Reliable Internet Routing

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Abstract

Network routing algorithms responsible for selecting paths to destinations have a profound impact on network reliability experienced by the network users. Unfortunately, performance of state-of-the-art routing algorithms often falls short of users' expectations.

(i) The flexibility with which operators of independently administered networks can choose their routing policies allows them to make selections that are "conflicting" and may lead to route oscillations. Oscillating routes have a negative impact on performance experienced by the user, and also cause overloading of the routers with control messages. (ii) Interdomain routing in the Internet is based on trust. As a result, false route announcements can be made by a malicious network operator. Such false announcements can be made even without knowledge of the network operator, e.g., due to accidentally misconfigurations or router hijacking. False route announcements may lead to denial of service, or worse yet, traffic can be intercepted without detection of both the sender and recipient. (iii) Even if network routes are stable and secure, unexpected equipment failures may cause performance degradation. It is difficult to pre-configure current routing protocols with all possible failures in mind, and not enough flexibility is offered to balance load in the network evenly.

This thesis addresses these three challenging problems. (i) We provide a new theoretical model of interdomain routing and derive the necessary and sufficient conditions that determine which policy combinations lead to route oscillations. Moreover, we also provide a practical polynomial-time algorithm that allows network operators to verify the existence of such conflicts. (ii) To secure routing against malicious attacks, we offer a new secure routing protocol that, unlike earlier attempts, is incrementally deployable. Our solution can protect both participants and non-participants if as few as 5–10 independently administered domains deploy our solution. (iii) To handle traffic engineering in the presence of failures, we propose a new architecture that optimizes load balancing for a wide range of failure scenarios. Our architecture supports flexible splitting of traffic over multiple precomputed paths, with efficient path-level failure detection and automatic load balancing over the remaining paths.

Collectively, the contributions of the dissertation provide tools that improve routing reliability and as a result network performance perceived by the user.

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Chapter 1

Introduction

Since the first message was sent on the ARPANET network, the predecessor of the modern Internet, the network has experienced a dramatic growth. Once reserved for academic use, the Internet presently connects billions of users around the world who rely on the infrastructure in their daily lives.

The Internet has become much more than just a network used to access information. In the past two decades, new important applications have emerged, such as electronic commerce, voice over IP, social networking, and many more. In addition, many applications such as online banking or online trading are business critical and time sensitive. As a result, users and businesses who rely on the Internet infrastructure require a high degree of reliability from the operators of the network. Reliability encompasses the ability to offer the network users high-bandwidth and low-latency service in the presence of accidental hardware failures or planned maintenance, and ability to deliver data securely even in the presence of malicious attacks on the Internet infrastructure. This thesis addresses these challenging problems.

Network routing, the selection of paths to destinations, is perhaps one of the most important features of the Internet that determines the performance, security, and reliability of the network. For example, selecting the “right” paths can reduce congestion and decrease queuing and propagation delay, improving the performance for the users. Furthermore, dynamic rerouting is a critical operation that ensures that connectivity is re-established after a failure of a link or router. Rerouting is also important when traffic demands of the users change. Finally, security of network routing protocols has implications on the reliability of the entire network – a malicious user may for example manipulate the routing protocol so that network traffic is forwarded to him, or to cause widespread connectivity disruptions.

In Section 1.1 we describe the basic operation of two main families of routing protocols – intradomain and interdomain routing protocols. Then, in Section 1.2 we describe the shortcomings of the current design and the impact on the safety, security and reliability of the network. We also outline the proposed solutions to address these important issues.

1.1 Routing in the Internet

Interdomain routing concerns the problem of calculating the *paths across domains* that the traffic needs to traverse to reach the destination. Intradomain routing determines the *path inside a single administrative domain* that the traffic needs to take to reach the destination. The two problems are very different. Intradomain routing is done in a single network and the owner of the network has a full control and information about the network topology, load, configuration, etc. Interdomain routing concerns exchanging traffic between separate networks whose owners, who are business competitors, do not have full information about the other networks. For this reason, interdomain and intradomain routing rely on different routing protocols and face different challenges.

1.1.1 Interdomain Routing

The Internet consists of tens of thousands of autonomous systems (ASes) that are independently owned and operated. To achieve global connectivity, ASes exchange information about reachability. This information exchange is facilitated by the Border Gateway Protocol (BGP) [92]. BGP is a path vector protocol, that is, when an AS uses BGP to announce a route to its neighbor, the announcement contains a list of all other ASes that the path traverses before reaching the destination. The adjacent ASes exchange the BGP messages between their edge routers, which are also sometimes referred to as BGP speakers.

ASes are typically Internet Service Providers who have business relationships with their neighboring ASes. These business relationships determine any transit fees. While business relationships are confidential, a model [19] that is believed to correspond to reality classifies business relationships into two categories: customer-provider and peer-peer. In customer-provider relationship, the customer has to pay the provider for all traffic that traverses the link between the ASes, no matter what the direction of the traffic. In peer-peer relationships, the peers forward traffic for each other free of charge.

The nature of business relationships determines which routes are preferred by ASes. For example, given the choice between a customer, peer and provider route, the AS will prefer the customer route which is the most profitable. Business relationships also play a role even after a BGP speaker selects the single route to the destination that it prefers – a BGP speaker will not announce a provider route to another provider as it would have to pay to both providers for the transit traffic. For this reason ASes need the flexibility to choose among multiple paths, and the option to announce the selected path to an arbitrary subset of their neighbors. BGP allows such flexibility – if an AS learns about multiple routes from its neighbors, it can apply an arbitrary policy to choose the preferred path, and decide which neighbors to announce the path to.

BGP is a protocol based on trust. When a route announcement is received, autonomous systems cannot verify whether a path announced by a neighboring BGP speaker corresponds to an existing physical path, and whether that path is available to the neighbor. For this reason, BGP is extremely vulnerable to malicious attacks where an attacker compromises a router to make false routing announcements, and to misconfigurations where a speaker mistakenly announces an incorrect route.

1.1.2 Intradomain Routing

Network operators need intradomain routing protocols that ensure network connectivity even as the network topology changes due to link additions, hardware failures, or during planned equipment maintenance. In addition, network operators desire to balance the load in their networks to avoid congestion. One protocol satisfying these goals is Open Shortest Path First (OSPF) [69]. OSPF is a link state routing protocol, i.e., a protocol that collects information from routers about their connectivity (the state of their links). Then, the routers construct a graph representing the network, and traffic is sent on the shortest path according to link weights that were pre-assigned to each link. If a router finds multiple shortest paths, traffic is split evenly on the outgoing links. The link state information is maintained by each router and if it changes, it is flooded in the network. The benefits of using OSPF include the ability to react to link failures – when a link fails the information is immediately flooded in the network and all of the routers can compute new shortest paths that avoid the failed link. Furthermore, proper link weight assignment allows load balancing. However, OSPF only allows to split the traffic on paths of the same minimal cost. This approach does not allow much flexibility, and if the same link weights are used before and after a failure, the performance may be suboptimal. Moreover, finding appropriate link weights is computationally hard.

Multiprotocol Label Switching (MPLS) [82] is a routing protocol that can be used to provide control over which flows traverse which paths. MPLS attaches labels to data packets, and forwarding decisions are made based purely on the content of the label. When a packet is received by a router, a label swap operation is performed. The old label is popped and another label is pushed on top of the label stack, and the packet is forwarded to the appropriate neighbor. An advantage of MPLS is that it can be applied to all data packets, such as ATM, SONET or Ethernet packets, irrespective of the lower-layer details of the corresponding protocols and technologies. MPLS can be used in conjunction with any standard IP routing algorithm to determine the routes that should be used. MPLS is often used in conjunction with OSPF and RSVP [17]. OSPF is used to calculate the desired set of routes, as described above, and the Resource Reservation Protocol (RSVP) is then used to configure the routers on the end-to-end paths.

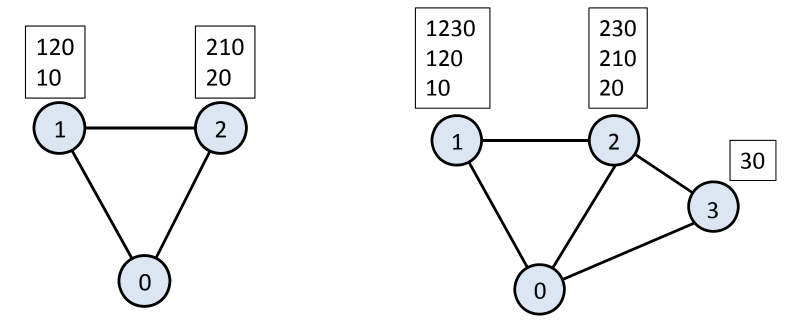
When a link fails, several mechanisms can be used to recover from the failure. Local path protection mechanisms are used to redirect traffic from a failed link onto an alternate path that connects the two link end points. Example of local path protection is MPLS Fast Reroute. The router that manages the backup path is called the Point of Local Repair (PLR), and the router where the backup path merges with the original path is called the Merge Point (MP). The primary benefit of Fast Reroute is its speed because the PLR can start forwarding packets on the precalculated backup path immediately after the failure is detected. Unfortunately, Fast Reroute often does not provide adequate performance because it can cause congestion in the neighborhood of the failed link. A more flexible mechanism that allows some end-to-end path restructuring is needed to balance the load more evenly. For this reason, network operators are often forced to perform end-to-end route reoptimization after a failure event.

1.2 Challenges with Network Reliability

Here we discuss three significant challenges that network operators must address to provide reliable service, and the respective solutions that we propose in this thesis. In Section 1.2.1 we explain how a combination of "conflicting" route preferences may lead to BGP oscillations with severe consequences on service availability and performance. In Section 1.2.2 we discuss how attacks on the BGP system can lead to widespread service disruptions. Finally Section 1.2.3 describes the challenges of path restoration after a component failure in the network of an Internet service provider.

1.2.1 Interdomain Routing Policies Leading to Oscillations

Although Autonomous Systems are free to choose their route preferences and route export policies, certain policy combinations can lead to permanent oscillations in the routing system. In these oscillations, routers exchange control plane messages in a cyclical fashion indefinitely. An example of such a policy choice is illustrated in in Figure 1.1(a). The two BGP speakers represented by nodes 1 and 2 are configured as follows. To reach node 0, node 1 prefers the indirect route through node 2 over the direct route. Similarly, node 2 prefers the indirect route through node 1. An oscillation may occur as follows. Initially the nodes select the direct route 10 and 20, respectively, and the nodes simultaneously update each other about the route they are using. Subsequently, they both switch to the two indirect routes 210 and 120, creating a transient loop between nodes



(a) Permanent oscillations may occur.

(b) The network will stabilize with every node picking its most preferred route.

Figure 1.1: Presence of a dispute wheel is necessary but not sufficient for oscillations.

1 and 2 that the network traffic follows. The loop is removed as soon as the routers update each other about their new route. However, if this update is once again performed simultaneously, the routes 10 and 20 will be selected, and the entire process can repeat itself.

Avoiding BGP oscillations is imperative as oscillations can have negative effect on both the control plane (routing messages exchanged by the routers) and the data plane (data packets traversing the network). In the control plane, oscillations increase the number of route updates, which may overload the routers that are not able to process the messages at high enough speed. This problem is significant because routers must already process route updates for some $350,000$ address prefixes in the Internet, causing a heavy load even without oscillations. Oscillations can also cause unacceptable delays or packet drops in the data plane. As illustrated above, loops in the data plane may arise and data packets may be dropped. Path changes during oscillations can also interfere with TCP that monitors end-to-end delays and expects regular timing of acknowledgements; oscillations may lead to severe performance degradation for the user.

The seminal work of Griffin, Shepherd, and Wilfong [44] developed the Simple Path Vector Protocol (SPVP), a theoretical model of BGP that provides the framework to study BGP stability. The most well-known result in that framework provides a sufficient condition for convergence – absence of a structure called dispute wheel is sufficient to prevent oscillations. Unfortunately, it is unlikely that the exact conditions of convergence in the SPVP model can be formulated – the problem of deciding convergence in an SPVP-like model is PSPACE-hard [30]. Furthermore, the model abstracts many implementation details of the BGP protocol, and it is possible to find network configurations where the BGP protocol oscillates but the SPVP model does not.

Checking the lack of oscillations by verifying the absence of dispute wheels is also problematic. First of all, a number of network configurations that have dispute wheels do not oscillate. Consider for example the configuration in Figure 1.1(b) that has a dispute wheel, but converges to the stable outcome where each node selects the most preferred route through node 3 to the origin 0. As dispute wheels can occur for legitimate reasons [37], a less strict condition that guarantees the absence of oscillations is desired. Another problem is that verifying whether a configuration contains a dispute wheel is an NP-hard problem.

Chapter 2 of this thesis develops a new theoretical model of routing that allows formulation of the necessary and sufficient conditions of convergence. We provide a simple polynomial-time algorithms that verifies these conditions. Furthermore, our new model of BGP includes features that have been omitted in the earlier models, making the model more accurate while retaining its simplicity.

1.2.2 Interdomain Routing is not Secure against Malicious Attacks

BGP is a protocol based on trust that does not authenticate route update messages. As a result, it is possible to launch a malicious attack that leads to traffic interception or loss of connectivity, as illustrated in Figure 1.2. Here the address blocks 12.8.1.0/24 and 12.34.0.0/16 belong to nodes 5 and 7, respectively. Node 1 is malicious and announces the two address blocks that it does not own. Some nodes will believe the malicious announcements and send traffic to the adversary. For example, if node 3 sends traffic to an address in the address block 12.34.0.0/16, it will reach node 1 instead of the valid destination in node 7.

Incorrect prefixes can be announced with malicious intentions, as described above, or due to misconfigurations. BGP attacks or misconfigurations have been occurring in the Internet with an alarming frequency and spectacular consequences.

In February 2008 Pakistan Telecom brought down YouTube worldwide for several hours when it tried to block local access to the service [50]. They mistakenly sent new routing information to PCCW, an ISP in Hong Kong, that propagated the route further.

In May 2003, spammers hijacked an unused block of IP address space owned by Northrop

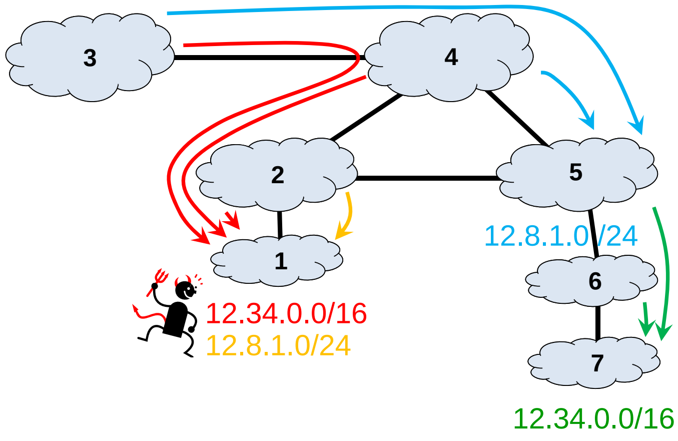


Figure 1.2: Malicious node 1 announces address prefixes that it does not own. Some nodes will believe the malicious node and forward it the traffic addressed to the concerned prefixes.

Grumman, and used the address space to send spam email [13]. It took two months for Northrop Grumman to resolve the situation and get the rogue routing announcements blocked.

For about 18 minutes on April 8, 2010, Chinese Telecom rerouted traffic destined to about 15% of the address space through servers in China. This incident reportedly [2] also affected traffic destined to US government and military sites, including those for the US Senate, four branches of the military, the office of the secretary of defense, and NASA. The network traffic was forwarded further to the destination. Such forwarding can be done transparently without the owner's knowledge, as described in [78].

Several cryptographic countermeasures have been proposed in the literature. However, deployment of these new protocols is complicated by the fact that participation of all ASes in the Internet is required to achieve significant security benefits. Since the Internet currently consists of approximately $35,000$ independently administered ASes, many of which are small regional ISPs or enterprise networks, speedy deployment of a new version of BGP cannot be expected. We argue that for any practical solution that will be eventually deployed, incentives must be provided to early adopters, and security benefits must exist even for small scale deployments.

Since privacy can be achieved using conventional cryptographic protocols by the two communicating users, this work focuses on service availability provided by the network. Our solution offers a novel secure BGP design that achieves significant security gains even for small scale deployments, e.g., deployments with as few as $5 - 10$ participating ASes. Our results indicate that such remarkable security benefits can be achieved through a combination of several mechanisms.

First, we require the participants to detect (and remove from further consideration) the malicious routes. This can be achieved by using, e.g., data plane monitoring that uses cryptographic techniques to verify if data was delivered to the user. Routes that do not pass this test are removed because they are either not reliable, or they are malicious. Second, we require that at least some "large" ASes (e.g., major national ISPs) participate. Third, we require that all participants announce the protected address prefixes of the other participant collectively – this makes it difficult for the adversary to launch a geographically widespread attack because it must compete with the announcements of all the participants. Finally, we require that the participants use overlay routing to reach the valid destination, as illustrated next. Assuming that nodes 3, 5, and 7 in Figure 1.2 are participants, node 3 can successfully forward data to the prefix 12.34.0.0/16 owned by node 7 by using the overlay to send the traffic first to node 5, which relays the traffic to node 7. Note that node 3 can reach the intermediate node 5 by using any address in the address block 12.8.1.0/24.

In Chapter 3 we describe our solution in detail and evaluate its security benefits using simulations. Our simulations rely on an accurate model of the Internet topology that was constructed by inspecting routing tables of Internet routers. The simulations model the behavior of BGP routers and simulate the impact of potential attacks.

1.2.3 Traffic Engineering after a Failure

Network operators need to carefully balance the load in the network in order to efficiently use the existing link capacity. In addition, they need to handle planned equipment maintenance and unplanned failures gracefully, without noticeable disruptions to the users. This challenging task is further complicated by the fact that traffic patterns change significantly during the day, and network operators need to satisfy strict service level agreements (SLAs) which specify, e.g., the maximum average delay that the network traffic can experience.

Figure 1.3 illustrates the advantages and disadvantages of two possible techniques that can be used to protect against failures – local path protection and global path protection. In local path protection, the failure is repaired locally by sending the traffic on an alternate route between the two endpoints of the failed link. The figure illustrates the disadvantage of this approach – congestion in the neighborhood of the failure where the rerouted traffic shares a link with another flow. Global path protection sends the traffic on an alternate end-to-end path, which allows to

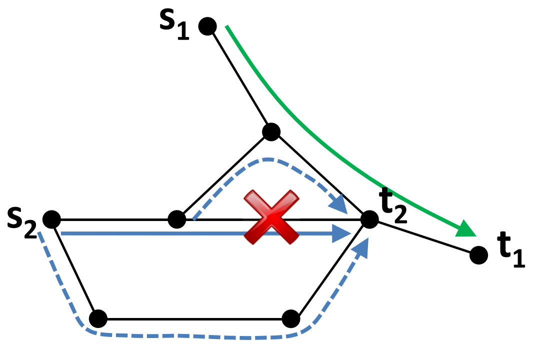


Figure 1.3: Two flows are shown by solid arrows. After a link failure, local path protection can be used to reroute the affected traffic (top dashed arrow), or global path protection can be used (bottom dashed arrow).

spread the load in the network more evenly, but at the cost of re-optimizing the route selection to minimize congestion. In practice, a combination of both techniques is used, local path protection for its speed, and global path protection because it allows the network operator to re-optimize the flow of traffic in the entire network according to the new conditions.

Our goal is to offer an alternative architecture to optimize load balancing under a wide range of failure scenarios that does not require route re-optimization after a failure. In Chapter 4 we propose to use an architecture that supports flexible splitting of traffic over multiple precomputed paths, with efficient path-level failure detection and automatic load balancing over the remaining paths. We systematically explore several possible load balancing algorithms that differ in their complexity and the amount of state they need to store in the network routers. Since the more complex solutions should allow better load balancing than their simpler counterparts, we explore this tradeoff experimentally. To perform accurate simulations, we use traffic measurements, topology, and failure data from a large ISP.

Our experiments allowed us to identify an architecture at the "sweet spot", an architecture that achieves near-optimal load balancing under a variety of failure scenarios with a relatively small amount of state in the routers. Besides its simplicity and deployability of the solution using current hardware, additional benefits include ability to use single configuration for traffic that varies throughout the day, and ability to achieve network propagation delay similar to delays experienced by routing algorithms that select paths to minimize this metric.

Chapter 2

Detecting Routing Protocol Oscillations

2.1 Introduction

The Border Gateway Protocol (BGP) [92], the de facto interdomain routing protocol in the Internet, offers autonomous systems (ASes) the flexibility to specify their custom routing policies. Unfortunately, this flexibility may result in policy choices that cause persistent oscillations. Such oscillations unnecessarily increase the number of BGP updates and negatively impact network traffic. Over the past decade, researchers have developed a good understanding of which combinations of routing policies lead to oscillations [23, 32, 37, 38, 43, 45, 86]. Most of these results were based on an abstract model of the interdomain routing system — namely the Simple Path Vector Protocol (SPVP) [44] — that captures how each node selects the highest-ranked path consistent with its neighbors' decisions.

This chapter shows that local engineering decisions, such as BGP timers and internal router structures, can produce short-term artifacts that lead to protocol oscillations not well modeled by SPVP. To capture how these local phenomena affect global convergence, we introduce an extension of SPVP called the Dynamic Path Vector Protocol (DPVP). Although DPVP is seemingly more complicated than SPVP, it actually yields to analysis more easily: we show that DPVP admits a *necessary and sufficient condition* of convergence. Furthermore, we give an algorithm that for most realistic settings efficiently determines whether a DPVP instance is safe, i.e., whether the BGP system as modeled by DPVP converges.

2.1.1 Spurious Selection of Lower-Ranked Routes

Earlier studies of interdomain routing assume that routers select and announce the most-preferred available route. However, routers in practice may temporarily announce other recently-available routes, or even withdraw a route when the destination appears reachable. We call such unexpected announcements and withdrawals *spurious updates*. These spurious updates can be caused by several router-level mechanisms that delay the propagation of update messages (to reduce overhead and improve stability) or limit visibility into the alternate routes (to improve scalability), including:

- **Route flap damping** [96]: Route flap damping temporarily suppresses a route if it appears unstable. As a result, a router may temporarily select a less-preferred route.
- **MRAI timers** [81]: The Minimum Route Advertisement Interval (MRAI) timer paces BGP update messages. Delaying message delivery can cause a router to temporarily select a lower-ranked alternate route.
- **Router queuing mechanisms**: The BGP message queues between routers delay the delivery of updates. These queues, coupled with optimizations that stop generating new messages when the queue grows large, can lead to delays in selecting the highest-ranked route.
- **Cluster routers**: Large routers are distributed,with BGP sessions terminating on different processor blades. To improve scalability, these blades do not exchange full information with each other, which may lead to a temporary selection of a less-preferred route.
- **Proposed router extensions:** Extensions to the BGP route-selection process were proposed to improve router reliability [57] or to reduce convergence time [93]. This changes the timing of routing decisions.

All spurious updates share two common properties: (i) a router can only send spurious updates for a short time after receiving information changing its most preferred route, and (ii) spurious updates are based on routes that have been recently available (including spurious withdrawals because "no route" is always available). DPVP allows any spurious update with these properties. We argue that such model is general enough to capture all spurious updates, but at the same time we show that the model is not overly broad.

Just as local routing policies can affect global convergence [44], the local engineering decisions that cause spurious updates can also trigger oscillations, and slow convergence exponentially. Eliminating all sources of spurious updates would require major changes to router design and the BGP protocol. Some of these mechanisms are important for reducing protocol overhead and improving scalability, making it unappealing to eliminate them entirely. Protocol designers, router designers, and network operators could strive to reduce the frequency and duration of spurious updates. However, it is not clear that such a quest is warranted or plausible. Rather than advocating for a world free of spurious updates, we argue for a better understanding of their consequences.

2.1.2 DPVP Convergence

While allowing spurious updates shrinks the set of BGP configurations that are safe from oscillations, we establish that *most* of the well-studied situations deemed safe under SPVP remain safe even under DPVP. In particular, we strengthen the SPVP-based results of [44] to show that even DPVP is safe in a network without a "dispute wheel" structure. Thus, spurious updates do not affect the large body of research on *safety in dispute-wheel-free settings*. In contrast, BGP safety in more general settings, as well as convergence time, *can* be adversely affected by spurious updates, as illustrated in Section 2.6.

Our main positive result on convergence that also demonstrates the power of DPVP is a *combinatorial necessary and sufficient characterization of safe DPVP instances, which is tractable under most typical settings*. We show that a DPVP instance is unsafe if and only if it admits a certain combinatorial structure we call a "CoyOTE" (explained in Section 2.8). Although DPVP adds the "complexity" of spurious updates over SPVP, this characterization is surprisingly nice in several aspects that have been elusive for SPVP:

- **Bijectivity**: The absence of CoyOTEs is necessary *and* sufficient for safety. Prior work has only yielded sufficient but not necessary [23, 37, 38, 43, 91], or necessary but not sufficient [32, 86] conditions of convergence. Griffin et al.'s best-known result [44] shows the absence of dispute wheels to be sufficient for safety. Also sufficient are the Gao-Rexford conditions that constrain the network's economic structure [38]. Cittadini et al. [23] derived necessary and sufficient conditions for the stricter criterion of "safety under filtering" [32], which requires convergence even if an arbitrary subset of routes are permanently filtered by each router. However, those conditions are just sufficient but not necessary for safety under a fixed filtering.

- **Tractability in most common cases**: Checking whether a network admits a CoyOTE under general routing policies is NP-complete, just like the weaker question of checking for the sufficient-only condition of No-Dispute-Wheel [44]. Luckily, we were able to find a polynomial time algorithm that verifies safety of BGP configurations for virtually any policy used by network operators in practice.

- **Verifiability**: Given a CoyOTE structure, one can easily *verify* its validity as a proof that a network is unsafe. On a more theoretical note, this also places the formal problem of DPVP safety in complexity class CoNP, relatively much easier than the PSPACE-complete problem of checking safety in a comparable SPVP setting [30].

Roadmap: In Section 2.2 we review the Stable Path Problem (SPP), a general framework for describing interdomain routing. In Section 2.3 we formally introduce our DPVP model of BGP built on top of SPP that captures the effects of spurious updates on worst-case BGP convergence. To demonstrate DPVP's versatility and applicability, Section 2.4 describes a variety of real and proposed router behaviors that could cause temporary announcements of lower-ranked routes, and demonstrates global oscillations caused by these behaviors, yet not predicted by the classical SPVP model of BGP. Section 2.5, conversely, establishes that DPVP is not over-broad: we show that any sequence of events allowed by DPVP might indeed occur from combinations of the above causes. Section 2.6 shows examples of theoretical results in the literature that, while correct under the SPVP model, no longer hold in the presence of spurious updates. We show the No-Dispute-Wheel DPVP safety condition in Section 2.7, and the necessary and sufficient condition for safety in Section 2.8. Section 2.9 presents an algorithm for checking DPVP safety in polynomial time for all "realistic" BGP policies. Finally, Section 2.10 shows that this "realistic policy" constraint is necessary — allowing truly arbitrary policies makes it NP-complete to verify safety.

2.2 The Stable Paths Problem (SPP)

Here we review the Stable Paths Problem (SPP) due to Griffin et al. [45]. The reader familiar with the SPP framework may proceed directly to Section 2.3.

The SPP [45] consists of a graph where each node represents a single BGP speaker, and a fixed node which all other nodes try to reach. Each node has its own set of permitted paths to the origin, and a ranking function that ranks the permitted paths in the order of preference. A solution of the SPP is a global assignment of nodes to permitted paths such that each node is assigned the highest ranked path that can be constructed based on the paths assigned at neighboring nodes. The formal definition of SPP follows.

The simple undirected graph $G = (V, E)$ with nodes $V = \{0, 1, 2, ..., n\}$ represents the network topology. Node 0 is the address *origin* and all other nodes try to establish a path to the origin. Let $\mathsf{neighbors}(v)$ denote the neighbors of node v.

Paths are represented as a sequence of nodes $(v_k\ v_{k-1}...v_1v_0)$ where for each $k \geq i > 0$ we have $(v_i, v_{i-1}) \in E$. An empty path is denoted ϵ. If two paths P and Q are not empty, and the last node in P is the same as the first node in Q, the *concatenation* of the two paths is denoted PQ. A *subpath* of the original path $P = (v_kv_{k-1}...v_1v_0)$ from node v_i to v_j for some $i > j$ is $P[v_i, v_j]$, and $P[v_i]$ denotes a subpath from v_i to the origin. We use $v \in P$ to denote that node v appears in path P.

The permitted paths to the origin are explicitly specified for each node. The set of *permitted paths* for each $v \in V$ is $\mathcal{P}^v$. Any path P that appears in the set is *permitted* at the node v. $\mathcal{P}^0$ is well defined and contains the only valid path to the origin, i.e., the empty path ϵ. Let the collection of all permitted paths be $\mathcal{P} = \{\mathcal{P}^v | v \in V\}$.

Route preference of each node $v \in V$ is captured by its *ranking function* λ^v. If two paths $P_1, P_2 \in \mathcal{P}^v$ and $\lambda^v(P_1) < \lambda^v(P_2)$ then P_2 is *preferred* to P_1 by v. Let the collection of all ranking functions be $\Lambda = \{\lambda^v | v \in V\}$.

Additional requirements pertain to the permitted paths and their ranking. For each λ^v and $\mathcal{P}^v$ we require:

(i) **Paths are simple:** every non-empty path in $\mathcal{P}^v$ is a simple path from v to the origin.

(ii) **Empty path permitted:** $\mathcal{P}^v$ contains the empty path ϵ.

(iii) **Empty path lowest ranked:** $\lambda^v(\epsilon) < \lambda^v(P)$ for all $P \in \mathcal{P}^v$.

(iv) **Strictness:** if $\lambda^v(P_1) = \lambda^v(P_2)$ then either $P_1 = P_2$ or the first edge of the two paths is the same.

A path assignment is a function π that maps each node v to a path in $\mathcal{P}^v$. $\pi(v) = \epsilon$ denotes that node v is not assigned a path to the origin. We write a path assignment as a vector $(P_1, P_2, ..., P_n)$ where $\pi(v) = P_v$, and the path of the origin to itself is omitted. Let $\mathsf{choices}(\pi, v)$ be the set of all possible permitted paths at v that extend the paths assigned to their neighbors:

$$\mathsf{choices}(\pi, v) = \begin{cases} \{(v\ u)\pi(u) | (v, u) \in E\} \cap \mathcal{P}^v & v \neq 0 \\ \{\epsilon\} & \text{o.w.} \end{cases}$$

Let W be a subset of permitted paths $\mathcal{P}^v$ such that each path has a distinct next hop. The best path in W is:

$$\mathsf{best}(W, v) = \begin{cases} P \in W \text{ with maximal } \lambda^v(P) & W \neq \emptyset \\ \epsilon & \text{o.w.} \end{cases}$$

The path assignment π is *stable* at node v if $\pi(v) = \mathsf{best}(\mathsf{choices}(\pi, v), v)$.

The SPP specification is a triple $S = (G, \mathcal{P}, \Lambda)$ consisting of the graph, permitted paths, and ranking functions. The specification S is *solvable* if there exists a stable path assignment π for S, otherwise it is *unsolvable*.

2.3 DPVP: BGP Model with Spurious Updates

To study the dynamic properties of BGP, we introduce the Dynamic Path Vector Protocol (DPVP), a formal model that allows transmission of stale information in spurious route updates. The DPVP model specifies the dynamics of routing information exchange between routers in the SPP framework. Section 2.3.1 informally explains how we model spurious updates. Then, Section 2.3.2 defines DPVP dynamics by specifying how a node exchanges routing information and selects a preferred route. A convenient shorthand notation that provides a compact description of a dynamic evolution of the DPVP model is introduced in Section 2.3.3.

2.3.1 Modeling the Spurious Updates

For a short period after receiving information that changes its best path, a router may temporarily transmit stale information in the form of *spurious* route announcements or withdrawals. An upper bound on the duration of the spurious behavior is required to prevent propagation of arbitrarily old information. The DPVP model introduces a universal fixed constant τ[1] that serves two purposes. First, it limits the interval after a route change at node v during which stale information may propagate from that node. Second, any stale information that propagates from node v at time t must have been available at node v at some point in the time interval $[t - \tau, t]$.

Specifically, the constant τ serves as an upper bound on the communication delay caused by queuing delays, the MRAI timer, the suppression period of route flap damping, and any other source of spurious behaviors, current or future. Indeed, we deliberately do *not* model the specific sources of spurious updates, so as to not limit our model to the sources thus far observed. Surely other sources may be buried deep inside current router designs, or may arise in the future, and we assert that modeling all of them with a generic finite cutoff is the right approach. That is, we expect that any future design decision that violates this model (i.e., potentially sends spurious updates indefinitely in an otherwise-stable system) would not be accepted by the network operator community.

2.3.2 Dynamic Path Vector Protocol (DPVP)

The current time of a global clock is denoted by t.

The internal state maintained by each node v consists of the following. The assigned path $\pi(v)$ represents the most preferred route that is consistent with the information received by the node at the present time. The structure $\textsf{rib-in}(v \Leftarrow w)$ maintained by node v contains the most recently processed information received from node w. The set $\textsf{recentRts}(v)$ contains all routes that node v has had recently available. This set includes any route that is available at the present time t according to the information in the $\textsf{rib-in}$ structure, as well as any route that was available in the time interval $[t - \tau, t]$. The state also includes variable $\textsf{stableTime}(v)$ which determines stability of the node as follows.

The stability of a node determines the properties of the information transfer from that

[1]Stability of an SPP instance in the DPVP model is independent of the actual numerical value of τ for $0 < \tau < \infty$.

node. The node v is stable if $t \geq \mathsf{stableTime}(v)$ and it is not stable otherwise. If a node v is *stable*, any information transfer in the system concerning the assigned path $\pi(v)$ must be accurate, i.e., the neighbors of node v learn the correct most recent route $\pi(v)$. However, if a node is *not stable*, then its neighbor may receive stale information consisting of any single path from $\mathsf{recentRts}(v)$.

The dynamic route information exchange is facilitated by *edge activations.* Simultaneous edge activations are allowed. When the edge (w, v) activates, the process shown in Figure 2.1 is executed. The "if" branch on lines 2–3 is executed if at the time of activation the node w is stable. The $\mathsf{rib\text{-}in}(v \Leftarrow w)$ variable is updated with the most recent information from node w. If node w is not stable, then lines 4–6 are executed, and node v learns information that is potentially stale. Stale information either contains a route withdrawal, or announcement of some route recently available at node w. The commands on lines 7–9 update the list of the recently available routes $\mathsf{recentRts}(v)$. Newly available routes are added, and if a route becomes unavailable at time t, it is scheduled for removal from $\mathsf{recentRts}(v)$ at time $t + \tau$. Finally, the if statement on lines 10–12 determines whether the best route available to node v consistent with the information received thus far changes. If the route changes then $\pi(v)$ is updated accordingly and the node is marked as unstable for a time period τ.

An edge activation sequence σ of sets $(E_0, E_1, \ldots)$ has E_t containing the edges that are activated at time t. An activation sequence is *fair* if each edge $e \in E$ appears in the sequence infinitely often, i.e., all node pairs continue exchanging routing information indefinitely.

```
activate($v \Leftarrow w$)
 1: old-rib-in := rib-in($v \Leftarrow w$)
 2: if $t \geq$ stableTime($w$) then
 3:     rib-in($v \Leftarrow w$) := $(vw)\pi(w)$
 4: else
 5:     pick some $P \in \{$recentRts$(w) \cup \epsilon\}$
 6:     rib-in($v \Leftarrow w$) := $(vw)P$
 7: if rib-in($v \Leftarrow w$) $\neq$ old-rib-in then
 8:     add rib-in($v \Leftarrow w$) to recentRts($v$)
 9:     remove old-rib-in from recentRts($v$) at time $t + \tau$
10: if $\pi(v) \neq$ best(rib-in, $v$) then
11:     $\pi(v)$ := best(rib-in, $v$)
12:     stableTime($v$) := $t + \tau$
```

Figure 2.1: The DPVP model for router v responding to the activation of edge $v \Leftarrow w$, i.e. v processing information from w.

A vertex activation sequence ρ of sets $(V_0, V_1, \ldots)$ has V_t containing the vertices that are activated at time t. When a vertex v activates, all its adjacent edges $(w, v) \in E$ activate simultaneously. We introduce vertex activations merely for convenience to allow more compact notation.

DPVP is stable at time t if the path assignment π is stable and it has not changed in the time interval $[t - \tau, t]$. Note that if DPVP is stable, it is impossible for nodes to exchange stale information, and the state cannot change at any later time.

DPVP is safe if any fair activation sequence, from any starting state, always converges to a stable state. Note that safety in the DPVP model is independent of the numerical value of the constant τ if $0 < \tau < \infty$ because the model does not limit the number of activations that can occur during any time interval τ. We define safety under filtering in the same way as [23, 32] do. DPVP is **safe under filtering** if it remains safe under removal of arbitrary subsets of paths from an arbitrary subset of the $\mathcal{P}^v$s (this generalizes the removal of arbitrary nodes and edges).

2.3.3 A Shorthand Notation

We introduce a shorthand state transition notation that concisely describes allowed oscillations caused by spurious updates in the DPVP model. A systematic treatment of the causes and consequences of spurious updates follows in Section 2.4.

An unsafe example of a network configuration is depicted in Figure 2.2. The network contains three nodes which attempt to obtain a route to node 0. Each node is annotated with its permitted paths, and these paths are listed in the order of decreasing preference. For example, node 1 prefers the path 1230 over 10. To demonstrate that the configuration is unsafe, we must find an oscillation, i.e, an initial path assignment, an activation sequence that activates every edge, and possible spurious announcements that cause a cyclical change of the path assignment. As long as the same activation sequence and spurious announcements are repeated, the oscillation persists.

The shorthand state transition notation that captures a possible oscillation in Figure 2.2 is as follows:

$$(10, 20, 30) \xrightarrow{2,3} (10, 210, 30) \xrightarrow{1;(1 \Leftarrow 2:230)} (1230, 210, 30) \xrightarrow{2} (1230, 20, 30) \xrightarrow{1} (10, 20, 30).$$

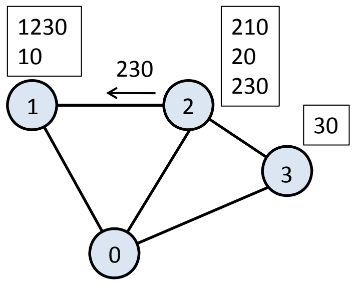


Figure 2.2: Example of an oscillation in DPVP. Node 2 exports a low ranked route 230.

The initial path assignment is $(10, 20, 30)$, nodes 1, 2, and 3 have paths 10, 20, and 30 respectively. The nodes or edges activated in each step are listed above the arrow. For example, the path assignment $(10, 210, 30)$ is reached from the initial state by activating nodes 2 and 3. If a spurious announcement is made, this is also described above the arrow. For example, $\xrightarrow{1;(1\Leftarrow 2:230)}$ represents an activation of node 1 where node 1 learns about route 230 from its neighbor 2. The spurious announcement of route 230 is allowed by the DPVP model because $\mathsf{recentRts}(2)$ contains route 230 and the path assignment of node 2 keeps changing during the outlined oscillation.

It is important to realize that not every shorthand notation that can be written down corresponds to a valid evolution in the DPVP model. Consider for example the following:

$$(10, 210, 30) \xrightarrow{1,2,3;(1\Leftarrow 2:230)} (1230, 210, 30) \xrightarrow{1} (10, 210, 30).$$

This notation is invalid because node 2, which is a stable node with a fixed path assignment 210, cannot spuriously announce the low ranked route 230 in DPVP.

2.4 Expressiveness of DPVP

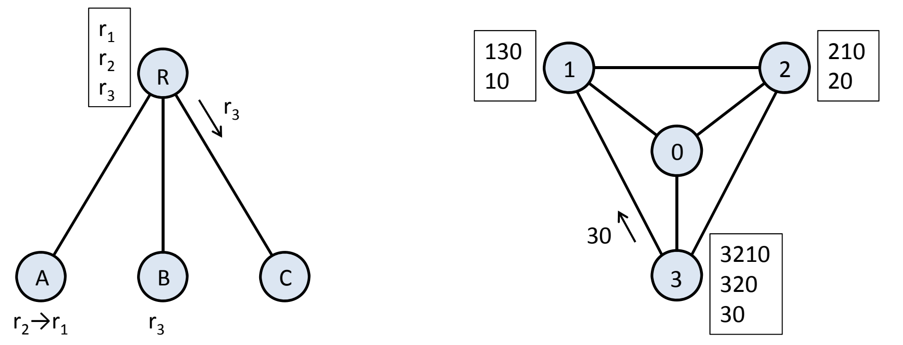
Having specified DPVP formally, we need to establish that it is a realistic model of BGP. We discuss several key sources of spurious updates in BGP, and demonstrate how the DPVP model captures them. These sources of spurious updates include (i) **mechanisms that introduce delays** to improve stability and reduce overhead, e.g. route flap damping and MRAI timers, and (ii) **mechanisms that limit route visibility** due to scalability requirements, e.g. specifics of router implementations. While this list is necessarily non-exhaustive, DPVP is an *abstract* model, and hence it is able to capture other sources of spurious updates as well.

We show specific examples of network configurations where the sources of spurious updates, such as route flap damping or router architectures, cause persistent BGP oscillations. These oscillations are correctly captured by the DPVP model, but the earlier established models of BGP are usually not able to model them.

2.4.1 Route Flap Damping

The route flap damping mechanism is used to limit the propagation of unstable routes [96]. When it is enabled, a BGP router maintains a penalty associated with *every* prefix announced by *each* BGP neighbor. Upon receiving a route update from a neighbor, the router increases the penalty. If the penalty exceeds a given suppression threshold, the route is tagged when it is inserted into the RIB. Tagged routes are not used in the route selection process, and a route with a different next hop will be used. The penalties decay exponentially in time, and if a route doesn't change, its tag is eventually cleared and the route may be used. Next, we show how route flap damping may cause spurious updates, which may in turn lead to permanent oscillations.

A spurious route announcement occurs when the route flap damping mechanism temporarily suppresses a route that would otherwise be preferred. Consider Figure 2.3(a) where the router R initially learns routes r_2 and r_3 from its neighbors A and B. The router announces



(a) After the update $r_2 \rightarrow r_1$, the less preferred route r_3 is temporarily selected.

(b) If node 3 suppresses routes from node 2, it must announce the route 30. This may cause oscillations.

Figure 2.3: The effects of route flap damping.

route r_2 to router C. If the route r_2 is updated to route r_1, and the penalty associated with BGP speaker A exceed the suppression threshold, router R temporarily suppresses route r_1 and selects and exports route r_3. After the penalty decreases, route r_1 is selected and exported. This appears as a spurious announcement to router C. Route flap damping may also cause a spurious withdrawal. For example, if router R was only connected to routers A and C, the same route update would lead to a route withdrawal from router C. These spurious updates are allowed in the DPVP model.

A permanent oscillation caused by route flap damping may occur in Figure 2.3(b). First, we will convince ourselves that the configuration is safe in the absence of spurious updates: node 2 must choose either 210 or 20, and thus node 3 will choose either 3210 or 320. Therefore node 1 must choose 10 and the stable state is $(10, 210, 3210)$. However, the following oscillation is possible in DPVP:

$$\begin{aligned}&(10, 20, 320) \xrightarrow{2} (10, 210, 320) \xrightarrow{3} (10, 210, 3210) \xrightarrow{1;(1 \Leftarrow 3:30)} \\ &(130, 210, 3210) \xrightarrow{1,2} (10, 20, 3210) \xrightarrow{3} (10, 20, 320).\end{aligned}$$

Indeed, this oscillation may occur due to route flap damping. Initially, node 2 activates and changes its route from 20 to 210. When node 3 activates, it processes the route update from node 2, which triggers the route flap damping mechanism. Although node 3 enters the state 3210 in DPVP, in the real BGP system the route 3210 is suppressed and the route 30 is used instead. This explains why in the next activation the spurious announcement 30 is made, and the system enters state $(130, 210, 3210)$. Assuming that the damping penalty decreases, the subsequent activations do not contain any spurious updates, and the system eventually enters the state it started in. This example demonstrates that route flap damping may cause unexpected oscillations that are not predicted by the earlier models of BGP.

2.4.2 MRAI Timer

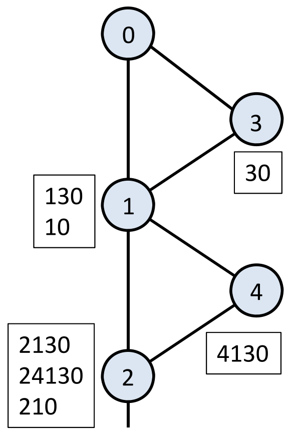
The MRAI timer [80, 81] may also exhibit unexpected spurious updates. When a new route is announced to a peer, subsequent route updates are postponed until the MRAI timer expires[2]. The

[2]The default value is 30 seconds in eBGP and 5 second in iBGP. However, the values used in practice range between 0 and 30 seconds.

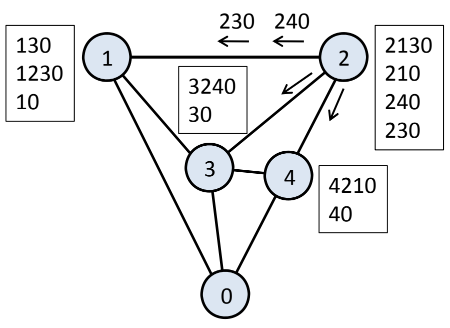
MRAI timer is applied to route announcements and, depending on the implementation, may [81] or may not [80] be applied to withdrawals. Some previous models of routing do not capture the asynchrony caused by MRAI timers. One such example is a variant of the Simple Path Vector Protocol (SPVP) with vertex activations [44, 46]. We show that when MRAI timers are used, an AS may unexpectedly lose connectivity or select a route which would not be selected in the SPVP model of routing. This may in turn lead to unexpected oscillations.

An unexpected spurious announcement caused by the MRAI timer is illustrated in Figure 2.4(a). The simplified variant of the SPVP model with node activations does not allow node 2 to select the route 24130. This can be explained as follows. Node 2 can only learn route 24130 after node 1 learns route 130, but then node 2 should select route 2130. However, the behavior of real BGP with MRAI timers differs. Let's assume that node 1 learns route 10 and exports it to node 2. Then it learns route 130, but cannot export it to node 2 because the corresponding MRAI timer has not expired yet. Then node 2 may select the route 24130 learned from node 4, but cannot select route 2130 until the timer in node 1 expires. The announcement of route 24130 by node 2 is possible in our DPVP model as a spurious announcement. The SPVP model with *edge* activations also allows this announcement.

A permanent oscillation caused by MRAI timers may occur in Figure 2.4(b). This gadget originally appeared in [22] as Figure 3. They show that this gadget is safe in the SPVP model with



(a) Node 2 temporarily selects route 24130.



(b) MRAI timer in node 2 may cause permanent oscillations.

Figure 2.4: The effects of MRAI timers.

node activations, but it may oscillate in SPVP with edge activations. We show that the gadget may also oscillate as a result of node 2 using the MRAI timer. Indeed, our DPVP model allows the following oscillation:

$$(1230, 240, 30, 4210) \xrightarrow{2} (1230, 230, 30, 4210) \xrightarrow{3;(3\Leftarrow 2:240)} (1230, 230, 3240, 4210)$$
$$\xrightarrow{1;(1\Leftarrow 2:240)} (10, 230, 3240, 4210) \xrightarrow{2} (10, 210, 3240, 4210) \xrightarrow{1,3,4;(\{1,3,4\}\Leftarrow 2:230)}$$
$$(1230, 210, 30, 40) \xrightarrow{2,4} (1230, 240, 30, 4210).$$

Initially, node 2 activates and changes its route from 240 to 230. This route change prevents the node from immediately announcing the new route 230 to its neighbors due to the MRAI timer. Therefore, when nodes 3 and 1 activate in the next two rounds, they receive the stale route 240. The MRAI timer is also invoked during the second to last round of activations when nodes 1, 3, and 4 receive the stale route 230. In conclusion, MRAI timers may be responsible for spurious announcements that cause permanent oscillations.

2.4.3 Lack of Route Visibility

Oscillations can be also caused by spurious updates resulting from a temporary loss of route visibility. We describe such losses of visibility due to the increasingly popular cluster-based router architectures.

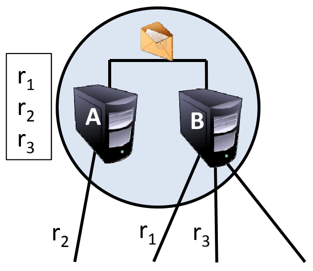
Distributed cluster-based routers parallelize functionality across multiple cores and across multiple server blades within each router [4, 27]. These architectures are becoming more common due to the need to scale to larger port densities and traffic demands at a reasonable cost. A router consists of multiple control processor blades, each handling a subset of the BGP sessions. Each blade runs its own software and exchanges reachability information with other blades. While the details of this information exchange differ from one implementation to another, scalability requires each processor blade to usually only announce the currently used route (the best route) to the other blades. Due to asynchrony, a blade may be temporarily unable to see a more preferred route learned by some other blade. This may lead to spurious updates being sent.

A spurious route announcement due to loss of visibility may occur, for example, after the most preferred route is withdrawn, and the second best route is not visible. Consider the example in Figure 2.5(a) which consists of two communicating processor blades. The cluster-based router prefers route r_1 to r_2 which is still more preferred than r_3. When all three routes are available, blade B selects route r_1 and announces it to the other blade A. If route r_1 is withdrawn, blade B must temporarily select route r_3 while it waits to learn about route r_2 from the other blade. Therefore, blade B spuriously announces route r_3 to other external BGP speakers. A spurious withdrawal can be caused if both routes r_1 and r_3 are withdrawn simultaneously. In such a case blade B must withdraw the route r_1 from its external BGP neighbor until it learns a valid route from the other blade.

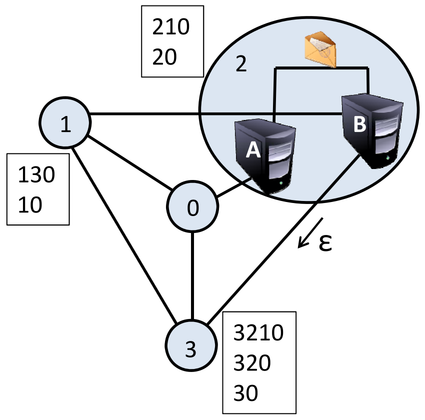
A permanent oscillation due to temporary lack of route visibility may occur in Figure 2.5(b). First, we observe that if no router sends spurious updates, the configuration is safe. This is the same configuration as in Figure 2.3(b) and hence the stable state must be $(10, 210, 3210)$. However, the following oscillation is allowed in DPVP:

$$(130, 210, 3210) \xrightarrow{2} (130, 20, 3210) \xrightarrow{1,3;(3\Leftarrow 2:\epsilon)} (10, 20, 30) \xrightarrow{2} (10, 210, 30) \xrightarrow{1,3} (130, 210, 3210).$$

In the second round of activations, node 3 receives a spurious withdrawal from node 2. This is explained as follows. Initially, node 2 was in state 210 and blade B used and exported the route 210. However, after node 1 switched to state 130, the route 210 was implicitly withdrawn, and



(a) After route r_1 is withdrawn, blade B temporarily announces r_3.



(b) The temporary lack of visibility of route 20 by processor blade B causes permanent oscillation.

Figure 2.5: The internal architecture of routers (or ASes) is a cause of spurious updates.

blade B was temporarily left without a route. Before blade B learned about route 20 from the other blade, it sent a spurious withdrawal to node 3. Once again, this example demonstrates that spurious updates may cause unexpected oscillations.

This chapter focuses on applying the DPVP model to model the router-level structure of the Internet. However, *DPVP may also be used to model entire ASes as nodes*, if the policies of routers inside an AS are consistent. This model is much coarser, and abstracts away many relevant intra-AS details, but is often reasonable since, from an external viewpoint, there is often very limited information about intra-AS router structure. In such a scenario, **the internal structure of an AS** may cause spurious updates, as well. The situation is analogous to the one with cluster based routers, individual routers correspond to processor blades and the communication of these routers is facilitated by route reflectors [12]. A subset of routers is assigned to each route reflector, which exchanges routing information with these routers and with other reflectors. Each router only learns one best route from each of its route reflectors, hence causing similar loss of visibility as we observed with cluster-based routers. This loss of visibility can be modeled by earlier BGP models, such as SPVP, *if* the internal structure of each AS is known, and each router and route reflector is represented as a separate node; on the other hand, DPVP allows us to consider questions of safety while remaining agnostic about the ASes' internal structures.

2.4.4 Router Queues

Unlike the classical SPVP model [45], DPVP does not explicitly model queues of BGP updates at each edge endpoint. However, the DPVP model captures any behavior that a queue could produce. If a node in the SPVP model processes a particular message from its queue that is older than the most recent update from the same sender, it may be modeled in DPVP as a spurious update. If a message is dropped from the queue, this may be modeled as a DPVP spurious withdrawal. On the other hand, DPVP intentionally allows a more varied behavior than per-edge queues do. While features like route flap damping may be modeled by particular patterns of dropped messages in a queue, some patterns of spurious updates valid in DPVP will not correspond to any possible queue behavior. It is worth noting that *an edge in DPVP does not necessarily model a single BGP session*, but rather models all BGP messages that may be exchanged between two potentially large, complex, distributed modern routers, which may share multiple physical links for redundancy, or might even have several BGP sessions. Since much is unknown in the public domain about the details of bleeding-edge intra-router architectures that vendors use today or may use in the near future, we intentionally limit our assumptions about exactly how two DPVP nodes will interact. We only assume the bare minimum restriction that we expect all routers to obey: the router acts on some reasonable timescale consistently like a monolithic BGP speaker, independently of its internal complexities.

Of course, if a DPVP instance is used to model AS-level behavior, there are many more possible sources of spurious updates from inter-router interaction. Those spurious events would not be well represented at all by any model involving a queue assigned to a node representing a whole AS.

2.4.5 Experimental Architectures

New router architecture proposals that aim to improve convergence or other properties of the routing system by changing the route selection algorithm also introduce spurious updates. For example, [57] builds a bug-tolerant router that runs multiple diverse copies of router software in parallel and uses voting to select the correct route. They achieve diversity by varying the ordering and timing of routing messages, and spurious inconsistencies in the decisions made by different virtual routers may cause spurious updates. Another example is [93] which uses heuristics to group BGP updates into priority classes to speed up convergence. Higher-priority updates are processed and propagated sooner, while lower-priority ones are delayed. Using examples similar as the ones in the previous subsections, it is easy to demonstrate that these mechanisms can also cause permanent oscillations. These oscillations are again accurately modeled by DPVP.

2.5 All Spurious Updates can be Realized

We demonstrate that any spurious behavior that is allowed in the DPVP model can be realized in the real interdomain routing system. Since we focus on the most general cases, our examples necessarily rely on complex configurations. We note, however, that the examples in Section 2.4 show that many spurious updates in the DPVP model are caused by much simpler configurations.

In general, we need to show that when the most preferred path available to a router changes from r_1 to r_2, the router may spuriously announce an arbitrary sequence of route updates $r_3, r_4, r_5, ...$

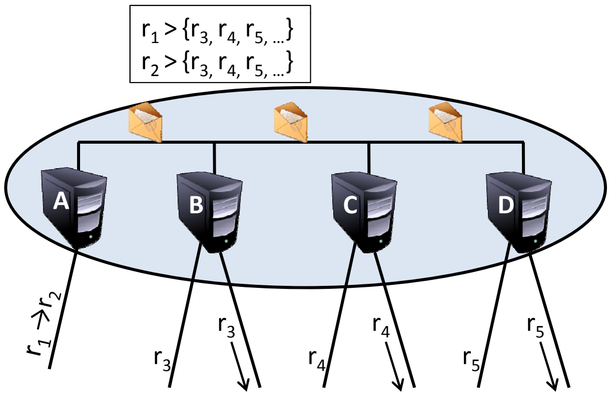


Figure 2.6: If blade A damps route r_2 after the update $r_1 \to r_2$, blades B, C and D announce routes r_3, r_4, and r_5 after they learn about the loss of route r_1.

to some neighbors before making the final announcement r_2. For full generality, we need to separately consider the two possible causes of the change of the most preferred route from r_1 to r_2. In the first, an update with route r_2 implicitly withdraws route r_1. In the second, route r_2 becomes available in addition to route r_1. We analyze the two cases separately.

The first case is illustrated in Figure 2.6 that depicts a cluster-based router consisting of four processor blades. The blade A receives an update with route r_2, which implicitly withdraws route r_1. Let us assume that this update triggers the route flap damping mechanism [96], and route r_2 is temporarily suppressed. Therefore, the router blade A announces to the other blades that no route is available to it. This information first reaches blade B, which identifies route r_3 as the currently best route that is available to it, and announces it to its external BGP neighbors. Similarly, blades C and D spuriously announce routes r_4 and r_5 after they hear from blade A. Assuming that the route flap damping ceases, and the internal state of the cluster-based router becomes consistent before further announcements are made, the final route r_2 is announced next. We conclude our analysis by noting that some of the routes $r_3, r_4, r_5, ...$ can be the empty routes ϵ, and hence spurious withdrawals can be made.

Figure 2.7 illustrates the remaining case where a new route r_2 becomes available in addition to a less preferred route r_1 which remains available as well. This example relies on the use of the Multi Exit Discriminator (MED). If one or more routers in an AS learn multiple routes from the same BGP speaker, these routes may be tagged with MED values. Routes tagged with smaller MED values are strictly preferred over routes with higher MED values. In our example in Figure 2.7,

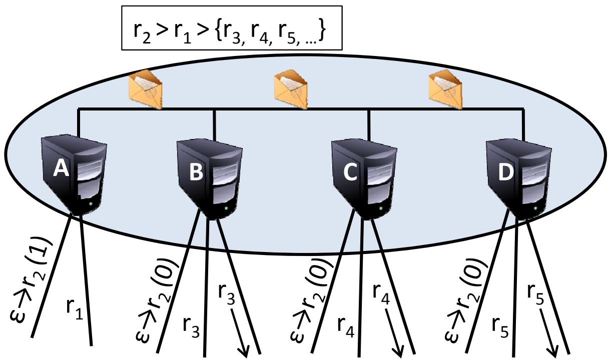


Figure 2.7: Route r_2 becomes available (numbers in parentheses are MEDs). If a small asynchrony causes damping in blades B, C, and D only, then these blades announce routes r_3, r_4, and r_5 respectively.

route r_2 has its MED values shown in parentheses. Route r_2 is announced to multiple blades on multiple interfaces perhaps because multiple parallel links connect the cluster-based router to the next hop on route r_2. This is a common occurrence in practice because high speed connections often consist of many parallel lower speed links. The link serviced by blade A has a higher MED value than the other ones, perhaps because it is a backup link, or because the network operator decided that the link needs to be taken down for maintenance in the near future, and traffic needs to be diverted.

The following dynamics in the system in Figure 2.7 leads to the desired spurious behavior. Initially, route r_1 is announced to the outside by each blade. Because each blade is running its own copy of the routing software, the blades must independently decide if an externally learned route should be damped. The timing of message processing may cause a slight asynchrony where one blade allows a route, whereas it is damped by another blade. Here we will assume that blades B, C, and D damp the newly learned route r_2 while blade A allows it. Therefore, blades B, C, and D will learn that blade A would like to use route r_2 (which should not be used due to its MED), and therefore they will export the spurious routes r_3, r_4, and r_5 similarly as in the previous example. After the internal state of the cluster-based router becomes consistent, the final route r_2 is announced.

2.6 Impact on Convergence

After having established that spurious updates may cause permanent oscillations in configurations that are otherwise stable, it is natural to ask if any existing results concerning convergence of BGP change with the introduction of spurious updates. We show in Section 2.6.1 that an exponential slowdown in convergence time may occur. Furthermore, in Section 2.6.2 we show an example of safety conditions in the literature that ensure safety in the absence of spurious updates, but that no longer hold in their presence.

2.6.1 Slower than Expected Convergence

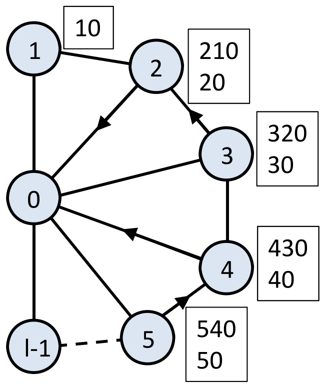
Understanding and improving the convergence time [46, 73] has been a central question in the BGP literature. It has been established that while the lower bound on convergence is in general exponential [31], a more favorable bound can be obtained in a Gao-Rexford [38] model of routing. In the Gao-Rexford model, every pair of neighboring ASes has a customer-provider relationship or a peering relationship, and no AS can become an indirect provider of itself. Furthermore, every AS prefers customer routes over peer or provider routes.

A recent paper of Sami et al. [86] shows that in the Gao-Rexford setting, the convergence time of BGP is linear in the depth of the customer-provider hierarchy, or more precisely it is at most $2l+2$ phases where l is the length of the longest directed customer-provider chain in the AS graph. We define the term *phase* below. We show an example of a network that, when spurious messages are allowed, the convergence takes $(2k+1)^{l-2}$ phases where k is the number of spurious messages that a node is allowed to announce after each route change. Our example, which is based on a topology that appears in the original work of Sami et al. [86], shows that *spurious updates may cause an exponential slowdown of convergence* even in the Gao-Rexford setting.

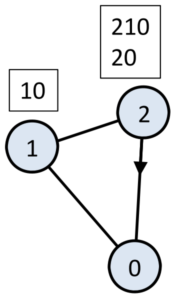
In their work a *phase* is defined as a period of time in which all nodes get at least one update message from each neighboring node, and all nodes are activated at least once after receiving updates from their neighbors. When a node activates, it processes the messages it previously received from all of its neighbors. The example which we describe next demonstrates that because the model of Sami et al. does not capture spurious route updates, it may lead to overly optimistic conclusions.

Figure 2.8(a) depicts a network with l nodes. Node 1 prefers route 10, and every other node i prefers route $i(i-1)0$ over the direct route $i0$. These path preferences satisfy the Gao-Rexford constraints if node 0 is a customer of every other node and node $i-1$ is a customer of node i for $2 \leq i \leq l-1$. The length of the longest directed customer-provider chain $0, 1, 2, ..., l-1$ is l. The arrows in the figure describe the initial routing choice of each node. Node 1 chooses the empty route, the even numbered nodes route directly to the origin 0, and all odd numbered nodes except node 1 route through the counter-clockwise neighbor. Except for node 1, the state is stable.

To show that the convergence in Figure 2.8(a) may take at least $(2k+1)^{l-2}$ phases in a model where each node is allowed to make at most k spurious updates after each route change, we first analyze the convergence of the smaller topology in Figure 2.8(b). The smaller topology is a special case of the larger one with $l=3$, and we will show convergence in $(2k+2)$ phases. In each phase, each node must activate at least once. Spurious updates are only used when we explicitly mention them. The initial state is $(\epsilon, 20)$. We assume that the two nodes activate simultaneously. After the first phase the state is $(10, 20)$, and after the second phase $(10, 210)$. In the third phase node 1 spuriously withdraws the route from node 2. The system reaches state $(10, 20)$ after the third phase, and state $(10, 210)$ after the fourth. This sequence is repeated, node 1 sends spurious updates in every odd phase for a total of k spurious updates, and the state flips back and forth between $(10, 20)$ and $(10, 210)$. The kth spurious update is made in phase $2k+1$ and the final state is reached in phase $2k+2$. Note that the route of node 1 changed once from ϵ to 10, but the route of node 2 changed $2k+1$ times.



(a)



(b)

Figure 2.8: Slowly converging network configurations.

A slow convergence in Figure 2.8(a) is achieved with the following timing of spurious updates. When the route of node i where $1 \leq i \leq l-1$ changes, this node makes a spurious announcement of the previously chosen route after each of the nodes with a higher node number sent k spurious updates and these nodes reached a stable state. Using the fact that in Figure 2.8(b) the state of node 2 changes $2k+1$ times, we conclude that whenever the route chosen by node i changes, the route chosen by node $i+1$ changes $2k+1$ times. Therefore, the route chosen by node $l-1$ will change $(2k+1)^{l-2}$ times and the number of phases required for convergence is $(2k+1)^{l-2}$.

2.6.2 BGP Without a Reel Unsafe

Although BGP convergence without spurious updates has been studied extensively, prior work either concerns *sufficient or necessary* conditions for safety. A classical result shows that the absence of a structure called a *dispute wheel*[3] is sufficient for safety [45]. Safety is also guaranteed when routing policies satisfy the conditions of Gao and Rexford [38].

The strongest result concerning BGP safety prior to this work was by Cittadini et al. [23]. They provide the necessary and sufficient conditions for safety *with route filtering.* Filtering allows each node to remove an arbitrary subset of paths from the list of permitted paths. They prove that instances that do not contain a dispute reel [23] are safe under any filtering, and if an instance contains a dispute reel, then there exists a filtering that allows oscillations. Note that these conditions become *sufficient* conditions for safety in the general setting without filtering. The dispute reel result no longer holds *if we allow spurious updates.*

Consider the example in Figure 2.9. This is the same topology that appears as Figure 4 in the original work of Cittadini [23] as an example of a safe topology without a reel (the dispute wheel with pivot vertices 1, 2, and 3 is not a reel because each pivot vertex appears in three rim paths, violating Definition 2.7.2). However, the following oscillation may occur:

$$(10, 20, 30) \xrightarrow{1,2,3;(2\Leftarrow 1:130),(3\Leftarrow 2:210),(1\Leftarrow 3:320)} (1320, 2130,$$
$$3210) \xrightarrow{1,2,3} (10, 20, 30).$$

This is a valid oscillation in the DPVP model where the spurious announcements may be

[3]Dispute wheel and dispute reel are formally specified in Definitions 2.7.1 and 2.7.2.

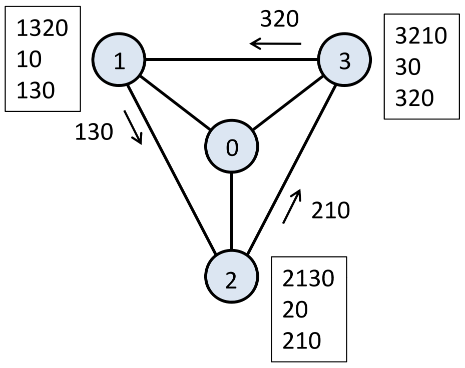


Figure 2.9: The graph does not contain a reel. However, spurious updates may cause oscillations.

caused, e.g., by the details of cluster-based hardware implementation of routers 1, 2, and 3. To make a spurious announcement, router 1 needs to have one router blade responsible for the BGP session with router 0, and another blade responsible for the other two BGP sessions. Routers 2 and 3 could use similar hardware architecture.

2.7 BGP Safety with Spurious Updates

The unexpected oscillations due to spurious updates beg the question of whether previous results on BGP safety continue to hold under the DPVP model. Fortunately, in Section 2.7.1 we are able to extend the well-studied No-Dispute-Wheel condition [45], sufficient for BGP safety in the SPVP model, to show that it is still sufficient for safety even with spurious updates. This result implies that the class of systems that oscillate due to spurious updates is relatively small, and most importantly, earlier results that use the absence of dispute wheel as a condition of safety hold even in the presence of spurious updates. While Section 2.6.2 showed that the absence of dispute reels is *not* sufficient for safety under filtering with spurious updates, Section 2.7.2 introduces a modified structure, a *two-third reel*, which we show to be necessary and sufficient under filtering.

2.7.1 No Dispute Wheel Implies Safety

The classical result by Griffin et al. [45] shows that BGP is safe in the SPVP model in the absence of *dispute wheels*, like the one in Figure 2.10, formally defined as follows:

Definition 2.7.1. [45] A *dispute wheel* $W = (U, \mathcal{Q}, \mathcal{R})$ of size k is a set of nodes $U = \{u_0, u_1, ..., u_{k-1}\}$ and sets of paths $\mathcal{Q} = \{Q_0, Q_1, ..., Q_{k-1}\}$ and $\mathcal{R} = \{R_0, R_1, ..., R_{k-1}\}$ such that the following conditions hold. For each $0 \leq i \leq k-1$, when all subscripts are interpreted modulo k:

(i) Q_i is a path from u_i to the origin.

(ii) R_i is a path from u_i to u_{i+1}.

(iii) $Q_i \in \mathcal{P}^{u_i}$ and $R_i Q_{i+1} \in \mathcal{P}^{u_i}$.

(iv) $\lambda^{u_i}(Q_i) \leq \lambda^{u_i}(R_i Q_{i+1})$.

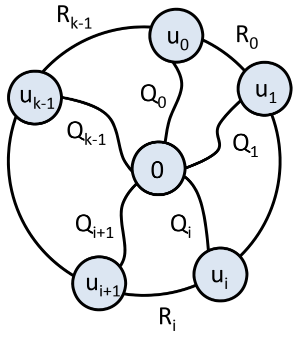


Figure 2.10: A dispute wheel of size k.

We strengthen Griffin et al.'s result to show that even with spurious updates, modeled by DPVP, BGP is *still* safe if there is no dispute wheel. This automatically strengthens the applicability of the large body of BGP literature that relies on the original No-Dispute-Wheel result under SPVP. We show:

Theorem 2.7.1. *DPVP instance with no dispute wheel is safe.*

We note that although the absence of a dispute wheel is sufficient for safety, it is *not necessary* in both DPVP and SPVP. That is, dispute wheels *can* occur in safe instances of the routing problem.

Rather than prove Theorem 2.7.1 separately, we actually derive it as a corollary of the stronger result in Section 2.7.2.

2.7.2 Safety with Filtering

Safety under filtering [32] studies convergence where an arbitrary subset of routes may be removed from the set of permitted paths $\mathcal{P}^v$. In [23] it was shown that the necessary and sufficient condition for safety under filtering in the classical SPVP model is the absence of a particular type of dispute wheel called a *dispute reel*:

Definition 2.7.2. [23] A *dispute reel* is a dispute wheel which satisfies the following conditions:

(i) **Pivot vertices appear in exactly three paths:** for each $u_i \in U$, u_i only appears in paths Q_i, R_i and R_{i-1}.

(ii) **Spoke and rim paths do not intersect:** for each $u \notin U$, if $u \in Q_i$ for some i, then no j exists such that $u \in R_j$.

(iii) **Spoke paths form a tree:** for each distinct $Q_i, Q_j \in \mathcal{Q}$, if $v \in Q_i \cap Q_j$, then $Q_i[v] = Q_j[v]$.

Section 2.6.2 showed that this result does *not* hold when we account for spurious updates. We define a generalized version of the dispute reel structure, which we prove to be exactly what is needed to identify the systems that are unsafe, but only due to spurious updates. That is, we prove:

Theorem 2.7.2. *BGP, as modeled by DPVP, is safe under filtering if and only if the network has no "two-third reel":*

Definition 2.7.3. A *two-third reel* is a dispute wheel which satisfies the second and third condition of dispute reel:

(i) **Spoke and rim paths do not intersect:** for each $u \notin U$, if $u \in Q_i$ for some i, then no j exists such that $u \in R_j$.

(ii) **Spoke paths form a tree:** for each distinct $Q_i, Q_j \in \mathcal{Q}$, if $v \in Q_i \cap Q_j$, then $Q_i[v] = Q_j[v]$.

The intuition for removing the first condition in the dispute reel definition is that spurious behavior of DPVP effectively allows us to "mangle" the rim of the reel, with pivots appearing in multiple rim paths. This would prevent divergence in SPVP, since a pivot would have to stick to one of its options in between its activations, preventing its participation in other pivots' rim paths. However, with spurious announcements, the pivot may keep spuriously announcing some of its other available routes in such a pattern as to keep the oscillation going.

Proof of Theorem 2.7.2: The theorem combines the two implication directions treated separately by Lemma 2.7.1 and Lemma 2.7.2. ■

Lemma 2.7.1. *If a DPVP instance S is unsafe under filtering, it has a two-third reel.*

Lemma 2.7.2. *If a DPVP instance S contains a two-third reel, it is not safe under filtering.*

We should first formalize the concept of a **DPVP evaluation cycle**, which, given that the system is finite, will arise in any unsafe instance. An evaluation cycle $C = (\Pi, M, \sigma)$ consists of a path assignment cycle $\Pi = \pi_1 \xrightarrow{\sigma_1(M_1)} \pi_2 \xrightarrow{\sigma_2(M_2)} \ldots \xrightarrow{\sigma_k(M_k)} \pi_{k+1}$, with $\pi_1 = \pi_{k+1}$, with a matching edge activation sequence σ and spurious message sequence M. A cycle is well-formed if all the paths can evolve as specified given the message sequence, and the message sequence can continue to be sent indefinitely. That is, no spurious announcements depend on any events that predate the cycle; and every node sending messages changes its chosen path at least once in the cycle, allowing it to keep sending spurious updates, as long as the cycle loops within τ time.

Let the set $\mathsf{values}(C, u)$ be the paths that node u adopts at some point in C. Let F be the set of **fixed** nodes: those that have a fixed path assignment throughout C; and let G be the other, oscillating, nodes. We will need this technical lemma:

Lemma 2.7.3. *Suppose $P \in \mathcal{P}^v$ is adopted by node $v \in G$ in the cycle C. Then we can write $P = QR$ where the first node on path R is in G and all other nodes further on in R are in F.*

Proof. $v \in G$, and the destination 0 is in F. ■

Proof of Lemma 2.7.1: We use proof techniques similar to the ones in the SPVP safety proof [45] to show that an unsafe instance has a two-thirds reel.

Let $C = (\Pi, M, \sigma)$ be a well-formed non-trivial cycle. Let $U \subseteq G$ be the nodes that sometimes adopt a path that consists of fixed nodes. U is nonempty, since there are oscillating nodes, and applying Lemma 2.7.3 to one of them gives a node in U.

We now construct a dispute wheel. Let u_0 be a node in U. Let $Q_0 = (u_0, w_0)Q'_0$ be the path of u_0 such that $w_0 \in F$. Since any such w_0 doesn't send spurious updates in the cycle, one can verify that there is only one such Q_0, and it is the lowest ranked path in $\mathsf{values}(C, u_0)$. Let

$H_0 \in \mathsf{values}(C, u_0)$ be the highest-rank path u_0 ever adopts. We have $\lambda^{u_0}(H_0) > \lambda^{u_0}(Q_0)$. Lemma 2.7.3 lets us decompose $H_0 = R_0 Q_1$, with $Q_1 = (u_1, w_1)Q_1'$, $u_1 \in U$, and R_0 being non-empty since Q_0 is unique at u_0. We repeat this inductively to yield a sequence (u_i), which loops back to u_0 since U is finite. The nodes u_i, spokes Q_i and rims R_i form the dispute wheel.

Finally, we prove that the dispute wheel also satisfies the two-third reel conditions of Definition 2.7.3. Suppose that condition (i) is not satisfied and there exists a node $u \in Q_i \cap R_j$. Since $u \in Q_i$, u and the rest of $Q_i[u]$ lie in F. Fixed nodes can't send spurious updates, so any repeating announcement of a path via u will end with $Q[u]$ (by induction back from u to the sender of the announcement). Thus, $R_j[u]$ must be a prefix of $Q_i[u]$, requiring $R_j[u] \subseteq F$, but this contradicts $R_j[u]$ ending in $u_{j+1} \notin F$. Suppose condition (ii) is violated, and there exist spoke paths $Q_i, Q_j \in \mathcal{Q}$ and a node $v \in Q_i \cap Q_j$ such that $Q_i[v] \neq Q_j[v]$. Then v cannot be a stable node, a contradiction.

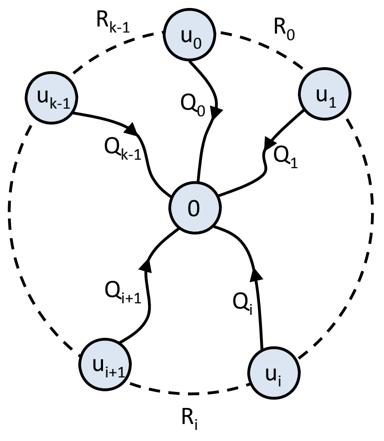
Any instance that is unsafe under filtering will thus have a two thirds reel after some routes are filtered. But if that subinstance has a two thirds reel, "unfiltering" those routes can't remove the reel, guaranteeing a two thirds reel in the original instance as well. ■

Proof of Lemma 2.7.2: Given a two-thirds reel, we first find the right parts of the system to filter out to cause the oscillation. We then define the path assignments that comprise the oscillation, and finally we show the activation sequence that allows infinite transitions between these path assignments.

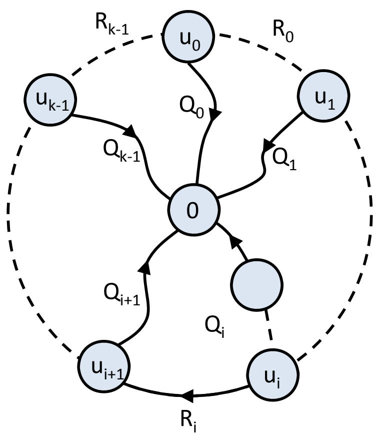
Given an SPP instance S containing a dispute wheel W, the *supporting instance* $S[W]$ is the minimal instance which contains the vertices, edges and paths of W. That is, the supporting instance is obtained by filtering all paths except the spoke and rim paths in the dispute wheel, and removing all edges and vertices outside of the dispute wheel.

The system oscillates by alternating between "all-spoke" and "one-rim" path assignments, like those in Figure 2.11. Formally, an *all-spoke* path assignment is a path assignment $\bar{\pi}$ such that $\bar{\pi}(u) = Q_i[u]$ if $u \in Q_i$, and $\bar{\pi}(u) = \epsilon$ otherwise. A *one-rim* path assignment is a path assignment $\bar{\pi}^i$ such that:

$$\bar{\pi}^i(u) = \begin{cases} Q_j[u] & \text{if } u \in Q_j, u \neq u_i \\ R_i[u]Q_{i+1} & \text{if } u \in R_i \\ \epsilon & \text{otherwise.} \end{cases}$$



(a) All-spoke path assignment $\bar{\pi}$.



(b) One-rim path assignment $\bar{\pi}^i$.

Figure 2.11: Special path assignments of a two-third reel.

Let S be the routing problem instance containing the two-third reel $R = (U, \mathcal{Q}, \mathcal{R})$ of size k. We consider the supporting instance $S[R]$ and construct a fair activation sequence with a sequence of possibly spurious messages that cause oscillations. The main idea is to alternate between the all-path assignment $\bar{\pi}$, and one-rim path assignment $\bar{\pi}^i$. Once a one-rim path assignment $\bar{\pi}^i$ is reached, the all-paths assignment is reached next, followed by the one-path assignment $\bar{\pi}^{i+1}$. This infinite sequence repeats itself, with the index $i + 1$ calculated modulo k. The paths of the nodes on the rim changes, and hence these nodes may send spurious updates.

The proof is illustrated in Figure 2.12. The black nodes are nodes that have a fixed path throughout the evaluation cycle. The paths of the white nodes change and we assume that the convergence time τ is long enough to allow these nodes to make spurious updates. Note that we *design* the oscillation with the property that every rim node oscillates; the system may also have other oscillations without this property.

We start with the empty state, $\pi(v) = \epsilon$ for all v. First, we activate the edges of each spoke path Q_i in outbound order from 0. Any non-pivot spoke node is not on the rim, and the pivots are all on separate spokes, so the filtering causes the spokes to build up the all-spoke path assignment $\bar{\pi}$. In the remainder, we will assume that, unless stated otherwise, each node $u \in R_j$ for each j always sends a spurious withdrawal (ϵ) to every neighbor. All other nodes make non-spurious announcements. Next we activate the edges on path $R_0 = (x_0 = u_0, x_1, \ldots, x_l = u_1)$ in reverse order, i.e., starting from node u_1. The node x_i spuriously announces route $R_0[x_i]Q_1$ to node x_{i-1}.

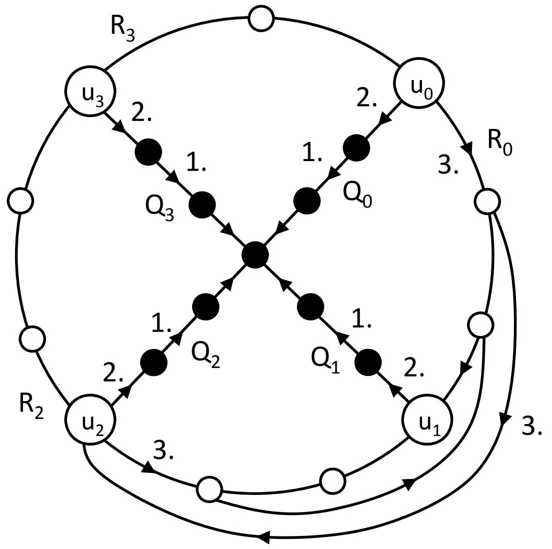


Figure 2.12: Sequence of activations that results in permanent oscillations.

Therefore, the path assignment $\bar{\pi}^0$ is reached. If all edges of each rim node are activated making the spurious withdrawal as specified before, the all-spoke path assignment is reached again. Continuing inductively, we reach assignment $\bar{\pi}^1$, and so on, yielding the infinite activation sequence. ■

Proof of Theorem 2.7.1 is now a direct corollary of Theorem 2.7.2. If there is no dispute wheel, there is no two-thirds reel, a special type of dispute wheel, and the resulting safety under filtering implies safety. ■

As should be evident from the ease with which we prove the sufficient conditions for safety in the DPVP model, spurious updates give us extra flexibility when finding oscillations, which makes proofs significantly *easier*. In the remainder of the chapter we solve the long-standing open problem of finding the necessary and sufficient conditions for safety.

2.8 The Necessary and Sufficient Conditions

We now formulate a sufficient and necessary condition for safety in the DPVP model. We show that a system is unsafe if and only if the configuration allows the existence of a particular combinatorial structure in the network. Since such a structure can serve as an easy-to-check proof that the system is unsafe in DPVP, the problem of deciding DPVP safety is in coNP, which contrasts the intractability results that show SPVP safety checking is PSPACE-complete [30]. Most remarkably, our results in Section 2.9 show that we can check a DPVP system for safety in polynomial time in most practical settings.

Informally, the network will be unsafe if and only if it admits a mapping of each node to a "selected path" and a partition of the nodes into two non-empty sets:

1. *Stable nodes* that in any infinite fair activation sequence eventually select a path consisting of only stable nodes, and, combined with all stable nodes' selected paths, comprise a consistent routing tree to the destination.

2. *Coy nodes* that are "coy" about joining that stable tree: there exists an infinite fair activation sequence in which they select a path that starts with another coy node as a next hop, preferring it over any paths that go straight to the stable tree.

We will shortly formally define this complete structure — which we dub a **CoyOTE** (for "Coy Optimum at Tree Edges") — by a set of requirements on a node partitioning and a path assignment that make one side of the partition "stable" and the other "coy", as roughly sketched above. We will show that, in a network that admits a CoyOTE structure, the stable nodes must reach a stable state under any activation sequence, while the coy nodes are capable of triggering DPVP oscillations.

Theorem 2.8.1. *An instance of DPVP oscillates if and only if it has a CoyOTE.*

To formally define the CoyOTE structure, we'll need some intermediate technical definitions:

We use $C \subset V$ to denote the set of coy nodes, and always define $S = V \setminus C$ to denote the complementary set of stable nodes. Given a path assignment Π and a coy set C, we define the set $\mathsf{stableChoices}(v, C, \Pi)$ of all paths available to v that go directly to the stable set: all $P \in \mathcal{P}^v$ such that $P = (v, u)P'$ where $u \in S$ and $P' = \pi^u$. If the set $\mathsf{stableChoices}(v, C, \Pi)$ is nonempty, let $\mathsf{bestStable}(v, C, \Pi)$ be the best such path: the unique $P \in \mathsf{stableChoices}(v, C, \Pi)$ for which $\lambda^v(P)$ is maximal. Otherwise let $\mathsf{bestStable}(v, C, \Pi)$ be ϵ.

Definition 2.8.1. The pair (Π, C) of a path assignment Π and a **coy set** $C \subset V$ (with the **stable set** defined as $S = V \setminus C$) is a CoyOTE if the following conditions hold:

(i) **Stable origin**: The origin is stable, $0 \in S$.

(ii) **Coy existence**: There are coy nodes, $C \neq \emptyset$.

(iii) **Best stable path at stable node**: A stable node is assigned its best stable path — for all $v \in S$, $\pi^v = \mathsf{bestStable}(v, C, \Pi)$.

(iv) **Coy node prefers a coy path**: a coy node is assigned a path learned from a coy node — for all $v \in C$ we have $\pi^v = (v, v_{\text{next}}, \ldots, 0)$ with $v_{\text{next}} \in C$. The assigned route must be higher ranked than v's most preferred stable route, $\lambda^v(\pi^v) > \lambda^v(\mathsf{bestStable}(v, C, \Pi))$.

(v) **Consistency with stable suffixes**: Every node's assigned path is "suffix-consistent" with (S, Π), as defined below.

Definition 2.8.2. A path π^v is **suffix-consistent** with (S, Π) if every suffix of π^v that starts with a node $s \in S$ is consistent with π^s: if $\pi^v = (v, \ldots, s, P_s, 0)$ and $s \in S$ (P_s is an arbitrary subpath), then $\pi^s = (s, P_s, 0)$.

We will need the following easy corollary of the definition:

Lemma 2.8.1. *In a CoyOTE:*

(i) If v is stable (and $\pi^v \neq \epsilon$), then $\pi^v = (v, P_v^S, 0)$, where P_v^S is a possibly-ϵ subpath containing only stable nodes.

(ii) If v is coy, then $\pi^v = (v, P_v^C, P_v^S, 0)$, where P_v^C is a non-empty subpath consisting of only coy nodes, and P_v^S is a possibly-empty subpath consisting of only stable nodes.

Proof. By induction on the length of paths assigned by Π. Assuming a node chooses a path where a coy node occurs after a stable node, by condition v there is a suffix $(s, c, \ldots, 0)$ with $s \in S$, $c \in C$, but then π^s violates condition iii. With condition iv, the rest follows inductively. ■

We now tackle the two directions of Theorem 2.8.1 separately: Lemma 2.8.2 and Lemma 2.8.4 establish that a CoyOTE is sufficient and necessary, respectively, for oscillations in DPVP.

Lemma 2.8.2. *If a DPVP instance has a CoyOTE, then the DPVP instance can oscillate.*

Proof. Given a CoyOTE (C, Π), it suffices to find an infinite fair activation sequence in which the state of the system continues changing.

Consider starting with an empty path assigned to each node (e.g., as if the destination just went online). The activation sequence starts by every stable node $v \in S$ with $\pi^v \neq \epsilon$ activating in the order of breadth-first search on the stable node tree, from the destination outward, so that each such node gets its path in π^v immediately, and never changes again.

After that, all the coy nodes $v \in C$ activate in a loop so that each one chooses the path π^v at some point. Pick some order of coy nodes, $c_{1\ldots k}$. For each c_i in order, we run 2 rounds of activations. In round 1, activate all the coy nodes on the coy prefix of the path π^{c_i} (nodes $P^C_{c_i}$ in Lemma 2.8.1), starting with the node at the edge of the stable tree and moving toward c_i, and have them announce, perhaps spuriously, their suffix of π^{c_i}. By c_i's turn, π^{c_i} will be available, so it or another coy-next-hop path will get selected. Round 2 activates all the same nodes in the same order, and has them send spurious withdrawals. With no coy-next-hop paths available after that, c_i will change its path selection. Repeat this sequence for all the c_is in a loop. Each coy node will keep changing its route, allowing it to keep sending spurious updates.

By condition iii, any stable node v with $\pi^v = \epsilon$ cannot have any paths in $\mathcal{P}^v$ available from its stable neighbors. Thus, it may only receive announcements of allowed paths from coy nodes. By having these remaining stable nodes activate after the second round of each c_i step above, they will never have any allowed paths available to them. ■

For the other direction, we need another technical lemma:

Lemma 2.8.3. *Let an oscillating node $c \in V$ in an instance of DPVP select path P at some point during the oscillation. If the path P contains node $s \in V$ and the node s does not oscillate, then node s must permanently select path $P[s, 0]$.*

Proof. Oscillating nodes only announce recently available paths. Since s will not be announcing spurious updates for longer than τ after it stops oscillating, any stale path containing node s and a suffix other than $P[s, 0]$ will not be announced anywhere for longer than $i\tau$ time after s converges and enough fair activation phases pass. ■

Lemma 2.8.4. *If a DPVP instance oscillates, it has a CoyOTE.*

Proof. Given the oscillating instance I and the corresponding activation sequence, we construct a CoyOTE (C, Π).

Let S_I be the non-oscillating nodes and C_I the oscillating nodes in the instance I. Any non-oscillating node $v \in S_I$ which permanently chooses path $(v,u)P$ learned from an oscillating node $u \in C_I$ can be made to oscillate by changing the activation sequence. Add an activation of the edge (u,v) with a spurious withdrawal, and a second activation of the edge with a (possibly spurious) announcement of path P. We keep adding nodes to C_I until no more nodes can be made to oscillate. We can set Π so that (C_I, Π) is a CoyOTE. For non-oscillating nodes $v \in S_I$, we set π^v to be the path permanently chosen. For oscillating nodes $v \in C_I$, we set π^v be the highest ranked of the path(s) that v chooses in the oscillation.

(C_I, Π) can be seen to be a CoyOTE: The origin is stable and C_I is non-empty, yielding conditions i and ii. After the above iteration, every node s remaining in S_I has a path π^s that is either ϵ or was learned from another node in S_I. Any path from s to a next hop also in S_I is continually available, so s chooses the stable path $\mathsf{bestStable}(s, C, \Pi)$ which must be available, yielding condition iii. If, for $v \in C_I$, the highest ranked path chosen infinitely often in the oscillation has a stable next hop, then, after the next hop stabilizes and τ time passes, that path must indeed become permanently available, which contradicts v continuing to oscillate, yielding condition iv. Lastly, Lemma 2.8.3 covers condition v. ■

2.9 Practical Algorithm for Safety Verification

We now consider the algorithmic question of checking safety under DPVP. We give an algorithm, "DeCoy", that greedily shrinks the candidate coy set, and reaches an empty coy set if and only if there is no CoyOTE. This algorithm is always correct, but its runtime depends on the local policies of the nodes. After analyzing the algorithm in general, we specify the large class of policies which enable it to run efficiently, which turns out to cover most BGP policies used in practice.

2.9.1 The "DeCoy" safety verification algorithm

The DeCoy algorithm greedily builds up the candidate stable set, starting with just the destination[4]:

[4]And possibly some nodes that are permanently disconnected from the destination, which are stable and retain the empty route.

```
initialize $S = \{0\} \cup \{\text{nodes not connected to } 0\}$; $C = V \setminus S$
while there exists a $v \in C$ such that:
  1. $v$ has a neighbor in $S$, and
  2. there is no path $P \in \mathcal{P}^v$ which is both:
       (a) preferred by $v$ over bestStable$(v, C, \Pi)$, and
       (b) suffix-consistent with $(S, \Pi)$
 do
   move $v$ from $C$ to $S$; set $\pi^v$ = bestStable$(v, C, \Pi)$
   for any $v \in C$ with no paths in $\mathcal{P}^v$ that are suffix-consistent with $(S, \Pi)$ do
     move $v$ from $C$ to $S$; leave $\pi^v = \epsilon$
if $C$ is empty then return "safe"
else return "unsafe: $(C, \Pi)$ is a CoyOTE"
```

Theorem 2.9.1. *A DPVP instance is safe if and only if DeCoy terminates with all nodes in the stable set S.*

Proof. **DeCoy terminates with $S = V \Rightarrow$ safety**:

We only need to show that if DeCoy adds vertex $v \in V$ to the stable set S and assigns it path π^v, then for any infinite fair activation sequence in DPVP, node v must permanently choose path π^v. We do so by induction. The claim is true for nodes added to the stable set at beginning, since these nodes in the DPVP model always permanently select the empty path ϵ. Next, let's assume that the claim holds for all nodes in set S after a particular pass through the while loop.

Consider node v that is added to S and assigned path π^v = bestStable(v, C, Π) on the next iteration of the while loop, outside the for loop. By contradiction, assume that there exists some infinite fair activation sequence where for any step in the sequence t there exists step $T > t$ in which node v chooses path $P \neq \pi^v$. The case of $\lambda^v(P) > \lambda^v(\pi^v)$ is precluded by the conditions 2(a-b) of the while loop, since, from the induction over the while loop, all nodes already in S will stabilize, and thus any announcements using paths not suffix-consistent with (S, Π) will disappear from the system within $n\tau$ time after that. We also cannot have $\lambda^v(P) < \lambda^v(\pi^v)$: since π^v will always be available to v after τ time after the next hop stabilizes, a less preferred path cannot be chosen after that. Lastly, since the next hop of π^v, more than τ time after stabilization, won't announce any other paths, P must have a different next hop, which precludes the possibility of $\lambda^v(P) = \lambda^v(\pi^v)$, by the strictness constraint of SPP.

Similarly, the claim holds for any node v assigned to S with an empty π^v in the for loop. We know that $n\tau$ time after S stabilizes, no paths in $\mathcal{P}^v$ will be announced anywhere, and thus v will be forced to remain without a route after that.

DeCoy terminates with non-empty $C \Rightarrow$ a CoyOTE exists:

Observe that after DeCoy finishes, the π^v for every node $v \in C$ is left unspecified. We complete the specification of the path assignment Π by setting π^v for every $v \in C$ to the highest ranked path in $\mathcal{P}^v$ that is suffix-consistent with (S, Π). This must be a non-empty path, ranked strictly higher than $\mathsf{bestStable}(v, C, \Pi)$. Otherwise, if $\mathsf{bestStable}(v, C, \Pi)$ is non-empty and there is no such path, v would satisfy the while loop's condition and would be added to the stable set. Alternatively, if the resulting π^v is ϵ — or, equivalently, if $\mathsf{bestStable}(v, C, \Pi)$ is empty and there is no allowed suffix-consistent path higher-ranked than ϵ — then there is no allowed suffix-consistent path at all. In that case, v would have been stabilized inside the for loop after its last allowed path stopped being suffix-consistent. This insures CoyOTE condition (iv).

CoyOTE conditions (i), (ii), (v) are immediate from the construction if $C \neq \emptyset$.

Note that throughout the algorithm, the set S grows, and path assignments are never changed after being set for the first time. As more nodes are stabilized and the number of constraints required for suffix-consistency increases, the set of candidate paths that are suffix-consistent with (S, Π) monotonically shrinks.

For a stable node s with a non-empty path, consider the stable next hop s' of the path $\mathsf{bestStable}(s, C, \Pi)$, using the *final* values of C, S, and Π. Suppose s' was added to the stable set after s, so it would have been still in C when s was being added. Since $\pi^{s'}$ was suffix-consistent when it was assigned to s', by monotonicity, it was suffix-consistent earlier, too. But, since that path became $\mathsf{bestStable}(s, C, \Pi)$ by the end of the algorithm, it must have ranked higher than whatever path s was assigned as its current $\mathsf{bestStable}(s, C, \Pi)$, using the running value of C and Π when s was stabilized (when s' wasn't stable yet). Thus, s''s path would have negated condition (2) of the while loop, preventing s from being stabilized. Thus, s' was stabilized before s, and, by the time s was stabilized, we had $\mathsf{bestStable}(s, C, \Pi) = (s, \pi^{s'})$, confirming CoyOTE condition (iii).

For a stable node s with an empty path, by the monotonic shrinking of the set of suffix-consistent paths, if there were no allowed suffix-consistent paths when π^s was assigned ϵ, none would have appeared since, guaranteeing that $\mathsf{bestStable}(s, C, \Pi)$ would be ϵ at termination, too, and confirming condition (iii) here, too. ■

The efficiency of the DeCoy algorithm turns out to be intricately linked to the tractability of *optimizing over policies*, including both router preferences (embodying, e.g., BGP local preference settings), and allowed path sets (embodying, e.g., filtering rules). We first define the fully general requirements on *optimizable policies* that allow DeCoy to run efficiently, and then, in Section 2.9.2, discuss the wide range of realistic policies that fit into this framework.

Definition 2.9.1. We say that a node v implements *optimizable policy* if v's permitted path set $\mathcal{P}^v$ and ranking function λ^v allow this node to find, in polynomial time, the highest-ranked path $P \in \mathcal{P}^v$ that is suffix-consistent with a given stable tree (S, Π) (which, being a tree, is itself suffix-consistent).

Theorem 2.9.2. *If all nodes implement optimizable policy, DeCoy runs in polynomial time.*

Proof. We can search through the $O(n)$ paths in stableChoices(v, C, Π) by brute force, to find the bestStable path. The optimizable policy constraint then lets us check whether the most preferred suffix-consistent allowed path is stable — which is needed to evaluate the second condition of the while loop. The for loop condition is checked by seeing if the optimization returns the empty path as the best suffix-consistent allowed path. The other steps are always efficient.

Nodes are stabilized at most once, so there are a total of at most $|V|$ iterations involved in both loops. Thus, even a brute-force implementation will finish in $O(|V|f(n)^2)$ time, where the efficient policy evaluation runs in time $f(n)$ as a function of the input. With most specific policies, we expect the best runtime to be substantially faster. ■

DeCoy is thus polynomial-time as long as a node's policy lets it efficiently figure out "what route do I most prefer, given what the rest of the world has permitted me?" We assert that a policy which does *not* allow a tractable answer to this question is somewhat strange. In a heuristic sense, that would entail the router "not knowing what is best for it", even when shown the full network structure and available paths, and instead awaiting whatever the notoriously unpredictable network dynamics give it, with no easily-computable goal in mind.

2.9.2 Verifying Safety in Practice

We now present a broad range of realistic policies that turn out to be optimizable. The BGP policies actually used by ASes are varied and complex, but fairly well understood [19]. Typically, a router implements a decision process that first applies import policies to filter out some routes received, then applies a local ranking to select the most preferred available route, and then applies export policies to determine whom to announce it to.

For the purposes of DPVP analysis, we distinguish two types of policies: (1) **preference** policies prefer some routes less than others, but don't preclude spurious announcements of such routes (modeled with the λ local ranking function); and (2) **filtering** policies which remove a route at such an early point, or at such a low level of the router, precluding even spurious updates involving that route (modeled with the permitted paths set $\mathcal{P}^v$). BGP import/export constraints will often be filtering policies. But BGP policies often thought of as filters, such as "avoidance policies" (like "avoid paths through country X") , are often actually just a depreferencing step in the preference function, without any clear mechanism forbidding spurious announcements of such a route.

Below, we present examples of common policies classified as either filters or preferences in a particular way. But we prefer that the reader remains agnostic about what kinds of policies to file under which of these two categories. This decision depends on the details of the mechanisms generating spurious updates *or* on the algorithm executor's decision to seek a more conservative/robust notion of safety: safety under a wider range of possible (mis)behaviors. Moving a policy from "filtering" status to "preference" status can only allow more spurious update possibilities, yielding a more robust notion of "safety", which we think is natural to seek. Remarkably, in all the cases we consider below, moving policies we consider as filtering policies to the preference stage would keep all the policies efficient.

A filtering policy requires that a path not be used even in absence of any other options, so we expect that each such black-and-white policy has a strong incentive behind it, typically produced by economic or security constraints. In most typical scenarios, a non-stub AS is only likely to filter based on some combination of the next and previous hop on the path in question (as needed, e.g., for Gao-Rexford economic constraints [38]), and the destination of the packet (as needed for blackholing, motivated by security or politics).

In particular, we admit as **"typical filtering policies"** any disjunction of polynomially many filtering rules that filter paths that are:

(a) learned from neighbor u and announced to neighbor v, or

(b) learned from neighbor u, or

(c) announced to neighbor v, or

(d) violating Gao-Rexford constraints, or

(e) leading to a blacklisted destination.

Rule (a) is quite general and rules (b) through (d) can be implemented by applying rule (a) polynomially many times.

This notably excludes filtering paths that go *through* another transit provider somewhere further in the path, but not on the next hop. The latter type of filtering, for security or for censorship, is unlikely to be effective, due to path diversity in the Internet.

Most **preference policies** observed in practice are covered by the following (in an *arbitrary* order):

(i) split the space of paths into some polynomial number of "categories", each of which disallows the use of some subset of nodes and/or edges, and prefer "categories" in a particular order, or penalize each category differently.

(ii) within each category, prefer routes based on a *stratified shortest path* metric with a logarithmic number of strata, and polynomially many preference levels within each stratum.

The inner category of stratified shortest path metrics, introduced by Griffin [42], allows a rich space of preferences based on a general algebraic semi-ring structure that allows us to "lexicographically stack" several precedence levels of preference constraints that are algebraically isomorphic to a shortest path problem. This allows many common policies beyond just basic shortest-hop-length. For instance, such a policy can be as complicated as:

1. prefer **customer routes**; then
2. penalize paths **passing through country** X by p_x, and through countries Y and Z, by p_{yz}; then
3. pick a route using the **shortest hop path**;
4. of those routes, pick a route that that minimizes some **linear combination** of:

   (a) **number of "undesirable" nodes it passes** (e.g. low-performing nodes)

   (b) **average data-plane packet loss rate over the path**;

5. then pick the one w/the **lexicographically 1st next hop**.

The number of strata in this unusually complicated example is 4, and no more than a small constant number of strata should ever be needed for realistic policies. Indeed, Griffin's work on such policies in [42] restricts the consideration to $\leq$ 3-stratum policies.

The outer splitting of policies into categories with disallowed nodes/edges allows a yet wider range of policies that don't fit neatly into the stratified shortest path framework. E.g., it allows a fixed penalty for path visiting some set of nodes (rather than counting how many nodes are visited), or it allows any route going through some next-hops (e.g. Gao-Rexford customers) to be strictly preferred over other next-hops (Gao-Rexford providers). This in particular allows the full implementation of any Gao-Rexford business policies, in addition to complex stratified shortest path policies within each business relationship class. The only meaningful restriction that we know of is that the number of categories be polynomial. If there is a list of more than a logarithmic number of "bonuses", each assigning an additive penalty to a path with a particular non-stratified-shortest-path-based feature, that would generate a superpolynomial number of categories. While such a policy *could* be implemented with using common router policy description languages, it is not clear that operators would do so, since it creates an extremely fine-grained distinction between superpolynomially many *combinations of bonuses*, which seems unnecessary in any context we know of. Together, these two levels of preference policy correspond closely to the decision processes described in [19].

Theorem 2.9.3. *Any combination of "typical" filtering and preference policy classes above constitutes optimizable policy.*

Proof. We need to demonstrate a polynomial-time algorithm to find a node's most preferred suffix-consistent path that would not be filtered by the nodes on it.

We start by modifying the graph to ensure suffix-consistency and compliance with filtering policies, and then run a modified version of the Bellman-Ford algorithm, applied separately to each of the polynomially many categories of the preference policy.

To ensure suffix-consistency with stable set S, convert into one-way edges (1) all edges from $V \setminus S$ to S, in that direction; and (2) the edges of the routing tree, formed by the (suffix-consistent) path assignments of stable nodes with non-empty paths toward d. Then remove edges between two stable nodes that are not used in the routing tree. Any permitted path that is consistent with these changes is suffix-consistent, since once it hits a stable node, it can do nothing but follow the stable tree to the destination.

Then remove any nodes that filter paths to the destination in question.

Then, separately for each category of the preference function, remove the disallowed edges or nodes from the network graph.

We then set a function λ^v by transforming each stratum of the stratified-shortest-path preferences into the isomorphic shortest path problem, and combining all the strata into a single preference value that follows the lexicographic ordering. For example, a 2-stratum policy with top stratum ranging from 1 to 9, and a second stratum ranging from 1 to 99 would be combined into a λ^v ranging from 101 to 999, with the first digit corresponding to the first stratum, and the last two digits corresponding to the second stratum. The combined λ^v still has a polynomial range, due to the constraints on the strata.

We then run a Bellman-Ford variant to compute paths to destination minimizing λ^v. The only additional constraint is that we must account for node filtering policies, which we do at each node as the Bellman-Ford wavefront propagates. Specifically, we maintain not just the shortest path at each node, but rather each node tracks for each next-hop-and-previous-hop pair (u, v) separately the shortest path that can be imported from u and exported to v.

The correctness argument follows from Bellman-Ford. The runtime scales linearly with the complexity of the preference policy evaluation, and increases only polynomially from maintaining shortest path options for each (u, v). ■

2.10 Hardness of Some Safety Verification Formulations

We will prove that in the most general settings it is NP complete to verify whether a DPVP instance is safe. The intractability relies on route preferences that are typically not used in practice, such as preference for longer routes over shorter ones, preference for routes of a particular fixed length, or filtering policies applied to transit traffic.

BGP permits the use of regular expressions to specify routing policies. In order to show that verifying the safety of BGP becomes NP complete when regular expressions are used to make the problem inputs (which include routing policies) more compact, we define regular expressions that are allowed to contain the "." metacharacter representing any one autonomous system. For the purpose of our proof, any other definition of regular expressions that allows the "." metacharacter is equivalent.

A *regular expression* is a sequence of metacharacters "." and node numbers. The metacharacter "." represents any node number. For example, the regular expression 3..0 matches both paths $3, 7, 4, 0$ and $3, 1, 9, 0$. When regular expressions are allowed, ranking functions Λ become *compact ranking functions*: they provide a ranking of regular expressions rather than individual paths. Path P_1 is preferred to path P_2 if the highest ranked regular expression that matches path P_1 is preferred to the highest ranked regular expression that matches path P_2. If two paths match the same regular expression then let the lexicographically smaller path be preferred. If a path does not match any regular expression it is not permitted.

Theorem 2.10.1. *The problem SAFE-REGEXP-DPVP of determining the safety of an instance of DPVP with compact ranking functions is NP-complete.*

Proof. The problem is in NP because if we are given a CoyOTE structure (Π, C), we can check its validity by verifying that the conditions in Definition 2.8.1 are satisfied. The path ranking function λ^v is polynomial-time computable and $\lambda^v(P)$ can be evaluated in polynomial time for any v and P. Hence we can also evaluate $\mathsf{bestStable}(v, C, \Pi)$ in polynomial time, and check each condition of Definition 2.8.1.

The rest of the proof uses a reduction from the Hamiltonian cycle problem, which is one of Karp's 21 NP-complete problems [55]. An instance of the Hamiltonian cycle problem is an undirected graph $G = (V, E)$. The problem asks if there exists a Hamiltonian cycle (a closed loop) that visits each distinct node exactly once.

Suppose we are given an instance I of the Hamiltonian cycle problem. We now construct an instance of the SAFE-REGEXP-DPVP problem D that is not safe if and only if I has a Hamiltonian cycle.

We construct the graph and route rankings of the SAFE-REGEXP-DPVP problem as illus-

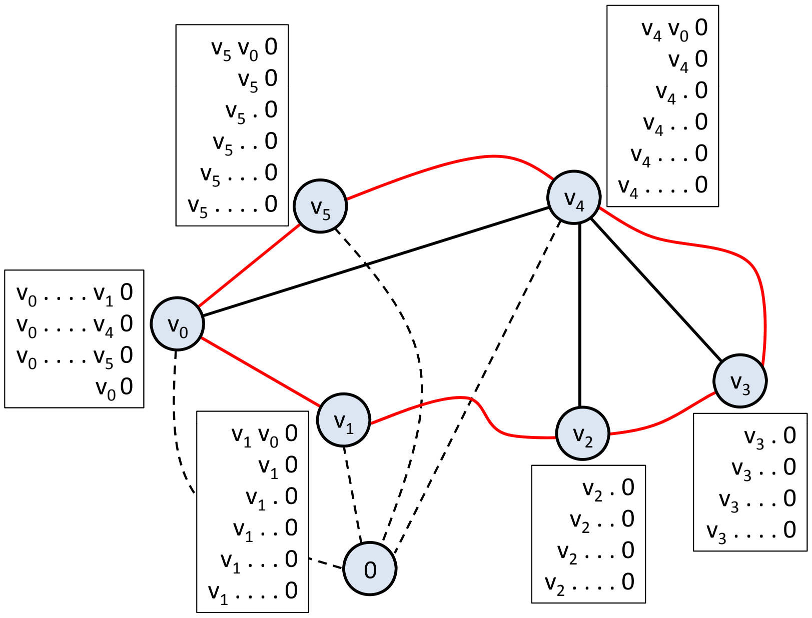


Figure 2.13: Example of construction. Solid lines represent edges in the original graph.

trated in Figure 2.13. Let $G = (V, E)$ be the graph in the original instance I, let $n = |V|$, and let $T = \{v_i : (v_0, v_i) \in E\}$ denote the set of neighbors of node $v_0 \in V$. We modify the graph G by adding vertex 0 representing the origin. We also add edge $(v_i, 0)$ for each $v_i \in \{T \cup v_0\}$. Finally, we specify the compact ranking functions. Node v_0 allows route $v_0 ... v_i 0$ for each $v_i \in T$ where the number of the dots is such that the length of the route is n. Node v_0 also allows $v_0 0$. Each node $v_i \in T$ allows route $v_i v_0 0$, and routes $v_i ... 0$ where the number of dots is such that the length of the route is between 1 and $n - 1$, and shorter routes are strictly preferred to longer ones. Finally, each node $v_i \in V - \{T \cup 0\}$ allows routes $v_i ... 0$ where the number of dots is such that the length of the route is between 2 and $n - 1$, and shorter routes are strictly preferred to longer ones.

We now show that if I contains a Hamiltonian cycle then D contains a CoyOTE. Let the Hamiltonian cycle in I be $(v_0, v_1, ..., v_{n-1}, v_0)$. We construct a CoyOTE (Π, C). Let $C = \{v_0, v_1, v_2, ..., v_{n-1}\}$. Let $\pi^{v_0} = v_0 v_1 v_2 ... v_{n-1} 0$, for each $v_i \in T$ let $\pi^{v_i} = v_i v_0 0$, and for all $v_j \neq 0$, $v_j \notin T$ let $\pi^{v_j} = v_j v_{j+1} ... v_{n-1} 0$. It is easy to verify that the conditions of Definition 2.8.1 are satisfied and hence (Π, C) is a CoyOTE.

It remains to show that if D contains some CoyOTE (Π, C), then I contains a Hamiltonian cycle. First we show by contradiction that $v_0 \in C$. Assume $v_0 \notin C$. There are two cases, either

$\pi^{v_0} = v_0...v_i0$ or $\pi^{v_0} = v_00$. If $\pi^{v_0} = v_0...v_i0$ then by Lemma 2.8.1 each $v \in \pi^{v_0}$ is stable, and hence each $v \in V$ is stable. This contradicts the fact that (Π, C) is a CoyOTE because condition ii of Definition 2.8.1 is not satisfied. If $\pi^{v_0} = v_00$ we reach a contradiction as follows. Nodes $v_i \in T$ are stable because their highest ranked path v_iv_00 is permanently available. It is easy to verify that the remaining nodes $v_i \in V - \{T \cup 0\}$ are also stable because they strictly prefer shorter paths over longer ones, reaching the same contradiction as before.

Since $v_0 \in C$ we must have $\pi^{v_0} \neq v_00$ because $0 \notin C$. Then $\pi^{v_0} = v_0...v_i0$ and we know that π^{v_0} is a valid loop free path in the graph $G = (V, E)$ of instance I. $v_i \in T$ and hence $(v_i, v_0) \in E$. Therefore $v_0...v_iv_0$ is the sought Hamiltonian cycle in the original graph G. ■

Next we show that if we allow each node $v \in V$ to filter routes that contain a node belonging to an arbitrary subset of nodes A^v, safety verification becomes NP-complete. We extend the route filtering and route preference rules of Section 2.9.2 to include this filtering step and show NP-completeness.

Theorem 2.10.2. *The problem SAFE-DECISION-DPVP of determining the safety of an instance of DPVP where node $v \in V$ filters routes containing any node $a \in A^v$ and otherwise implements the filtering and preference rules of Section 2.9.2 is NP-complete.*

Proof. The problem is in NP by similar observation as in the proof of Theorem 2.10.1. To show NP completeness we reduce from 3-SAT [55].

We transform a 3-CNF formula of 3-SAT into an instance of SAFE-DECISION-DPVP that is safe if and only if the formula is not satisfiable. This transformation is depicted in Figure 2.14. For each clause the graph contains the following gadget. If the ith clause in the formula is $(x \vee y \vee z)$, the graph contains nodes c_i and c_i' connected through three intermediate nodes labeled x, y and z. The graph also contains nodes 0, 1 and 2. Node 2 is connected to c_1, c_i' is connected to c_{i+1}, and the last node c_i' is connected to 0.

Node 1 prefers routes learned from node 2 and node 2 prefers routes learned from node 1. Nodes labeled with a variable x avoid every other node labeled with its negation $\neg x$. Node c_1 also avoids node 2. The original formula can be satisfied if and only if node 2 can use a path to the origin 0 passing through the nodes corresponding to the variables that satisfy each clause. If and only if this path exists nodes 1 and 2 are in dispute and the DPVP instance is not safe. ■

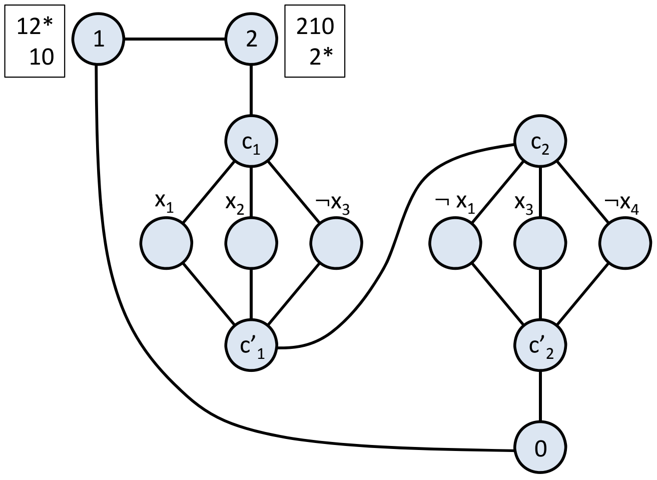


Figure 2.14: Example of construction for a 3-CNF formula $(x_1 \vee x_2 \vee \neg x_3) \wedge (\neg x_1 \vee x_3 \vee \neg x_4)$. Nodes labeled x_i avoid nodes labeled $\neg x_i$.

2.11 Summary

We have introduced DPVP, a new dynamic model of BGP, which captures spurious announcements (those of recently seen, but not most preferred, routes) that can result from a number of BGP implementation details. DPVP proved to be a powerful tool with favorable properties that allowed us to prove a necessary and sufficient condition for DPVP convergence, a question that remained elusive with earlier models of BGP. Furthermore, our DeCoy algorithm for verifying BGP safety under DPVP runs in polynomial time for a rich group of policies that are representative of the configurations used in practice. This resolves an important question, of possible practical interest to both researchers and network operators who need to verify the correctness of their configurations. Designing a distributed privacy-preserving version of the DeCoy algorithm that doesn't require ASes to reveal their routing policies is the subject of our ongoing investigation.

To model a variety of BGP implementation phenomena, DPVP admittedly allows a very broad variety of spurious message patterns. Any one router implementation would likely not be capable of emitting every spurious message pattern that DPVP would allow. So what does it mean to require that a system be safe under DPVP? We believe that there are three valid, potentially practical approaches to thinking about DPVP safety, depending on what we assume about the Internet:

1. Throughout this chapter, we pursued the literal interpretation of DPVP: the mechanisms in Sec. 2.1.1 can, under some circumstances, allow for any DPVP-modeled sequence of spurious updates. Without visibility into the implementation details of every proprietary router architecture in the world, safety under the full range of DPVP spurious updates gives us a strong guarantee on convergence across a wide range of router behaviors, and, we believe, is future-proof against the likely scope of future tweaks to BGP behavior.

2. Between SPVP and DPVP, there is a range of models allowing some spurious update patterns and not others. By restricting the spurious patterns that you consider possible, more systems can become safe, so DPVP becomes a "sufficient model" for safety. I.e., for any intermediate model with restrictions on spurious updates, DPVP safety constitutes a broad class of systems that are safe, and are, with optimizable policies, efficiently checkable by DeCoy. This is useful even for the corner case of SPVP proper: *A system with no dispute wheels has no CoyOTE, but there are also systems with no CoyOTE that* do *have dispute wheels. All these are DPVP-safe, and thus SPVP-safe.* So, even with no spurious updates, this work expands the range of systems we can efficiently label safe.

3. Like SPVP, DPVP can also be used as a coarse model of AS-level behavior. As long as we assume that the whole AS has a consistent ranking of paths, we can treat the AS, instead of the router, as one DPVP node. As with SPVP, this has the advantage of not requiring intra-AS knowledge. Furthermore, DPVP's spurious updates provide a much more reasonable scope of AS-level behavior, as any DPVP update pattern can easily arise from, e.g., iBGP behavior within the AS. Except for intra-AS differences in path ranking, DPVP otherwise seems to also encompass a wide and possibly complete range of BGP behaviors of an AS "node", and thus gives us a useful tool for coarse AS-level modeling.

We believe that each of these three frameworks is worthy of further exploration in its own right.

Chapter 3

Preventing Address Prefix Hijacks

3.1 Introduction

Today's Internet routing system is extremely vulnerable to attacks where adversarial networks announce routes for address blocks they do not own. In fact, "hijacking" another network's IP prefix is so easy that it often happens by accident [14–16,65]. The consequences of prefix hijacking, and other forms of bogus routes announcements, are serious because the packets destined to the victim prefix are instead delivered to the adversary, who may drop the traffic, impersonate the destination, modify the payload, or snoop on the communication. For example, during the infamous "AS 7007" incident, a significant fraction of all Internet traffic was mistakenly directed to a small ISP for several hours [15].

The best way to defend against prefix hijacking is the subject of much debate. The role of secure routing protocols, in particular, has received considerable attention. The debate has been dominated by a "purist" philosophy that advocates the ubiquitous deployment of a secure version of the Border Gateway Protocol (BGP). The purist approach seems natural, if not mandatory, since BGP is the glue that holds the disparate parts of the Internet together. Purist solutions are advocated in public forums, such as the RPSEC working group of the IETF [6] and the North American Network Operators Group [7]. In fact, the debate focuses primarily on *which* secure routing protocol should be adopted (e.g., S-BGP or soBGP) [8], rather than *whether* a single solution should prevail. The Internet policy community has also discussed the possibility that the

U.S. government might mandate S-BGP deployment [5].

Although ensuring that routing-protocol messages are authorized is clearly useful, the purist approach is problematic for both technical and economic reasons:

- Ubiquitous deployment of a secure routing protocol requires the cooperation of more than 35,000 autonomous systems (ASes). The large number of ASes prevents market forces from driving deployment, and government intervention may be both hard to realize (due to the global nature of the Internet) and undesirable (since it may stifle innovation).

- Smaller groups of like-minded ASes are much more likely to have aligned incentives that enable a partial deployment of a security solution. In a small group, one large company may have sufficient incentive to finance the participation of other members, or all of the ASes in the group (say, of large backbone providers) may decide to share the cost for their mutual gain.

- Groups benefit from deploying customized security solutions. No one interdomain security solution satisfies all of the security objectives, and, therefore, different groups may want to strike different trade-offs, based on their customer requirements and deployment costs.

In this chapter, we argue that small groups of cooperating ASes should be the starting point for securing interdomain communication.

Interdomain communication needs to be protected against attacks on availability, confidentiality, and integrity. Ultimately, ensuring confidentiality and integrity requires end-to-end mechanisms, such as end-host encryption and authentication. As such, in designing and evaluating secure routing techniques, we focus primarily on improving end-to-end *availability*, though some of our solutions also improve the confidentiality and integrity of communication. Rather than *guaranteeing* availability—something that is inherently difficult to do, even for full deployments of S-BGP—we focus on significantly raising the bar for the adversaries to disrupt the delivery of traffic to the group members. For example, we would like to limit the number of places where an adversary can launch a successful attack, or require several colluding adversaries before an attack can succeed.

The problem is challenging because the non-participating ASes—who make no effort to detect or avoid the routes announced by the adversaries—outnumber the group members by several orders of magnitude. This imposes several serious constraints on the space of solutions. First, the group members must use the conventional (insecure) version of BGP to exchange routing information with the non-participants, meaning the participants cannot completely upgrade to a new, secure protocol. Second, a non-participant may unwittingly propagate an adversary's route announcement to the group members, reducing the likelihood the participants learn *any* valid routes. Third, the traffic exchanged between group members often traverses non-participating ASes, who may unintentionally direct traffic toward the adversaries.

In this chapter, we argue that the proposed secure variants of BGP are not equipped to overcome these obstacles. On the surface, secure origin BGP (soBGP) [101] is the most promising starting point for a partial deployment, since it is backwards compatible with BGP. Rather than authenticating the BGP messages themselves, soBGP has the routers verify the contents of BGP announcements against a *registry*, populated with information about prefix ownership and the AS topology. However, our evaluation under realistic AS-level topologies shows that small-scale deployment of the registry does not offer significant security gains. The group members can improve availability by applying more accurate techniques for detecting invalid routes, such as control-plane anomaly detectors or data-plane probing techniques. However, our experiments show that even a *perfect* detector would not make small groups effective at circumventing the adversary, even if several large ISPs participate in the group.

Our experiments suggest that to be effective, the members of the group must take two additional actions. First, they must cooperate to expose additional path diversity, to ensure that they have valid routes to the destination. Second, they must be proactive in inducing non-participants to select valid routes. In this chapter, we present the design and evaluation of two novel mechanisms that the group can use to achieve these goals:

- **Secure overlay routing**: To circumvent the adversaries, the group members form a secure overlay network we call an SBone. In contrast to conventional overlays, an SBone connects *networks* rather than end hosts, collects path-quality measurements that are robust to adversaries, and avoids mapping virtual links on to compromised paths through the Internet.

- **Hijacking the hijacker**: To prevent non-participants from directing traffic toward the adversaries, all participating ASes originate BGP announcements for the prefixes the group wants to protect, and then forward the traffic over the secure overlay to the legitimate destination. “Shouting” the group’s prefixes substantially improves availability, in exchange for a small increase in routing-table size and path lengths.

Our experiments show that these two techniques, combined with accurate detectors and the support of a few large ISPs, allow a small group (e.g., of 5 to 10 ASes) to achieve remarkable security gains at reasonable cost.

To quantify the effectiveness of small groups, we perform extensive simulations on a snapshot of the Internet’s AS-level topology, annotated with the inferred business relationships between neighboring ASes. Simulation is necessary for an accurate evaluation because the composition of the group, the location of the adversaries, the connectivity between ASes, and the routing policies all have a profound influence on whether ASes learn and select legitimate routes. To date, synthetic models that accurately capture both the Internet’s structure and BGP routing policies remain elusive, leading us to simulate security solutions on the existing Internet topology, rather than an abstract model. To understand the influence of group composition, we consider several models of group formation, including random group memberships and participation by ASes based on their node degree. In practice, we envision that groups of ASes will form based on shared incentives or the desire of large ISPs to offer enhanced security as a value-added service.

The remainder of the chapter is organized as follows. The next section expands on our economic argument that small groups should form the basis of secure interdomain communication [11]. Section 3.3 presents a brief overview of prefix-hijacking attacks. Section 3.4 shows that small-scale deployments of soBGP are not very effective, and Section 3.5 shows that even perfect techniques for detecting invalid routes are not sufficient. In Section 3.6, we show how to improve availability for communication *between the participating ASes* through secure overlay routing. Then, in Section 3.7, we show how to coax the *non-participating ASes* into directing traffic towards the group members. Section 3.8 elaborates on how our solutions defend against sub-prefix hijacking attacks. Section 3.9 discusses how a small group of like-minded ASes should deploy our solutions in practice. Section 3.10 presents related work, and Section 3.11 concludes the chapter with a discussion of future research directions.

3.2 Economic Case for Small Groups

Here we motivate our design decision to support the formation of multiple coexisting small groups as a basis for providing secure interdomain communication. First, we argue that a *purist* solution that requires the ubiquitous deployment of a secure routing protocol prevents market forces from driving adoption. Then, we argue for a *pluralist* approach that supports customized security solutions for groups of various sizes and is consistent with market forces.

3.2.1 Economics of Groups and Goods

Secure interdomain communication requires action in *groups*. Although there are techniques that an AS acting alone can use to reduce the likelihood of attacks (such as applying protective filters to routing protocol messages and data packets), these techniques cannot ensure confidentiality, integrity, and availability for interdomain communication. Symmetric encryption instead, the simplest technique to ensure confidentiality, requires bilateral cooperation to establish a security association and encrypt/decrypt the data. The formation of *groups* of ASes is, therefore, essential for interdomain communication security.

Because the ASes in the commercial Internet are independent, payoff-maximizing entities, it is important to consider the economic incentives of individual ASes to join groups that provide interdomain communication security services. From this perspective, interdomain communication security is an economic *good* that the group provides to its members by deploying common security mechanisms. Depending on the interaction among the group members goods are, in general, classified as (1) purely public, (2) purely private, and (3) impurely public, with different economic implications. Pure public goods are *non-rival*, i.e., consumption of the good by one member does not diminish the availability of the good to other members, and *non-excludable*, i.e., the privilege of consumption of the good is unrestricted. An example of a pure public good is public television broadcasting. In contrast, pure private goods are rival and excludable, for example, recorded music sold in music stores. Impure public goods are partially rival or partially excludable, such as cable television broadcasting.

The appropriate classification of a good depends on the group's incentive structure for production and consumption. In fact, technological innovations can transform a good from one class to another. For example, encryption changed television broadcasting from a pure public good to an impure public good. As another example, peer-to-peer file-sharing applications are rapidly transforming recorded music from a pure private good to a pure public good. The rest of this section argues, first, that the purist view treats secure interdomain communication as a pure public good, which prevents market forces from driving adoption, and, second, that a pluralist approach in which smaller groups coexist better matches the economic incentives of those groups.

3.2.2 Purism is not Economically Viable

The ubiquitous deployment of a secure routing protocol that purism requires is unappealing because it implies *non-excludability.* Consider, for example, an exclusion mechanism based on fees. The option of charging a fee to prospective customer networks for connecting them to your secure routing protocol, implies the possibility of networks that decline to pay the fee, therefore, leading to partial, non-ubiquitous deployment of the protocol. I.e., the option of receiving an economic gain through the deployment of a secure routing protocol is inconsistent with the ubiquity requirement of purism. In the absence of other sources of revenue (e.g., advertising), non-excludability leads to *market failure*, i.e., no supply of the good, or a level of provision that is grossly inefficient. This is the situation today, where no secure interdomain routing protocol is deployed, despite a pressing need for better security.

Avoiding market failure under non-excludability typically requires government intervention, such as regulation [24]. However, resorting to regulatory action to mandate the ubiquitous deployment of a secure routing protocol is unnecessary, and in fact may stifle the creation and deployment of superior alternatives. Instead, we believe it is possible for market forces to drive the deployment of security mechanisms, including the existing and novel secure routing protocols, just not based on the purist view. In the rest of this section, we advocate *pluralism*, i.e., the coexistence of multiple groups of various sizes, and discuss market-based incentive structures for secure communication.

3.2.3 Pluralism is Incentive Compatible

As mentioned earlier, whether a group will form to counteract a threat will ultimately depend on the economic incentives of individual ASes to join. The theory of collective action [74] argues that small and medium-sized groups are more effective in providing public goods than large ones. In a small group, one large member may have sufficient incentive to provide the good by himself, essentially financing the participation of the other members. For example, a large corporation may finance the deployment of encryption devices at smaller business partners for business-to-business transactions. In a medium-sized group, a good can be provided by strategic interaction and bargaining. For example, large backbone providers may form a coalition to deploy a secure routing protocol to protect their customers.

Accommodating independent variable-sized groups, instead of mandating the formation of a single large group that purism advocates, would enable market forces alone to drive the provision of communication-security goods, in two main ways. First, as noted above, a small or medium-sized group may provide a pure public good based solely on alignment of incentives or bargaining. Second, exclusion mechanisms can be leveraged to provide goods as impurely public or purely private. For example, a coalition of networks that had deployed a secure routing protocol may charge non-member networks to use its routes.

In the rest of this paper, we focus our attention on techniques that small groups can employ to secure interdomain communication. Before discussing these techniques we provide in the next section background material on how an adversary can attack interdomain routing.

3.3 Prefix Hijacking

The nature of the BGP protocol makes it vulnerable to attacks. Each AS may legitimately announce or *originate* in BGP one or more destination address blocks, which are called *prefixes*. Each router is allowed to select and announce a single *best* path per prefix. A BGP-speaking router trusts that the information its neighbors propagate has been generated in a legitimate fashion. This assumption of trust can be exploited by adversaries, who may use BGP attacks to breach the availability, confidentiality, and integrity of interdomain communication. In addition, as each BGP speaker announces at most one path per prefix, incremental deployments of countermeasures are difficult due to low path diversity.

Each prefix is usually announced by a single AS. However, this condition may be violated in practice either for legitimate reasons [104], such as origination of anycast prefixes, or illegitimate

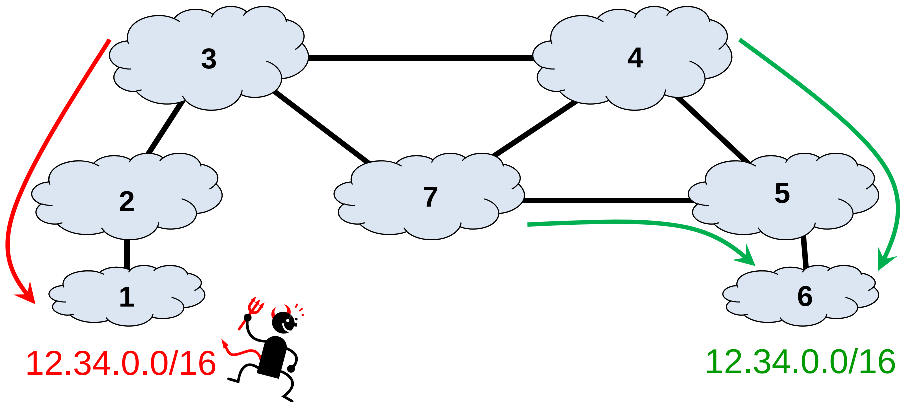


Figure 3.1: Announcement of prefix 12.34.0.0/16 from two origins partitions the network into two subsets.

ones, such as malicious attacks. In the latter case, the adversary gains control of the address block, and may either act as a sink and discard the received traffic, or act as a man-in-the-middle and forward packets to the legitimate destination. The term *prefix hijacking* usually refers to the origination of a victim prefix by an adversary instead of the legitimate origin AS. We will refer to this attack as a *simple origination attack*. An AS that selects a malicious route will propagate it to its neighbors, who may select it as well. For example, in Figure 3.1 prefix 12.34.0.0/16 is initially announced by AS 6, and all source ASes point their routing tables toward AS 6. If AS 1 also announces the same prefix, the network is partitioned in two subsets of ASes according to the origin AS they have chosen for the prefix: ASes 2 and 3 point their routing tables toward AS 1, whereas ASes 4, 5, and 7 point their routing tables toward AS 6.

Another type of attack is the *path-spoofing attack* in which the adversary announces a forged AS path to the victim prefix so that the adversarial AS appears upstream of the legitimate origin AS. Path spoofing is an intelligent attack, motivated by the adversary's desire to evade detection. However, the attack increases the length of the AS path, which may make some ASes less likely to select the malicious route.

In a third variant of prefix hijacking, the adversary breaks the victim's prefix into multiple sub-prefixes and originates those instead. In this way, although the network will maintain routes to both the original prefix and the sub-prefixes, because of the *longest prefix matching* rule used in the data plane, traffic will be directed to the adversary-controlled subprefixes. We refer to this attack as *sub-prefix hijacking*.

There are other variants of prefix hijacking, such as *wormhole* [49] attacks. Wormhole attacks are a countermeasure the adversary can employ against secure routing protocols. Wormhole attacks are not discussed in this chapter because in our evaluation scenarios the adversary is able to employ strictly more effective attacks.

3.4 Deploying soBGP in Small Groups is Ineffective

Secure routing protocols such as S-BGP and soBGP have been designed assuming ubiquitous deployment. In this section, we consider a partially deployable variant of soBGP, a protocol that has been widely discussed as an alternative to S-BGP [8]. Using simulations we demonstrate that small-scale deployments of soBGP provide only limited benefits to the adopters. These results motivate our exploration of the conditions that enable small groups to be effective.

3.4.1 Partial soBGP Deployment

The soBGP protocol is designed around a cryptographically-secured registry of routing information. The registry contains information about the prefixes each AS is authorized to advertise in BGP, as well as the pairs of ASes that are BGP neighbors. BGP advertisements are validated against the registry to ensure, first, that the origin AS in the advertisement has been authorized to advertise the corresponding prefix and, second, that all links in the AS-path of the advertisement match links included in the registry.

We consider a cryptographically secured registry of routing information similar to the one used by soBGP. However, our registry is *partial* in that it contains the routing information for only a subset of the ASes in the Internet. We call the set of ASes that publish information in the registry the *participants*, and all other ASes the *non-participants*. For each participant, the registry contains a list of prefixes the AS is allowed to originate as well as a list of the AS's neighbors. Besides having their information published, participants use the registry to validate BGP advertisements. A participating AS discards a BGP advertisement if it contains information that *contradicts* the registry. For example, the AS will discard a route to a registered prefix if the origin AS number is wrong. To evade detection, the adversary must spoof the origin of the route and possibly additional hops in the AS-path (such as the second-to-last-hop), to avoid contradicting the registry. Although the resulting advertisements do not contradict the registry, the longer AS-path makes them less likely to be selected by other ASes.

3.4.2 Evaluation Methodology

Our experiments evaluate the effectiveness of the aforementioned partial deployment of soBGP. The simulation techniques, and the data sets, that are introduced in this section are used throughout the chapter. Our experiments simulate the propagation of BGP route announcements on the AS-level Internet topology, as well as how the announcements affect the routing tables of each AS. Route propagation is profoundly influenced by AS business relationships, such as customer-provider, peer-peer or sibling-sibling. Since the goal of ISPs is to generate profit, and since customers have to pay their providers, ASes prefer routes learned from customers over routes learned from peers or providers; if multiple routes of the same class are available, the AS prefers shorter AS paths over longer ones. The business relationships also determine whether an AS exports the chosen route to its neighbors. An AS exports a route learned from its customer to all of its neighbors, whereas a provider or a peer route is exported only to customers.

Our simulations use and extend BSIM [53], which provides a convenient environment to simulate policy-based route propagation on an arbitrary AS topology. BSIM accurately captures the influence of business relationships on how ASes select and export routes. As input, we used an AS topology (annotated with the inferred business relationships) from June 2007 available from CAIDA [20]. This is considered to be one of the most accurate and most complete AS topologies available. The topology is constructed from snapshots of the routing tables from RouteViews servers [83], and it contains $25,304$ ASes. Routing table inspection at the end of each experiment allows us to determine what fraction of ASes selected valid routes to the victim prefix.

To measure the impact of an attack, we compute the average number of ASes that accept a route to the legitimate origin AS over a sequence of 100 experiments. In each experiment, the set of participant ASes and the adversarial AS are selected at random, and the victim is a randomly chosen member of the group. The variance of the quantities we measure can be high. This is because the outcome of each experiment critically depends on the location of adversary and the composition of the group. For example, if the adversary is the victim's sole provider, *all* ASes select the attacker's route, but the opposite configuration offers perfect security. Using the mean as the performance metric enables us to estimate how difficult it is for the adversary to mount a successful attack against the defending group. As the mean increases or decreases the number of places where the adversary can launch a successful attack increases or decreases correspondingly. Fortunately, as we demonstrate later in the chapter, our proposed solutions not only increase the mean but also decrease the variance.

There are two possible strategies the adversary can use to maximize the impact of an attack against a victim prefix belonging to a participant. In the first, the adversary launches a simple origination attack, listing its own AS number as the origin AS. This attack is intended to affect the *non-participants*, as the participants can easily detect that the offending route contains contradictory information. In the second, the adversary spoofs the shortest possible path that does not contradict the registry (i.e., the shortest AS path from any non-participant to the true origin AS). This attack is intended to evade detection by the *participants*. We assume that an AS has accepted a route to the adversary if it is affected by at least one of these attacks. This approach is justified by the fact that the adversary can combine the two attacks. For example, the adversary can use one of the attacks and if a particular AS of interest does not accept the adversarial route, the adversary can withdraw the first announcement and then use the other attack.

3.4.3 Simulation Results

In the first experiment, we consider a registry that has been formed by a random group of participants. Random participation is an apt deployment model when the locations of the participants are chosen based on criteria other than securing interdomain communication, for example, when the multiple sites of a single organization decide to act jointly. The adversary attacks a randomly chosen victim participating in the registry from a single randomly chosen AS. Figure 3.2 shows the percentage of participant and non-participant ASes that have accepted a route leading to the legitimate origin AS for the victim prefix.

The percentage of non-participant ASes that are able to reach the victim remains flat at approximately 50% as participation grows up to 30 members. In the absence of any protection mechanisms the victim would also be reachable on average by 50% of the ASes.[1] Therefore, if the participation is random, the partial registry is unable to help the victim, which remains as

[1]This observation is easy to verify using the symmetry between the victim and adversarial ASes.

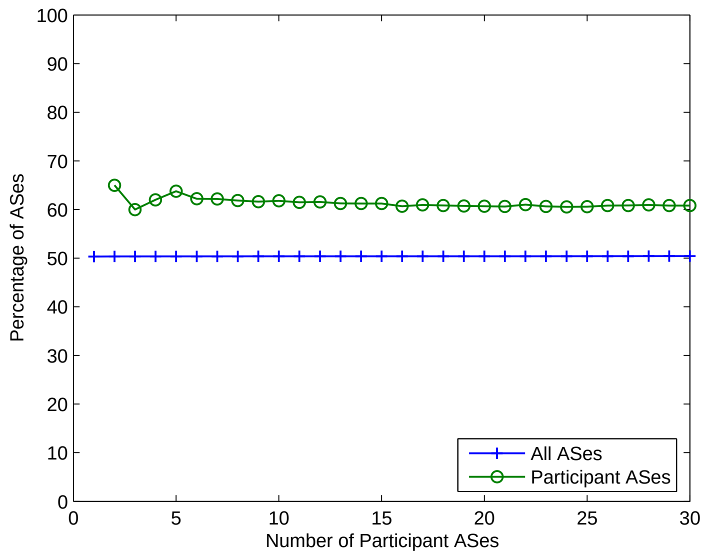


Figure 3.2: Up to 30 randomly selected member ASes participate in a routing registry. Figure depicts the percentage of ASes able to reach the prefix of the victim.

vulnerable as if there were no participants to prevent the propagation of the adversarial routes. This result can be explained by the following observation. Since most ASes in the Internet are stubs, the random selection of participants implies that stubs are most likely to be selected. However, stubs are *terminal* points that do not propagate any routes. Therefore, randomly selected participants can at most help themselves avoid the adversarial routes. Furthermore, although the percentage of participant ASes that are able to reach the victim prefix is higher, it does not exceed 60%. This result can be explained by the limited upstream connectivity of the stub participants that have only a few alternate routes to choose from. In fact, inspection of the AS topology used in the simulation reveals that approximately 35% of the stubs are connected to a single upstream provider, having a single route per destination prefix, and that approximately 40% of the stubs are connected to two upstream providers, therefore, having at most two routes per destination prefix.

It is also worth noting that, in addition to the poor average security gain that the partial registry attains, the security gain has high variability. For example, in a group of 10 randomly selected members, the legitimate origin of the victim prefix is reachable by *every* other member in only 22% of the simulation instances, whereas in 5% of the simulation instances *no* member can reach the legitimate origin.

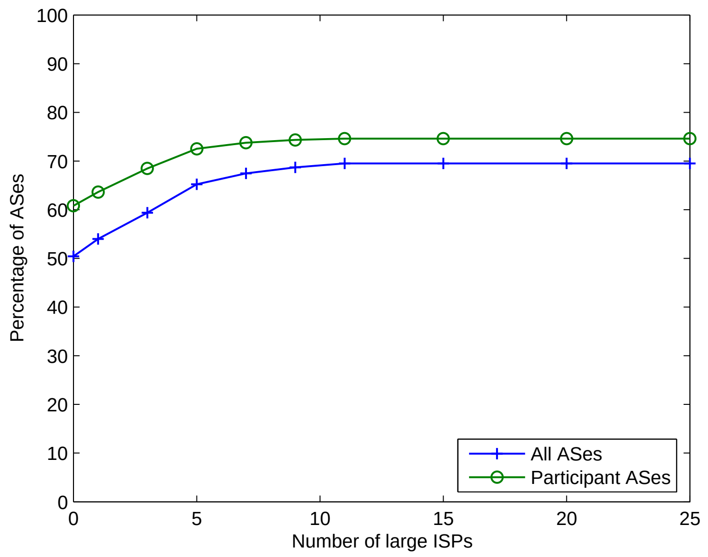


Figure 3.3: Up to 25 large ISPs plus 30 randomly selected ASes participate in a routing registry. The figure depicts the percentage of ASes able to reach the prefix of the victim.

In the next experiment, we consider a different participation model. The random set of participants is helped by enlisting in the group a set of *deputies*, which are large ASes having high node degree. Because of their rich connectivity, deputy ASes have the highest potential to prevent the propagation of adversarial routes by filtering these routes and propagating legitimate routes instead. Figure 3.3 shows the percentage of participant and non-participant ASes that have accepted a route to the victim prefix leading to the legitimate origin AS. The origin has been selected at random among a set of 30 participants. We decided to fix the number of participants to 30 as this corresponds to a large group of participants providing an upper bound on the effectiveness of any smaller group. The adversary attacks from a single randomly chosen AS. We consider a group of deputy ASes ranging in size from 0 to 25. Although the benefits are significant in comparison to the results obtained by random deployment, performance is poor even if the number of deputy ASes is large. For example, more than 25% of the ASes still cannot reach the legitimate origin, even if the group includes the the 25 highest-degree nodes. The corresponding percentage for non-participants is even larger. It is also worth noting that the formation of groups of 25 or more large ISPs is not realistic because it requires significant coordination by networks that are otherwise business competitors and have been notoriously reluctant to form sizable coalitions.

3.5 Perfect Filtering of Invalid Routes is Not Sufficient

In Section 3.4, the group members do not enjoy significant security benefits, in part because they cannot accurately detect invalid routes. In this section, we explore whether more accurate detection techniques are enough to make small groups effective. We first explain how participating ASes can detect and filter invalid routes using existing proposals for detecting control-plane and data-plane anomalies. Then we present simulation experiments that show that, while ideal filters offer clear security benefits, the adversary is still able to inflict substantial damage.

3.5.1 Accurate Detection of Invalid Routes

Although protocols like S-BGP and soBGP rely on cooperation in creating and maintaining registries, several new solutions have been proposed so that individual ASes can detect (and potentially filter) invalid routes. Some solutions detect invalid routes in the control plane (i.e., by analyzing BGP announcements), whereas others operate in the data plane (i.e., by monitoring the forwarding path).

Control-plane techniques: Anomaly-detection techniques running in the control plane can identify suspicious routes based on a history of past BGP announcements [54,61,79]. These techniques essentially allow individual ASes to use historical data to construct a more complete *de facto* registry of prefix ownership and AS-level connectivity. Although early anomaly-detection techniques only detected simple origination attacks [54, 61], recent solutions are able to detect more sophisticated path-spoofing attacks launched by intelligent adversaries [79]. Although these techniques provably detect spoofed paths, they are vulnerable to false alarms that mistakenly flag valid routes as suspicious.

Data-plane techniques: An alternate approach to detecting suspicious routes is to detect anomalies in the forwarding path to the (alleged) destination. For example, a significant change in the number of hops in the path, or differences in the end-host properties, would suggest that a hijack has occurred [48, 105]. However, these techniques are vulnerable to intelligent adversaries who actively try to evade detection. More sophisticated data-plane techniques are possible when the source and destination ASes cooperate to detect availability problems along the path. For example, the communicating ASes could employ passive-monitoring techniques, like coordinated sampling [28] and stealth probing [10], to monitor loss and delay of the data packets traversing the forwarding path.

When data-plane measurements or control-plane anomalies suggest that the current path is invalid, a participating AS can simply filter the offending route, and select an alternate path (if one is available). The proposed detection techniques vary in how well they identify invalid routes, withstand intelligent adversaries, and avoid false positives. In the rest of this section, we evaluate the effectiveness of small groups that are armed with a *perfect* detector of invalid routes. As such, our results provide an *upper bound* on how well the proposed detection techniques might work in practice, when deployed in a small group of ASes. Since our results show that even a perfect detector is not sufficient to make small groups effective, we do not evaluate the individual detection techniques.

In conducting the evaluation, we must consider how an intelligent adversary would adapt the attack strategy in the face of these detection techniques. In particular, the adversary no longer has any incentive to spoof the AS path, since the participating ASes can easily detect and discard the invalid route. Instead, the adversary's best strategy is to launch a simple origination attack, in the hope of enticing as many non-participants as possible to select an invalid route. The impact of the attack is determined by the number of ASes that accept a route leading to the adversary, or are left with no route at all.

3.5.2 Simulation Results

In this simulation, we evaluate the effectiveness of ideal filtering when the victim AS is selected at random among a set of 30 randomly-chosen participants. As in the previous section, we decided to fix the number of participating ASes at 30, as this corresponds to a relatively large group of participants providing an upper bound on the effectiveness of any smaller group. The adversary attacks from a single randomly chosen AS, and we consider the benefits of enlisting 0 to 25 large ISPs as deputies.

Figure 3.4 shows the percentage of participant and non-participant ASes that select a legitimate route to the victim prefix. When 25 deputy ASes augment the group of 30 participants, around 15% of the group members cannot reach the victim. This represents a significant improvement over the partial soBGP deployment evaluated in the previous section, where more than 25% of the

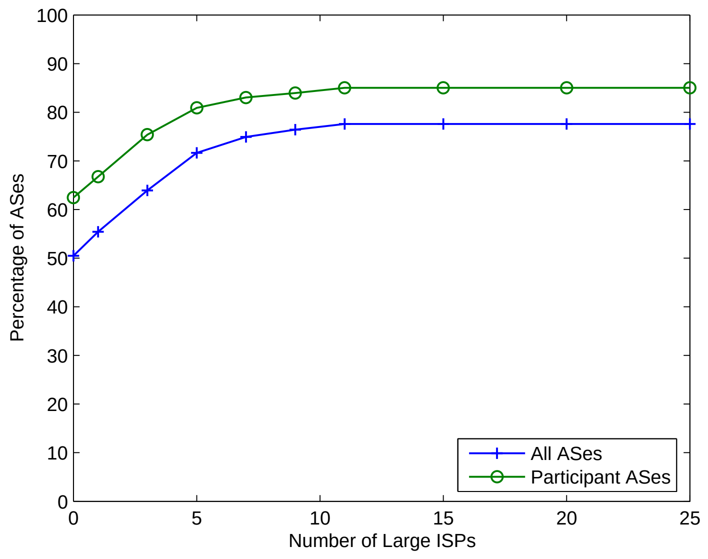


Figure 3.4: Ideal filters are employed at 30 member ASes and up to 25 additional large ISPs. Figure depicts the percentage of ASes able to reach the prefix of the victim.

participants could not reach the victim. Non-participants also benefit when the group members and deputy ASes can accurately filter invalid routes. Compared to the previous section, the percentage of non-participants able to reach the victim prefix increased from 65% to 75% assuming 10 deputy ASes participate. The non-participants benefit because the deputy ASes select valid routes, essentially blocking the propagation of invalid routes. This decreases the likelihood that a non-participant inadvertently selects an invalid route.

Despite the security gains attainable over partial soBGP deployment, we believe that too many ASes must participate before significant benefits are achieved. Accurate detection of invalid routes and the support of large ISPs are helpful, but not sufficient. Since the non-participants cannot detect invalid routes, they propagate these routes to the participants in lieu of legitimate routes. As such, some participants do not learn *any* valid route. Many of these ASes are stub networks with just one or two upstream providers, which significantly limits the number of BGP routes they learn. Although these ASes can detect that the routes they learn are invalid, they do not have enough options to select a valid alternative. In the next section, we present a technique that overcomes this limitation by providing the participating ASes with additional routes. This new technique offers significantly better security gains, even for groups as small as 5–10 ASes.

3.6 Secure Overlay Routing

This section describes how a small group of ASes can effectively secure interdomain communication between its members. The group members form a secure overlay network that offers alternate overlay paths a participating AS can use when no valid BGP route is available. First, we introduce the Security Backbone (SBone), which protects intra-group traffic despite *deployment gaps* separating the participants. Then, we present simulations that quantify the substantial security gains for a realistic Internet topology. Although primarily designed to protect *availability*, the SBone can also enhance the confidentiality and integrity, as discussed at the end of the section.

3.6.1 Security Backbone (SBone)

The participating ASes must effectively handle deployment gaps, i.e., non-member networks that are uncooperative or even hostile. To enable participants to circumvent availability problems, we connect the group members by a mesh of virtual links forming an overlay network. We call this overlay network a Security Backbone (SBone). The SBone differs from traditional overlays (e.g., RON [9]) in that it is an overlay of *networks* rather than individual end hosts or servers. In traditional overlays, the participating hosts have little or no control over the routes their upstream providers pick. In contrast, because the SBone is created by the administrators of the participating ASes, it has visibility into (and control over) BGP routing. In fact, the SBone may run directly on the routers in the participating ASes.

The virtual links are created by connecting members networks with IP tunnels that encapsulate and decapsulate the data packets. For each pair of group members X and Y, we create two tunnels: one from X to Y, and one from Y to X. The SBone improves availability in two main ways. First, multiple interdomain paths may exist between the endpoints of each virtual link, and traffic can be directed over any of these paths. That is, an SBone node could switch a virtual link from one underlying interdomain path to another, after detecting an availability problem on the original path. Second, if all of the underlying paths have availability problems, an SBone node can direct traffic through an intermediate SBone node via the overlay network.

The SBone relies on a monitoring system to detect availability problems and to disseminate the measurement results to other nodes. The SBone could apply any of the monitoring techniques outlined earlier in Section 3.5.1, but techniques tailored to detecting availability attacks are especially appealing. For example, the SBone nodes could easily apply highly accurate availability-monitoring techniques that require cooperation between the communicating end points [10, 28], or active probes sent by separate probe machines. These data-plane monitoring techniques allow the SBone to react to a broader range of attacks, including adversaries that propagate valid BGP advertisements while maliciously dropping or delaying the data packets—a problem traditional secure routing protocols like S-BGP cannot handle.[2]

3.6.2 Random Deployment

We evaluate how effective the SBone is in preventing routing attacks using simulation. As before, our simulator is based on BSIM, and the AS topology is the same as in the previous experiments. In the simulation, we assume that the participating nodes are equipped with ideal filters that are able to distinguish the routes leading to the adversary. Figure 3.5 shows the percentage of group members that are able to reach the legitimate origin AS through an overlay path. The origin AS, the defending group, and the adversarial AS have been selected at random among the set of all ASes. The size of the defending group reaches up to 30 members. We consider a case in which the group members are not able to select the BGP path of a virtual link, a scenario akin to a setting in which the group members are overlay hosts that do not have visibility into BGP, and a case in which the group members have control over BGP. The latter case is shown to outperform the former. Note, however, that participation of large groups is required to attain significant security gains in either case. A refinement is required so that deployment of the SBone in small groups attains significant security gains.

3.6.3 Reinforcement through Tier-1 ASes

In the previous experiment, the participant ASes were selected at random from the set of all ASes. Since the majority of ASes are stubs of limited upstream connectivity, this resulted in limited path diversity, and hence large groups were required to attain significant security benefits. Therefore, we consider a participation model similar to the one used in Sections 3.4 and 3.5 in which the random

[2]In fact, *secure* availability monitoring techniques could protect against especially insidious adversaries that try to bias the data-plane measurements to evade detection [10].

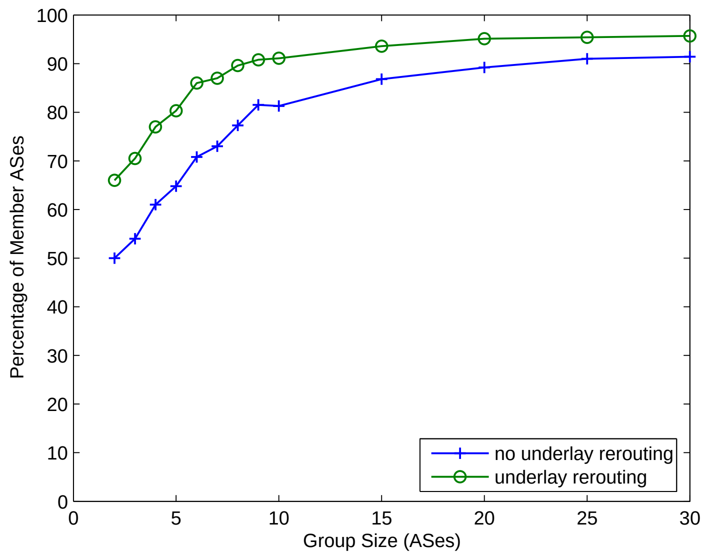
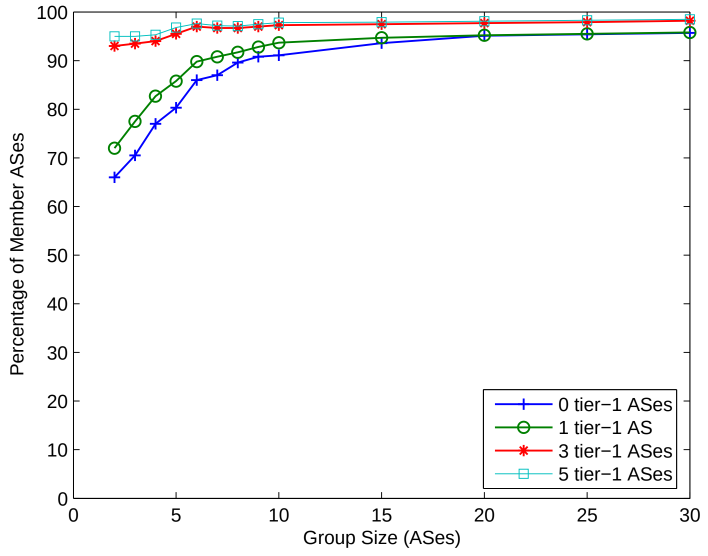


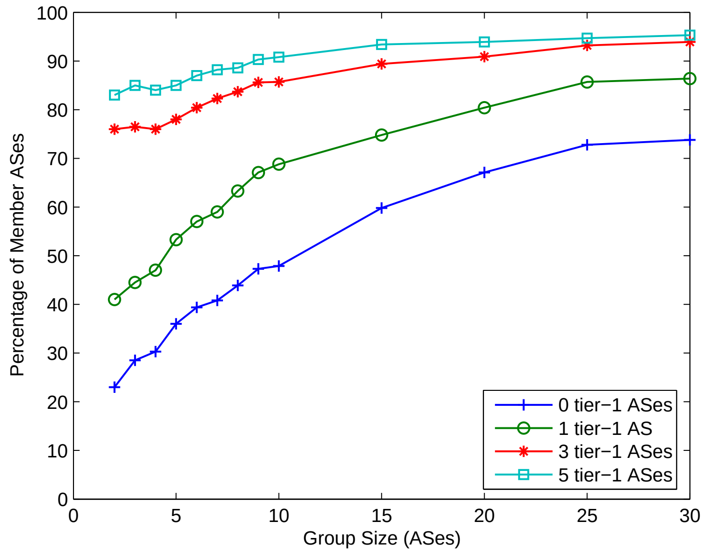
Figure 3.5: Percentage of up to 30 randomly selected ASes that are able to reach the legitimate origin through an overlay. Case with underlay rerouting allows a participant to control the selection of its BGP routes.

group of participants is reinforced by a few tier-1 ASes acting as deputies to assist the group. Because of their rich connectivity, deputy ASes are able to expose rich path diversity that is able to overcome the limitations the random group of participants faces. In the next experiment, we demonstrate that enlisting one or more such deputy ASes in a randomly selected group can thwart the adversary's power even if the overall size of the defending group that includes the deputies is small. Note that the participation of large ASes in the defending group could be arranged in exchange for pay.

Figure 3.6(a) shows the percentage of group members that are able to reach the legitimate origin of a randomly selected victim prefix. The deputy ASes are excluded from the set of possible origins and the adversary has also been selected at random. We assume that the adversary attacks both the participant's prefix and the tunnel endpoints of the overlay formed by the defending group. Both the participants and deputy ASes are equipped with filters that are able to distinguish the routes leading to the adversary. We consider four cases such that the group of deputy ASes consists of 0, 1, 3, and 5 members. We find that the percentage of participants able to reach the victim prefix exceeds 95% provided that the group of deputy ASes consists of 3 or more members. This



(a) 1 Adversarial AS



(b) 5 Adversarial ASes

Figure 3.6: Up to 5 tier-1 ASes assist the group of participants. Figure depicts the percentage of group members that have an overlay path to the prefix of the victim.

result holds irrespective of the size of the group of participants. We also find that if the size of the group of participants is small, the security gains improve substantially as the size of the group of deputy ASes increases from 0 to 5 members; larger groups of participants are able to achieve significant security gains without the help of the deputies.

Five Adversaries

So far, we have assumed that the adversary attacks from a single AS, and we have shown that the security benefits attainable by partially deployed secure routing protocols are poor even against the smallest adversarial group. In this section, we show that the SBone is resilient not only against a single adversary but also against larger adversarial groups. In particular, we consider an adversarial group that consists of 5 members.

Figure 3.6(b) shows the percentage of group members able to reach a victim prefix belonging to a randomly selected participant under the same assumptions as in the experiment shown in Figure 3.6(a). We observe that large adversarial groups have a significant impact on those defending groups that do not enlist deputy ASes. Furthermore, the effect of large adversarial groups diminishes as the number of deputy ASes that are enlisted increases. For example, in a defending group of 10 members that enlist 5 deputy ASes, 90% of the members are able to reach the legitimate origin of the victim prefix.

Comparison with Secure Routing Protocols

We use the security benefit attainable by an ideal secure routing protocol as a baseline for comparison of the benefit attainable by the SBone under large adversarial groups. Figure 3.7 shows the percentage of member ASes able to reach a victim prefix belonging to a randomly selected member AS that is attacked by 5 adversarial ASes. We consider two cases. In the first case, an ideal secure routing protocol has been used to protect the victim and, in the second case, the victim is protected by the SBone. We assume that the group of participants consists of 10 members. We consider four cases such that the group of deputy ASes consists of 0, 1, 3, and 5 tier-1 ASes. We find that the SBone is able to protect $20 - 30\%$ more participants than the ideal secure routing protocol.

It is also worth noting that the SBone has significantly lower variability in its performance than the ideal secure routing protocol. For example, in the case of 5 deputy ASes, the standard deviation corresponding to the SBone is 8%, whereas the standard deviation of the ideal secure

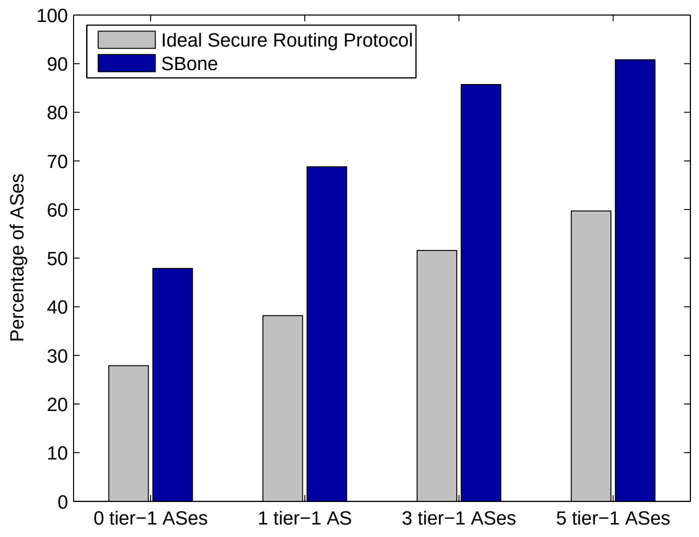


Figure 3.7: Comparison of the security performance of an ideal secure routing protocol and the SBone against 5 adversarial ASes. Up to 5 tier-1 ASes assist the group of 10 participants. Figure depicts the percentage of the participants able to reach the victim prefix.

routing protocol is 15%. Therefore, the security benefit of the SBone is more predictable than the benefit an ideal secure routing protocol would attain.

3.6.4 Enhancing Confidentiality and Integrity

The SBone also enables routing strategies that protect confidentiality and integrity by preventing adversaries from receiving sensitive traffic. For example, the SBone allows the group to select overlay paths that proactively avoid untrusted ASes owned by business competitors or known not to implement best common practices. Furthermore, in a manner complementary to filtering unwanted routes, group members can employ routing strategies that mitigate the risks of attacks. For example, the group members can proactively spread traffic over multiple paths, reducing the overall amount of traffic carried over any single path, and, therefore, substantially increasing the amount of resources an adversary would have to invest to observe all the traffic.

3.7 Shout

In this section, we present Shout, a technique that the defending group can employ to secure traffic originating from non-participating ASes that is destined to the participating ASes. Shout extends the benefits of the SBone to non-participating ASes that are entirely agnostic of the protection mechanisms that the participants apply. The ability to secure traffic originating from a potentially large number of non-participating ASes is in the vested interest of the participants as in this way the value they receive from their participation in the system increases. Noting that to receive the full benefits of the SBone an AS must eventually become a participant, Shout is intended for deployment during the transient period in which the SBone is incrementally building up to include a target group of members.

3.7.1 Hijacking the Hijacker

Shout is a destination-driven technique that attracts traffic from sources that may not be participating in the system and that may even be ignorant of the existence of the system. The ability to offer the service despite non-participating traffic sources decreases the degree of participation required for the system to be effective and facilitates the formation of small groups. Shout coaxes non-participants into picking routes leading to nearby participants instead of routes leading to adversarial ASes. Shout competes with the adversary to attract traffic from the non-participants using the adversary's own armory. Shout *hijacks the hijacker* by having the defending group of ASes simultaneously originate in BGP a participant's prefix. In this way, even if the adversary attacks the prefix receiving the protection of the group, the non-participants will prefer the routes leading to the group members over the adversary's routes.

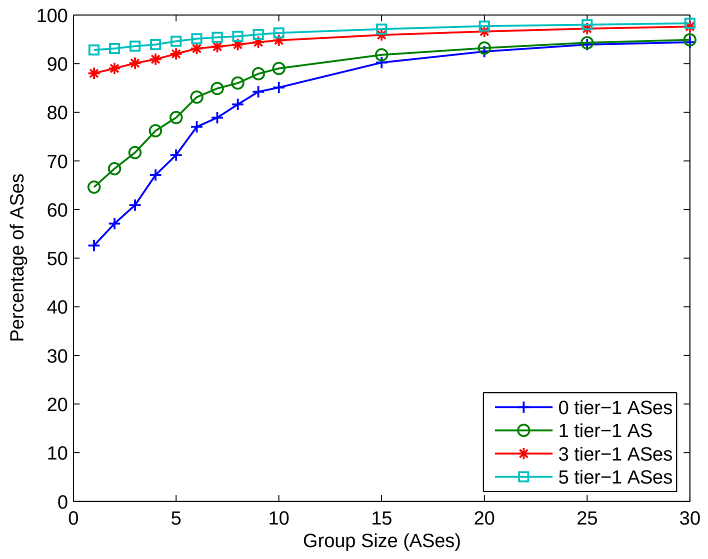
The simultaneous origination of the prefix from multiple ASes may cause the network to direct traffic from different sources attempting to communicate with the same IP address to different destination machines. Although there are cases where a behavior like this is a limitation, sometimes this behavior is desirable. An example is *anycast*, a service in which multiple hosts share the same IP address. Traffic destined to this address may be delivered to any of the corresponding hosts. Anycast is well suited for applications that do not maintain state across multiple packets, such as single request-response applications like DNS. In fact, anycast has been deployed to improve the resilience of root DNS servers [47]. Anycast becomes problematic in connection-oriented multi-packet transfers requiring state across different packets, like in TCP.[3]

For those applications that anycast is unsuited for, the traffic that enters the group must be delivered from the AS that first receives the traffic to the participant AS hosting the prefix. This delivery is carried out by the SBone. The effectiveness of the combined arrangement of, first, using Shout to coax non-participants into picking routes leading to participants and, then, using the SBone to deliver the traffic to its destination was evaluated using simulation. As before, our simulator is based on BSIM, and we use the same AS topology as in the previous experiments. Figure 3.8(a) shows the percentage of all ASes in the Internet that are able to reach a participant's prefix when the corresponding participant AS is selected at random and the adversarial AS is also selected at random. We assume that the adversary attacks both the participant's prefix and the tunnel endpoints of the overlay formed by the defending group. We consider four cases in which 0, 1, 3, and 5 tier-1 ASes are enlisted in the defending group. If the group has 10 or more members and 3 or more tier-1 deputies, more than 95% of the ASes in the Internet are able to reach the victim despite the attack.

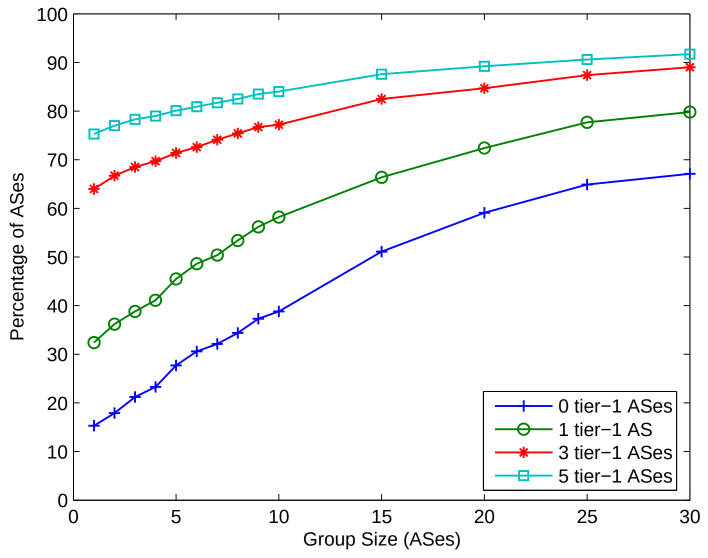
The case of an adversarial group that consists of 5 members is shown in Figure 3.8(b). As in the case of the SBone, large adversarial groups have a significant impact in those defending groups that do not enlist deputy ASes. Furthermore, the effect of large adversarial groups diminishes as the number of deputy ASes that are enlisted increases. For example, in a defending group of 10 members that enlists 5 deputy ASes, 85% of the group members are able to reach the legitimate origin.

Finally, the security benefits of the SBone are compared with the security benefits of ideal secure routing protocols in Figure 3.9. If the group of the participants consists of 10 members and the group of the adversary consists of 5 members, we find that the SBone is able to protect $20 - 40\%$ more participants than the ideal secure routing protocol.

[3]This is because routing changes in the middle of an ongoing connection can lead the packets to a different host, which may not have the appropriate state to continue the transfer.



(a) 1 Adversarial AS



(b) 5 Adversarial ASes

Figure 3.8: Up to 5 tier-1 ASes and up to 30 participants use Shout and SBone. Figure depicts the percentage of ASes in the entire Internet that have a path to the prefix of the victim.

3.7.2 Performance and Scalability

We use simulation to measure the impact of Shout on the performance and scalability of interdomain routing. The impact on performance is measured by the extent to which end-to-end paths

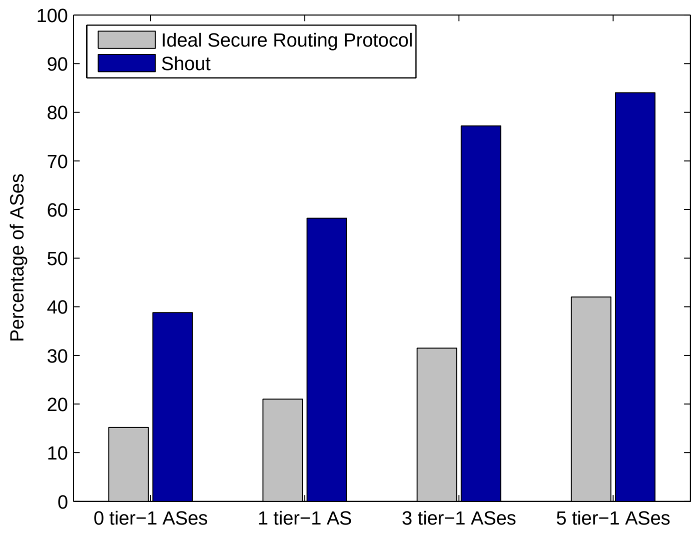


Figure 3.9: Comparison of ideal secure routing protocol and the SBone reinforced by Shout. Up to 5 tier-1 ASes assist the group of 10 participants. Figure depicts the percentage of all ASes able to reach the victim.

are prolonged, whereas the impact on scalability is measured by the increase in routing table sizes. We show that the increase in both quantities is small.

Shout forces the data traffic to take detours prolonging the end-to-end path, even when no adversary is launching an attack. Within each randomly-created group, we select a random destination AS, and enlist either 0, 1, 3, and 5 tier-1 ASes as deputies. We compute the length of the path from each (source) AS to the destination without Shout enabled. Then, we simulate the operation of Shout and determine the path the traffic would take to the destination AS, where part of the path takes the traffic from the source to the group and the remainder traverses the virtual link to the destination. Figure 3.10 shows the ratio of the path length after activating Shout over the path length before activating Shout, averaged over all source ASes. As expected, the ratio decreases as the number of deputy ASes increases. In a group where no tier-1 ASes participate, the ratio is at most 1.35. In a group in which 5 tier-1 ASes participate, the ratio drops below 1.15.

Considering now the impact of Shout on the routing-table size, we note that by originating a prefix from multiple ASes, Shout increases the number of alternate BGP routes for that prefix. Although in BGP each router selects exactly one route for each prefix, the alternate routes are

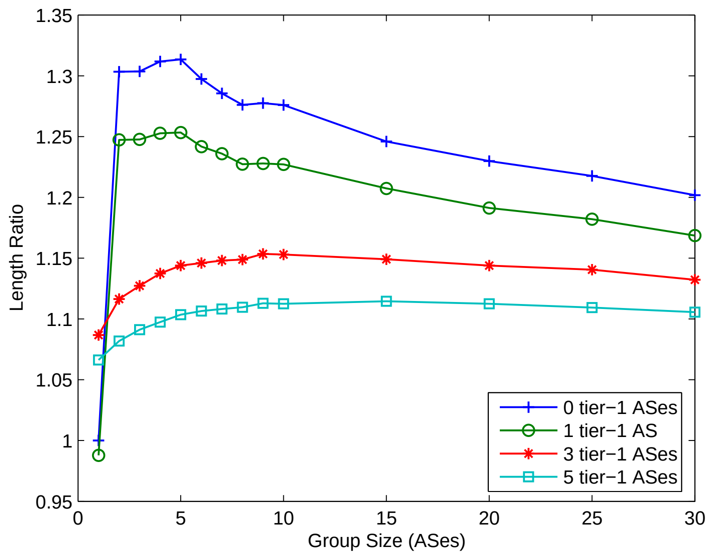


Figure 3.10: Impact of Shout on the lengths of paths used to reach the overlay origin. Ratios of the original and resulting hop counts depicted.

stored in the RIB, increasing the size of the RIB at those routers. To evaluate the impact of Shout on RIB size, we counted the number of routes each AS stores for the "shouted" prefix. The average RIB size never increased by more than 5%, across all simulation instances for up to 30 group members and 0, 1, 3 or 5 deputy ASes.

Our experiments show that the increase in path length and routing-table size are modest. Still, in some cases, the group members may not want to incur the small performance penalty or carry the extra traffic. Fortunately, the Shout mechanism can be employed *reactively* upon detecting a prefix-hijacking attack, using a control-plane anomaly detector like the ones described earlier in Section 3.5.1. When one of the group members, or a separate anomaly-detection system, detects a hijack, the group members can be instructed to activate Shout for the affected prefix. As long as the detector has a low false-positive rate, and few if any false negatives, invoking Shout reactively is both effective and efficient. Fortunately, several such control-plane anomaly detectors already exist [79]. Ultimately, the decision of whether or not to run Shout reactively depends on what trade-off the group wants to strike between efficiency/performance and delay in reacting to attacks.

3.7.3 Discussion

The goal of Shout is to increase the resources an adversary must expend to disrupt the communication of a participant with the non-participating ASes. Shout protects the direction of communication from the non-participating traffic sources to the participating destinations. Our system does not protect the reverse direction of communication. Despite this limitation, the resources the adversary must expend to attack a participant are substantial. The reason is since traffic destined to the participant is secured, the adversary must be in a position to attack the traffic originating from the participant. However, to attack this traffic the adversary must be able to attack a potentially large number of traffic destinations. Therefore, Shout substantially raises the bar for the adversary to perform a successful attack against a participant. It is worth noting that our system does not attempt to level the degree of protection offered to the communication between participants with the degree of protection offered to the communication between a participant and a non-participating AS. We believe that this compromise is justifiable by the fact that once a non-participant decides to participate, he can receive the full benefits of the system.

3.8 Defending Against Sub-Prefix Hijacking

Thus far our adversary attacked a prefix by announcing it in BGP. Here we examine a scenario in which the adversary breaks or *deaggregates* the victim's prefix into multiple sub-prefixes and originates those instead. Although the network will maintain routes to both the original prefix and the sub-prefixes, the *longest prefix matching* rule will ensure that traffic is directed to the adversary's subprefixes. We refer to this attack as *sub-prefix hijacking.* There is a limit to the granularity that the adversary can deaggregate a prefix. This limit is imposed by filtering rules employed by ISP networks that discard routes to prefixes more specific than a /24.

3.8.1 Defending the Participants

The tunnel endpoints of the SBone overlay can be easily protected against sub-prefix hijacking. To protect a tunnel endpoint, the corresponding participant can announce the address of the endpoint with a /24 prefix, preventing the adversary from announcing a more specific prefix covering the endpoint. If a network participates in more than one overlay simultaneously, it is possible to use the same tunnel endpoint address in each of the overlays. Therefore, each network needs to only announce one /24 prefix irrespective of the number of the overlays it participates in.

Since the tunnel endpoints of the SBone overlay are resilient to sub-prefix hijacking, the delivery of traffic among participants is also resilient to this attack. This is true because traffic sent by participants is delivered via the secure overlay, with the packets encapsulated with the IP address of the recipient's tunnel end-point. Therefore, at full participation, to attack the intra-group traffic, the adversary must resort to the prefix hijacking attacks we considered in the previous sections. Next, we describe how one can protect the traffic sent by non-participants against sub-prefix hijacking attacks.

3.8.2 Defending the Non-Participants

One way to protect non-participant traffic destined to the participants is to proactively deaggregate the prefixes of the participants into sub-prefixes and use Shout to advertise the sub-prefixes instead of the aggregate prefixes. However, such a proactive countermeasure would create extra routes leading to a potentially significant increase of the BGP routing-table size. In the following, we present a reactive technique that is able to relieve the routing system from the extra routes. This technique is useful because it can obviate the need for the current practice in which ASes deaggregate their prefixes proactively [90]. This is well aligned with the recent efforts to limit the growth of the BGP routing tables.

Monitoring BGP Advertisements

To protect traffic sent by non-participants and destined to the participants against sub-prefix hijacking, we use the participants to monitor BGP advertisements and detect the offending advertisements. Assuming each participant knows the prefixes every other participant should advertise, if a BGP advertisement announcing a subprefix of the known prefixes is detected by any participant, a sub-prefix hijacking must have occurred. If any participant detects this event, it notifies the rest of the members, who respond to the attack by advertising the same sub-prefix. Based on the results of the previous section regarding the impact of Shout on the routing-table size, this countermeasure does not increase the routing table entries by more than 5% of the number of sub-prefixes the adversary has already advertised.

Evading the Monitors is Ineffective

To prevent the countermeasures from taking effect, the adversary will attempt to conceal the attack from the participants. However, the adversary has limited control over how the sub-prefix advertisement propagates in the system. Control over the propagation of the advertisement can be exercised using the following trick: BGP has a mechanism for avoiding loops in which a BGP advertisement is discarded by an AS if that AS already appears in the AS-path. Relying on this mechanism, the adversary can prevent the participants from receiving the offending announcement by adding the *union* of the neighboring ASes of each participant to the AS-path of the announcement. In this way, loop detection will be triggered at the neighboring ASes preventing the announcement from reaching the participants.

In order to conceal the attack from the participants, the adversary must limit the impact of the attack on the non-participants. To measure the damage that the adversary can induce, we used simulation on the AS topology one more time. Figure 3.11 shows the percentage of ASes that are able to reach a randomly selected victim prefix against which the adversary mounts a sub-prefix hijacking attack. We assume that the adversary forges the offending BGP advertisement so that

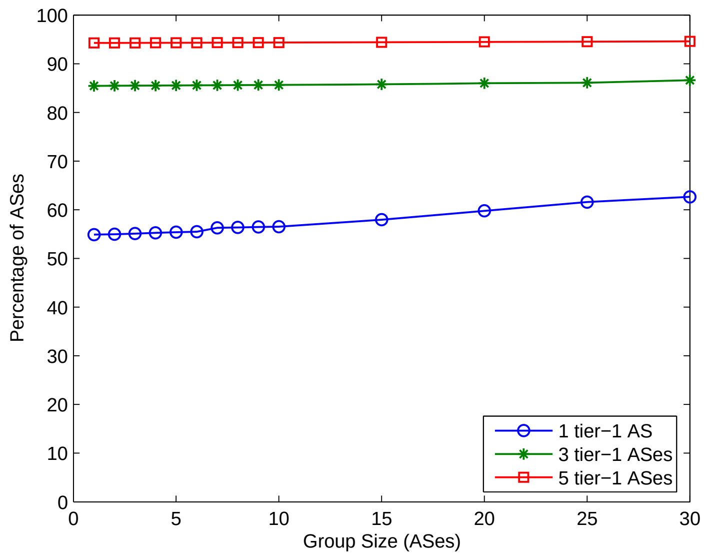


Figure 3.11: Percentage of ASes able to reach the legitimate origin of a victim prefix against which the adversary mounts a sub-prefix hijacking attack.

the attack is concealed from the set of ASes that are neighbors of the participants. We consider three cases in which 1, 3, and 5 tier-1 ASes comprise the set. It is shown that to conceal the attack from 5 tier-1 ASes, the adversary affects less than 5% of the ASes in the Internet.

Figure 3.11 shows that the impact of the adversary's attack decreases rapidly as the number of participating tier-1 ASes increases. This can be explained by the fact that tier-1 ASes have a large number of neighbors, rapidly increasing the number of ASes that must discard the offending advertisement to prevent detection. In fact, inspection of the topological map of the Internet reveals that the number of ASes that discard the malicious advertisement through loop detection is on order of several thousand. This observation gives rise to a simple countermeasure to thwart the adversary's ability to conceal a sub-prefix hijacking attack using the aforementioned trick. In this countermeasure, filters prevent the propagation of BGP advertisements containing more than a few hundred ASes in the AS-path. In practice, AS-paths of this length only appear due to configuration errors (or attacks), and anecdotal evidence suggests that certain ISPs already filter these long BGP advertisements. Moreover, such advertisements have been known to cause some routers to crash.

3.9 Deployment Considerations

In this section we explore practical issues that concern deployment of our system. So far we have been mostly considering the average level of security attainable when the members of a group are selected at random. However, the security attainable by any given group depends on the relative location of the group members. Although the desired security level may not be achievable using the resources of the group itself, enlisting one or more large ISPs, similarly as we did in our experiments, may be sufficient to achieve the target objective. We argue that it is important to consider both the location of the members and the goals of the group when enlisting deputy ASes.

SBone: Let's consider intra-group communication. We observe that the security of a path between two participants improves as their relative distance decreases. This is intuitive since as the number of AS hops in the path increases, the probability that an intermediate AS selects a compromised route increases as well. Therefore, if the group members are located close to one another, it may be possible to effectively secure intra-group communication without the help of additional deputy ASes. In contrast, groups that contain remotely located members can ensure secure intra-group communication by enlisting deputy ASes, effectively decreasing the relative distance of the remote group members.

Shout: Let's consider communication of non-participants and participants. We observe that, in contrast to intra-group communication, diversity in the location of the members improves the effectiveness of Shout. Therefore, enlisting deputy ASes can be helpful for groups of limited geographic presence. The choice of which deputy ASes to enlist critically depends on the footprint of their topology. For example, a group may prefer deputies with a strong presence in areas that contain vulnerable non-participants that are of particular interest to the group members.

We envision that the group can enlist deputy ASes in exchange for a fee. Since our solution allows creation of multiple co-existing groups, each group can enlist the nodes that are the most beneficial for the particular configuration and desired level of protection. To that end, the groups can employ the simulation techniques we introduced in this chapter to find the best deputies to enlist.

3.10 Related Work

Our research relates to earlier attempts to secure the interdomain routing system. Previous solutions have focused primarily on *cryptographic techniques* that ensure the validity of the route announcements [49, 58, 75, 101], or on *anomaly detectors* that detect (and in some cases filter) suspicious routes [48, 54, 61, 79, 105]. The cryptographic techniques are expensive and offer limited (if any) gains in small-scale deployments. Filtering invalid routes based on anomaly-detection techniques is effective, but only for larger group sizes (e.g., 50 or more ASes, including large ISPs). Our solution leverages these anomaly detectors as one way to detect, and avoid, compromised routes. Yet we show that, to be effective, small groups also need to increase path diversity and proactively coax non-participants to pick valid routes.

Our research is part of a recent body of work exploring interdomain security solutions that are successful in smaller-scale deployments. For example, the work in [100] argues that large ISPs can offer increased path diversity as a service, to offer their customers higher availability. In our work, we provide additional path diversity through secure overlay routing, and also propose Shout to help the *non-participants* reach the group members. As another example, the work in [21] models the conditions (in terms of costs and security benefits) that would lead all ISPs to adopt one of the various secure routing protocols. In contrast, we quantify the effectiveness of the protocols in small-scale deployments, and identify new mechanisms that enable small groups to be effective. Finally, this chapter builds on earlier work [11] that proposed the secure overlay routing technique in Section 3.6.1. However, the earlier paper did not present any simulation experiments or propose techniques for coaxing the non-participants to select valid routes.

3.11 Summary

In this chapter, we argued that small groups of cooperating ASes should form the basis of solutions for secure interdomain routing, with an emphasis on improving end-to-end availability. We evaluated and compared several techniques that small groups can employ, to identify four conditions that enable them to be effective. In particular, the participating ASes should: (i) apply accurate techniques for detecting and filtering compromised routes, (ii) cooperate to expose additional path diversity, (iii) actively induce the non-participating ASes to select valid routes, and (iv) enlist a few large ISPs to join the group. All four conditions are important—omitting any of them results in a much less effective solution. We also proposed and evaluated a novel approach, based on secure overlays and cooperative BGP announcements, that achieves these four goals, allowing groups as small as 5–10 ASes to enjoy substantial security benefits.

As small groups become effective in securing interdomain routing, other ASes may want to join the group. As more ASes join, both parts of our solution become more effective—the overlay exposes even greater path diversity and the cooperative BGP announcements reach an even larger fraction of the remaining non-participants. Interestingly, as the group grows even larger, the "shout" mechanism becomes increasingly less necessary because most important communication stays within the group. However, as the group grows in size, the assumption that all group members trust one another becomes less reasonable. At that point, it becomes important to protect the honest members of the group from malicious *participants*. For example, the members of the group could agree to deploy a secure routing protocol *within the overlay network*, in addition to the mechanisms we have already discussed for protecting against adversaries outside of the group.

In fact, the deployment of a secure routing protocol within the group, coupled with the group's growing size, could present a viable incremental deployment path for traditional cryptographic solutions, like S-BGP.

Chapter 4

Load Balancing in the Presence of Failures

4.1 Introduction

To ensure uninterrupted data delivery, communication networks must distribute traffic efficiently even as links and routers fail and recover. By tuning routing to the offered traffic, *traffic engineering* [76] improves performance and allows network operators to defer expensive outlays of new capacity. Effective *failure recovery* [77, 89]—adapting to failures by directing traffic over good alternate paths—is also important to avoid performance disruptions. However, today's networks typically handle failure recovery and traffic engineering independently, leading to more complex routers and less efficient paths after failures. In this chapter, we propose an integrated solution with much simpler routers that balances load effectively under a range of failure scenarios.

We argue that traffic engineering and failure recovery can be achieved by the same underlying approach—dynamically rebalancing traffic across diverse end-to-end paths in response to failures. This reduces the complexity of the routers by moving most functionality to the management system—an algorithm run by the network operator. Our architecture has three key features:

Precomputed multipath routing: Traffic between each pair of edge routers is split over multiple paths that are configured in advance. The routers do *not* compute (or recompute) paths, reducing router overhead and improving path stability. Instead, the management system computes paths that offer sufficient diversity across a range of failure scenarios, including correlated failures of multiple links.

Path-level failure detection: The ingress routers perform failure recovery based only on which *paths* have failed. A minimalist control plane performs path-level failure detection and notification, in contrast to the link-level probing and network-wide flooding common in today's intradomain routing protocols. This leads to simpler, cheaper routers.

Local adaptation to path failures: Upon detecting path failures, the ingress router rebalances the traffic on the remaining paths, based only on which path(s) failed—*not* on load information. This avoids having the routers distribute real-time updates about link load and prevents instability. Instead, the management system *precomputes* the reactions to path failures and configures the routers accordingly.

The first two features—multiple precomputed paths and path-level monitoring—are ideas that have been surfacing (sometimes implicitly) in the networking literature over the past few years (e.g., [18,52,70,100], and many others). Our architecture combines these two ideas in a new way, through (i) a specific proposal for the "division of labor" between the routers and the management system and (ii) an integrated view of traffic engineering and failure recovery within a single administrative domain. To support the simple network elements, the management system makes *network-wide* decisions based on the expected traffic, the network topology, and the groups of links that can fail together. The management system does *not* need to make these decisions in real time—quite the contrary, offline algorithms can compute the paths and the adaptations to path failures.

Our architecture raises important questions about (i) what configuration state the routers should have to drive their local reactions to path failures and (ii) how the management system should compute this state, and the underlying paths, for good traffic engineering and failure recovery. In addressing these questions, we make four main contributions:

Simple architecture for joint TE and failure recovery (Section 4.2): We propose a joint solution for traffic engineering and failure recovery, in contrast to today's networks that handle these problems separately. Our minimalist control plane has routers balance load based only on path-failure information, in contrast to recent designs that require routers to disseminate link-load information and compute new path-splitting parameters in real time [62, 68].

Network-wide optimization across failure scenarios (Section 4.3): We formulate and solve network-wide optimization problems for configuring the routers. Our algorithms compute (i) multiple paths that distribute traffic efficiently under a range of failure scenarios and (ii) the state for each ingress router to adapt to path failures. We present algorithms for two router designs that strike a different trade-off between router state and load-balancing performance.

Experiments with measurement data from a large ISP (Section 4.4): We evaluate our algorithms on measurement data from a tier-1 ISP network. Our simulation achieves a high degree of accuracy by utilizing the real topology, link capacities, link delays, hourly traffic matrices, and Shared Risk Link Groups (SRLGs) [51]. Our experiments show that one of our candidate router designs achieves near-optimal load balancing across a wide range of failure scenarios, even when the traffic demands change *dynamically*.

Deployability in ISP and data-center networks (Section 4.5): While our architecture enables simpler routers and switches, existing equipment can support our solutions. ISP backbones can use RSVP to signal multiple MPLS [82] paths, with hash-based splitting of traffic over the paths. In data centers, the fabric controller can configure multiple paths through the network, and the server machines can encapsulate packets to split traffic in the desired proportions.

The chapter ends with related work in Section 4.6, supporting proofs in Section 4.7 and conclusion in Section 4.8.

4.2 Simple Network Architecture

Our architecture uses simple, cheap routers to balance load before, during, and after failures by placing most functionality in a management system that performs offline optimization. The network-management system computes multiple diverse paths between each pair of edge routers, and tells each ingress router how to split traffic over these paths under a range of failure scenarios. Each edge router simply detects path-level failures and uses this information to adjust the splitting of traffic over the remaining paths, as shown in Figure 4.1. The main novel feature of our architecture is the way routers split traffic over the working paths; we propose two approaches that introduce a trade-off between router state and load-balancing performance.

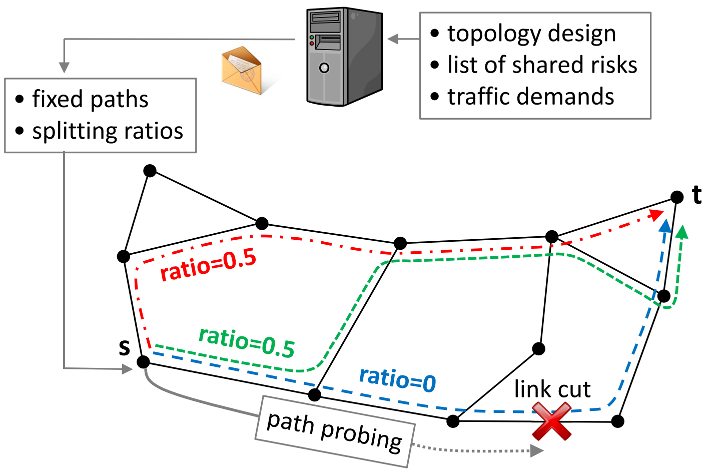


Figure 4.1: The management system calculates a fixed set of paths and splitting ratios, based on the topology, traffic demands, and potential failures. The ingress router learns about path failures and splits traffic over the remaining paths, based on pre-configured splitting ratios.

4.2.1 Precomputed Multipath Routing

Many routing protocols compute a single path between each router pair, and change that path in response to topology changes. However, dynamic routing has many downsides, including the overhead on the routers (to disseminate topology information and compute paths) and the transient disruptions during convergence. Techniques for making convergence faster tend to increase the complexity of the routing software and the overhead on the routers, by disseminating more information or updating it more quickly. Rather than reducing convergence time, or adding mechanisms to detect transient loops and blackholes, we avoid dynamic routing protocols entirely [18].

Our architecture uses multiple *preconfigured* paths between each pair of edge routers, allowing ingress routers to adapt to failures by shifting traffic away from failed path(s). With multiple paths through the network, the routers do not need to recompute paths dynamically—they simply stop using the failed paths until they start working again. This substantially reduces router software complexity and protocol overheads (e.g., bandwidth and CPU resources), while entirely side-stepping the problem of convergence. Instead, the management system computes these paths, based on *both* traffic-engineering and failure-recovery goals, and installs the paths in the underlying routers. The management system can select diverse paths that ensure connectivity in the face of failures, including multiple correlated failures.

Using multiple paths also leads to better load balancing, whether or not failures occur. Today's shortest-path routing protocols (like OSPF and IS-IS) use a single path, or (at best) only support *even* splitting of traffic over multiple *shortest* paths. Our architecture (like other recent proposals for multipath load balancing [29, 52, 60, 98]) allows flexible splitting of traffic over multiple paths. However, we do not require dynamic adaptation of the traffic splitting. Instead, the ingress router has a simple static configuration that determines the splitting of traffic over the available paths, while intermediate routers merely forward packets over pre-established paths. The management system optimizes this configuration *in advance* based on a network-wide view of the expected traffic and likely failures. This avoids the protocol overhead and stability challenges of distributed, load-sensitive routing protocols. Also, the management system can use knowledge about shared risks and anticipated traffic demands—information the routers do not have.

4.2.2 Path-Level Failure Detection

Most routing protocols detect failures by exchanging "hello" messages between neighboring routers and flooding the topology changes through the network. This approach requires small timers for fast failure detection, imposing additional overhead on the routers. In addition, many failures are triggered by planned maintenance [66], leading to *two* convergence events—one for the link failure(s), and another for the recovery—that both cause transient disruptions. In addition, "hello" messages do not detect all kinds of failures—some misconfigurations (e.g., having a maximum packet size that is too small) and attacks (e.g., an adversary selectively dropping packets) do not lead to lost "hello" messages.

Instead, our architecture relies on *path-level* failure detection. Each ingress-egress router pair has a session to monitor each of its paths (e.g., as in BFD [56]). The probes can be piggybacked on existing data traffic, obviating the need for separate "hello" messages when the path is carrying regular data traffic. This enables fast failure detection without introducing extra probe traffic, and the "implicit probes" provide a more realistic view of the reliability of a path [10, 39], since the packets vary in size, addresses, and so on. Another advantage is that the packets are handled by the hardware interfaces and, as such, do not consume processing resources (or experience software processing delays) at intermediate routers. (Still, the propagation delay along a path does impose limits on detection time in large topologies, an issue we discuss in more detail in Section 4.5.)

Although the ingress router doesn't learn which *link* failed, knowledge of the *path* failures is sufficient to avoid the failed path. In fact, since the routers need not be aware of the topology, no control protocol is needed to exchange topology information. In fact, only some of the ingress routers need to learn about the failure—only the routers that have paths traversing the failed edge. The other ingress routers, and the intermediate routers, can remain unaware of the link failure. Of course, the *management system* ultimately needs to know about topology changes, so failed equipment can be fixed or replaced. But this detection problem can be handled on a much longer timescale since it does not affect the failure-recovery time for data traffic.

4.2.3 Local Adaptation to Path Failures

In our architecture, a router is a simple device that does not participate in a routing protocol, collect congestion feedback, or solve any computationally difficult problems. Still, the routers do play an important role in adapting the distribution of traffic when paths fail or recover, at the behest of the management system. We propose two different ways the routers can split traffic over the working paths: (i) state-independent splitting which has minimal router state and (ii) state-dependent splitting which introduces more state in exchange for near-optimal performance, as summarized (and compared to an idealized solution) in Table 4.1.

Optimal load balancing: This idealized solution calculates the optimal paths and splitting ratios separately for each possible failure state, i.e., for each combination of link failures. This approach achieves the best possible load balancing. However, the approach is impractical because

	Optimal (Baseline)	**State-Dependent Splitting**	**State-Independent Splitting**
Router state	Exponential in total # of links	Exponential in # of pre-configured paths between two routers	Linear in # of pre-configured paths between two routers
Failure information	Link level	Path level	Path level
Optimality	Optimal	Nearly-optimal	Good
Time complexity	P	P if paths fixed	NP-hard

Table 4.1: Properties of the candidate solutions. The solutions differ in the amount of configuration state that must be stored in the routers, the information the routers must obtain about each failure, and the achieved traffic-engineering performance.

the routers must (i) store far too much state and (ii) learn about all link failures—even on links the router's paths do not traverse. Therefore, this solution would violate our architecture. However, the solution is still interesting as it provides a lower bound on the amount of congestion achievable by the other two schemes.

State dependent splitting: In this solution, each ingress router has a separate configuration entry with path-splitting weights for each combination of *path* failures to a particular egress router. For example, suppose a router has three paths to an egress router. Then, the router configuration contains seven entries—one for each of the $2^3 - 1$ combinations of path failures. Each configuration entry, computed ahead of time by the management system, consists of three weights—one per path, with a 0 for any failed paths. Upon detecting path failures, the ingress router inspects a pre-configured table to select the appropriate weights for splitting the traffic destined to the egress router. Our experiments in Section 4.4 show that, even in a large ISP backbone, having three or four paths is sufficient, leading to modest state requirements on the router in exchange for near-optimal load balancing.

State independent splitting: This solution further simplifies the router configuration by having a *single* set of weights across all failure scenarios. So, an ingress router with three paths to an egress router would have only *three* weights, one for each path. If any paths fail, the ingress router simply renormalizes the traffic on the remaining paths. As such, the management system must perform a *robust* optimization of the limited configuration parameters to achieve good load-balancing performance across a range of failure scenarios. Our experiments in Section 4.4 show that this simple approach can perform surprisingly well, but understandably not as well as state-dependent splitting.

4.3 Network-Wide Optimization

In our architecture, the network-management system performs network-wide optimization to compute paths and traffic-splitting ratios that balance load effectively across a range of failure scenarios. In this section, we first discuss the information the management system has about the network topology, traffic demands, and shared risks. Then, we explain how the management system computes the multiple diverse paths and the traffic-splitting ratios, for both state-dependent and state-independent splitting. We solve all optimization problems either by formulating them as convex optimizations solvable in polynomial time, or by providing heuristics for solving NP-hard problems. Table 4.2 summarizes the notation.

4.3.1 Network-Wide Visibility and Control

The management system computes paths and splitting ratios based on a network-wide view:

Fixed topology: The management system makes decisions based on the designed topology of the network—the routers and links that have been deployed. The topology is represented by a graph $G(V, E)$ with a set of vertices V and directed edges E. The capacity of edge $e \in E$ is denoted by c_e, and the propagation delay on the edge is y_e.

Shared risk link groups: The management system knows which links share a common vulnerability, such as connecting to the same line card or router or traversing the same optical

Variable	**Description**
$G(V, E)$	network with vertices V and directed edges E
c_e	capacity of edge $e \in E$
y_e	propagation delay on edge $e \in E$
S	family of network failure states
s	network failure state (set of failed links)
w^s	weight of network failure state $s \in S$
D	set of demands
u_d	source of demand $d \in D$
v_d	destination of demand $d \in D$
h_d	flow requirement of demand $d \in D$
P_d	paths available to demand $d \in D$
α_p	fraction of the demand assigned to path p
O_d	family of observable failure states for node u_d
$o_d(s)$	state observable by u_d in failure state $s \in S$
P_d^o	paths available to u_d in failure state $o \in O_d$
f_p^s	flow on path p in failure state $s \in S$
f_p^o	flow on path p in failure state $o \in O_d$
l_e^s	total flow on edge e in failure state s
$l_{e,d}^s$	flow of demand d on edge e in failure state s

Table 4.2: Summary of notation

fiber or amplifier [51]. The shared risks are denoted by the set S, where each $s \in S$ consists of a set of edges that may fail together. For example, a router failure is represented by the set of its incident links, a fiber cut is represented by all links in the affected fiber bundle, and the failure-free case is represented by the empty set $\emptyset$. Operators also have measurement data from past failures to produce estimates for the likelihood of different failures (e.g., an optical amplifier may fail less often than a line card). As such, each failure state s has a weight w^s that represents its likelihood or importance.

Expected traffic demands: The management system knows the anticipated traffic demands, based on past measurements and predictions of traffic changes. Each traffic demand $d \in D$ is represented by a triple (u_d, v_d, h_d), where $u_d \in V$ is the traffic source (ingress router), $v_d \in V$ is the destination (egress router), and h_d is the flow requirement (measured traffic). For simplicity, we assume that all demands remain connected for each failure scenario; alternatively, a demand can be omitted for each failure case that disconnects it. In practice, the management system may have a time sequence of traffic demands (e.g., for different hours in the day), and optimize the network configuration across all these demands, as we discuss in Section 4.4.3.

The management system's output is *set of paths* P_d for each demand d and the *splitting ratios* for each path. In each failure state s, the traffic splitting by ingress router u_d depends only on which *paths* have failed, not which failure scenario s has occurred; in fact, multiple failure scenarios may affect the same subset of paths in P_d. To reason about the handling of a particular demand d, we consider a set O_d of "observable" failure states, where each observable state $o \in O_d$ corresponds to a particular $P_d^o \subset P_d$ representing the available paths. For ease of expression, we let the function $o_d(s)$ map to the failure state observable by node u_d when the network is in failure state $s \in S$. The amount of flow assigned to path p in observable failure state $o \in O_d$ is f_p^o. The total flow on edge e in failure state s is l_e^s, and the flow on edge e corresponding to demand d is $l_{e,d}^s$.

The management system's goal is to compute paths and splitting ratios that minimize congestion over the range of possible failure states. A common traffic-engineering objective [35] is to minimize $\sum_{e \in E} \Phi(l_e^s / c_e)$ where l_e is the load on edge e and c_e is its capacity. $\Phi()$ could be a convex function of link load [35], to penalize the most congested links while still accounting for load on the remaining links. The final objective minimizing congestion across failure scenarios is

$$obj(l_{e_1}^{s_1}/c_{e_1}, ...) = \sum_{s \in S} w^s \sum_{e \in E} \Phi(l_e^s/c_e). \tag{4.1}$$

Minimizing this objective function is the goal of all the candidate solutions in the following section. The constraints that complete the problem formulation differ depending on the functionality placed in the underlying routers.

4.3.2 Computing Multiple Diverse Paths

The management system must compute multiple diverse paths that ensure good load balancing—and (most importantly) continued connectivity—across a range of failure scenarios. However, as shown later, computing the optimal paths for state-dependent and state-independent splitting is NP-hard. Instead, we propose a heuristic: using the collection of paths computed by the optimal solution that optimizes for each failure state independently. This guarantees that the paths are sufficiently diverse to ensure traffic delivery in all failure states, while also making efficient use of network resources.

The idealized optimal solution has a separate set of paths and splitting ratios in each failure state s. To avoid having variables for exponentially many paths, we formulate the problem in terms of the amount of flow $l_{e,d}^s$ from demand d traversing edge e for failure state s. The optimal edge loads are obtained by solving the convex optimization:

$$\begin{array}{lll} \min & obj(l_{e_1}^{s_1}/c_{e_1}, ...) & \\ \text{s.t.} & l_e^s = \sum_{d \in D} l_{e,d}^s & \forall s, e \\ & 0 = \sum_{i:e=(i,j)} l_{e,d}^s - \sum_{i:e=(j,i)} l_{e,d}^s & \forall d, s, j \neq u_d, v_d \\ & h_d = \sum_{i:e=(u_d,i)} l_{e,d}^s - \sum_{i:e=(i,u_d)} l_{e,d}^s & \forall d, s \\ & h_d = \sum_{i:e=(i,v_d)} l_{e,d}^s - \sum_{i:e=(v_d,i)} l_{e,d}^s & \forall d, s \\ & 0 \leq l_{e,d}^s & \forall d, s, e, \end{array} \tag{4.2}$$

where l_e^s and $l_{e,d}^s$ are variables. The first constraint defines the load on edge e, the second constraint ensures flow conservation, the third and fourth constraints ensure that the demands are met, and the last constraint guarantees flow non-negativity. An optimal solution can be found in polynomial time using conventional techniques for solving multi-commodity flow problems.

After obtaining the optimal flow on each edge for all the failure scenarios, we use a standard decomposition algorithm to determine the corresponding paths P_d and the flow f_p^s on each of them. The decomposition starts with a set P_d that is empty. New unique paths are added to the set by performing the following decomposition for each failure state s. First, annotate each edge e with the value $l_{e,d}^s$. Remove all edges that have 0 value. Then, find a path connecting u_d and v_d. Although we could choose any of the paths from u_d to v_d, our goal is to obtain paths that are as short as possible. So, if multiple such paths exist, we use the path p with the smallest propagation delay. Add this path p to the set P_d and assign to it flow f_p^s equal to the smallest value of the edges on path p. Reduce the values of these edges accordingly. Continue in this fashion, removing edges with zero value and finding new paths, until there are no remaining edges in the graph.

Note that we can show by induction that this process completely partitions the flow $l_{e,d}^s$ into paths. The decomposition yields at most $|E|$ paths for each network failure state s because the value of at least one edge becomes 0 whenever a new path is found. Hence the total size of the set P_d is at most $|E||S|$. It is difficult to obtain a solution that restricts the number of paths as we prove in Section 4.7 that it is NP-hard to solve problem (4.2) when the number of allowed paths is bounded by a constant J. In practice, the algorithm produces a relatively small number of paths between each pair of edge routers, as shown later in Section 4.4.

4.3.3 Optimizing the Traffic-Splitting Ratios

Once the paths are computed, the network-management system can optimize the path-splitting ratios for each ingress-egress router pair. The optimization problem and the resulting solution depend on whether the routers perform state-dependent or state-independent splitting.

State-Dependent Splitting

In state-dependent splitting, each ingress router u_d has a set of splitting ratios for each *observable* failure state $o \in O_d$. Since the path-splitting ratios depend on *which* paths in P_d have failed, the ingress router must store splitting ratios for $min(|S|, 2^{|P_d|})$ scenarios; fortunately, the number of paths $|P_d|$ is typically small in practice. When the network performs such *state-dependent splitting*, the management system's goal is to find a set of paths P_d for each demand and the flows f_p^o on these paths in all observable states $o \in O_d$. If the paths P_d are known and fixed, the problem can be formulated as a convex optimization:

$$\begin{aligned} \min \quad & obj(l_{e_1}^{s_1}/c_{e_1}, ...) \\ \text{s.t.} \quad & l_e^s = \sum_{d \in D} \sum_{p \in P_d^o, e \in p} f_p^o \quad && \forall e, s, o = o_d(s) \\ & h_d = \textstyle\sum_{p \in P_d^o} f_p^o && \forall d, o \in O_d \\ & 0 \leq f_p^o && \forall d, o \in O_d, p \in P_d, \end{aligned} \tag{4.3}$$

where l_e^s and f_p^o are variables. The first constraint defines the load on edge e, the second constraint guarantees that the demand d is satisfied in all observable failure states, and the last constraint ensures non-negativity of flows assigned to the paths. The solution of the optimization problem (4.3) can be found in polynomial time.

Finding the optimal set of paths $\{P_d\}$ in problem (4.3) is NP-hard. Section 4.7 shows that it is NP-hard to construct a path (if one exists) that allows the ingress router to distinguish the failure state s. This is required to decide how to best balance the load. All our formulations where the routers cannot directly observe link failures are NP-hard. Therefore, we use the paths that are found by the decomposition of the optimal solutions (4.2), as outlined in the previous subsection. Since these paths allow optimal load balancing for the optimal solutions (4.2), they are also likely to enable good load balancing for the optimization problem (4.3).

State-Independent Splitting

In state independent splitting, each ingress router has a *single* configuration entry containing the splitting ratios that are used under any combination of path failures. Each path p is associated with a splitting fraction α_p. When one or more paths fail, the ingress router u_d observes state o and uses $\frac{\alpha_p}{\sum_{q \in P_d^o} \alpha_q}$ as the splitting ratio for path p (and 0 for all the failed paths). If the network elements implement such *state-independent splitting*, and the paths P_d are known and fixed, the management system needs to solve the following non-convex optimization problem:

$$
\begin{aligned}
\min \quad & obj(l_{e_1}^{s_1}/c_{e_1}, \ldots) & & \\
\text{s.t.} \quad & f_p^o = h_d \frac{\alpha_p}{\sum_{q \in P_d^o} \alpha_q} & & \forall d, o \in O_d, p \in P_d \\
& l_e^s = \sum_{d \in D} \sum_{p \in P_d^o, e \in p} f_p^o & & \forall e, s, o = o_d(s) \\
& 0 \leq f_p^o & & \forall d, o \in O_d, p \in P_d,
\end{aligned}
\tag{4.4}
$$

where l_e^s, f_p^o and α_p are variables. The first constraint ensures that the flow assigned to every available path p is proportional to α_p. The other three constraints are the same as in (4.3).

Unfortunately, no standard optimization techniques allow us to compute an optimal solution efficiently, even when the paths P_d are fixed. Therefore, we have to rely on heuristics to find both the candidate paths P_d and the splitting ratios α_p. To find the set of candidate paths P_d, we again use the optimal paths obtained by decomposing (4.2). To find the splitting ratios we mimic the behavior of the optimal solution as closely as possible. We find the splitting ratios for all paths p by letting $\alpha_p = \sum_{s \in S} \frac{w^s f_p^s}{h_d}$ where f_p^s is the flow assigned by the optimal solution to path p in network failure state s. Since $\sum w^s = 1$, the calculated ratio is the weighted average of the splitting ratios used by the optimal solutions (4.2).

4.4 Experimental Evaluation

To evaluate the algorithms described in the previous section, we wrote a simulator in C++ that calls the CPLEX linear program solver in AMPL and solves the optimization problems (4.2) and (4.3). We compare our two heuristics to the optimal solution, a simple "equal splitting" configuration, and OSPF with the link weights set using state-of-the-art optimization techniques. We show that our two heuristics require few paths resulting in compact routing tables, and the round-trip propagation delay does not increase. Finally, using real traffic traces obtained during a 24-hour measurement in the network of a tier-1 ISP we show that our solutions achieve excellent results without the need to perform any reoptimizations even in the presence of a changing traffic matrix.

Our experimental results show that the **objective value of state-dependent splitting very closely tracks the optimal objective**. For this reason, this solution is our favorite.

Although state-independent splitting has somewhat worse performance especially as the network load increases beyond current levels, it is also attractive due to its simplicity.

4.4.1 Experimental Setup

Our simulations use a variety of synthetic topologies, the Abilene topology, as well as the city-level IP backbone topology of a tier-1 ISP with a set of failures provided by the network operator. The parameters of the topologies we used are summarized in Table 4.3.

Synthetic topologies: The synthetic topologies include 2-level hierarchical graphs, purely random graphs, and Waxman graphs. 2-level hierarchical graphs are produced using the generator GT-ITM [102], for random graphs the probability of two edges being connected is constant, and the probability of having an edge between two nodes in the Waxman graph decays exponentially with the distance of the nodes. These topologies also appear in [34].

Abilene topology: The topology of the Abilene network and a measured traffic matrix are used. We use the true edge capacities of 10 Gbps.

Tier-1 IP backbone: The city-level IP backbone of a tier-1 ISP is used. In our simulations, we use the real link capacities, link round-trip propagation delays, and measured traffic demands.

The collection of network failures S for the synthetic topologies and Abilene contains single edge failures and the no-failure case. Two experiments with different collections of failures are performed on the tier-1 IP backbone. In the first experiment, single edge failures are used. In the second experiment, the collection of failures also contains Shared Risk Link Groups (SRLGs), link

Name	**Topology**	**Nodes**	**Edges**	**Demands**
hier50a	hierarchical	50	148	2,450
hier50b	hierarchical	50	212	2,450
rand50	random	50	228	2,450
rand50a	random	50	245	2,450
rand100	random	100	403	9,900
wax50	Waxman	50	169	2,450
wax50a	Waxman	50	230	2,450
Abilene	backbone	11	28	110
tier-1	backbone	50	180	625

Table 4.3: Synthetic and realistic network topologies.

failures that occur simultaneously. SRLGs were obtained from the network operator's database that contains 954 failures with the largest failure affecting 20 links simultaneously. For each potential line card failure, a complete router failure, or a link cut there is a corresponding record in the SRLG database. Therefore, failures that do not appear in the database are rare. The weights w^s in the optimization objective (4.1) were set to 0.5 for the no-failure case, and all other failure weights are equal and sum to 0.5.

The set of demands D in the Abilene and tier-1 networks were obtained by sampling Netflow data measured on Nov. 15th 2005 and May 22nd 2009, respectively. For the synthetic topologies, we chose the same traffic demands as in [34].

To simulate the algorithms in environments with increasing congestion, we repeat all experiments several times while uniformly increasing the traffic demands. For the synthetic topologies we start with the original demands and scale them up to twice the original values. As the average link utilization in Abilene and the tier-1 topology is lower than in the synthetic topologies, we scale the demands in these realistic topologies up to three times the original value.

In our experiments we use the piecewise linear penalty function defined by $\Phi(0) = 0$ and its derivatives:

$$\Phi'(\ell) = \begin{cases} 1 & \text{for} \quad 0 \le \ell < 0.333 \\ 3 & \text{for} \quad 0.333 \le \ell < 0.667 \\ 10 & \text{for} \quad 0.667 \le \ell < 0.9 \\ 70 & \text{for} \quad 0.9 \le \ell < 1 \\ 500 & \text{for} \quad 1 \le \ell < 1.1 \\ 5000 & \text{for} \quad 1.1 \le \ell < \infty \end{cases}$$

This penalty function was introduced in [35], and allows one to formulate the optimizations (4.2) and (4.3) as linear programs by adding auxiliary variables. The function can be viewed as modeling retransmission delays caused by packet losses. The cost is small for low utilization, and increases steeply as the utilization exceeds 100%.

Our simulation calculates the objective value of the optimal solution, state-independent and state-dependent splitting, and equal splitting. Equal splitting is a variant of state-independent splitting that splits the flow evenly on the available paths. We also calculate the objective achieved by the shortest path routing of OSPF with optimized link weights. These link weights were calculated using the state-of-the-art optimizations of [34], and these optimizations take into consideration the set of failure states S and the corresponding failure weights w^s.

Our simulations were performed using CPLEX version 11.2 on a 1.5 GHz Intel Itanium 2 processor. Solving the linear program for (4.2) for a particular failure case in the tier-1 topology takes 4 seconds, and solving the linear program (4.3) takes about 16 minutes. A tier-1 network operator can perform calculations for its entire city-level topology in less than 2 hours.

4.4.2 Performance with Static Traffic

Avoiding congestion and packet losses during planned and unplanned failures is the central goal of traffic engineering. Our traffic engineering objective measures congestion across all the considered failure cases. The objective as a function of the scaled-up demands is depicted in Figure 4.2 and 4.3. The results which were obtained on the hierarchical and tier-1 topologies are representative, we made similar observations for all the other topologies. In Figure 4.2 and 4.3, the performance of state-dependent splitting and the optimal solution is virtually indistinguishable in all cases. State-independent splitting is less sophisticated and does not allow custom load balancing ratios for distinct failures, and therefore its performance is worse compared to the optimum. It is not surprising that the equal splitting algorithm achieves the worst performance.

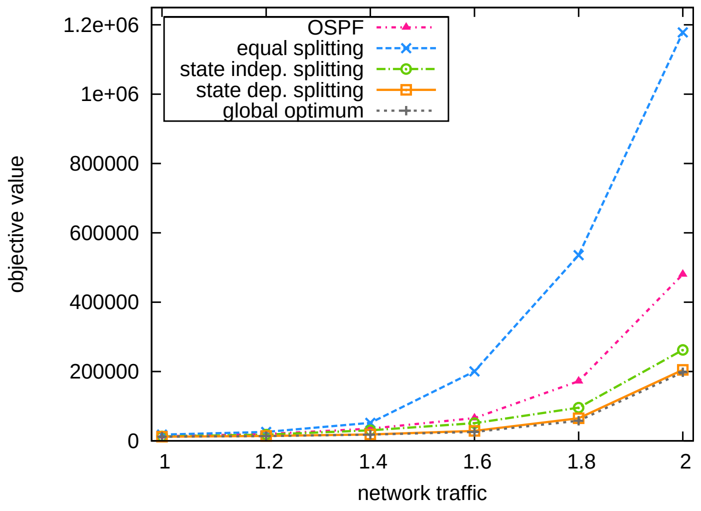
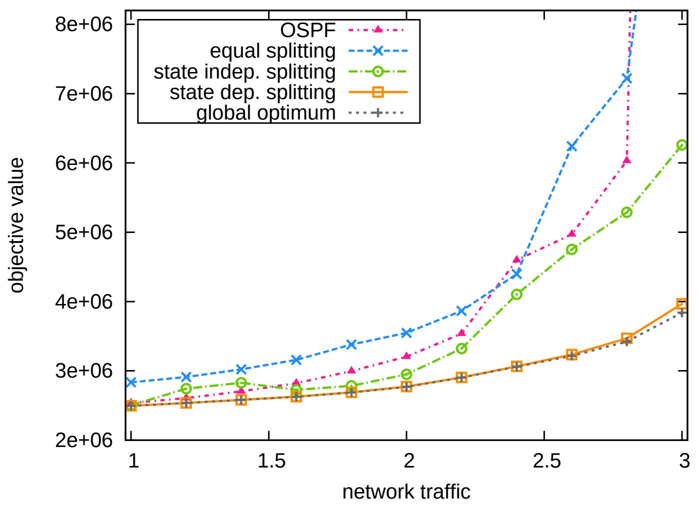


Figure 4.2: The traffic engineering objective as a function of an increasing traffic load in the hierarchical topology hier50a. The performance of the optimal solution and state-dependent splitting is nearly identical.



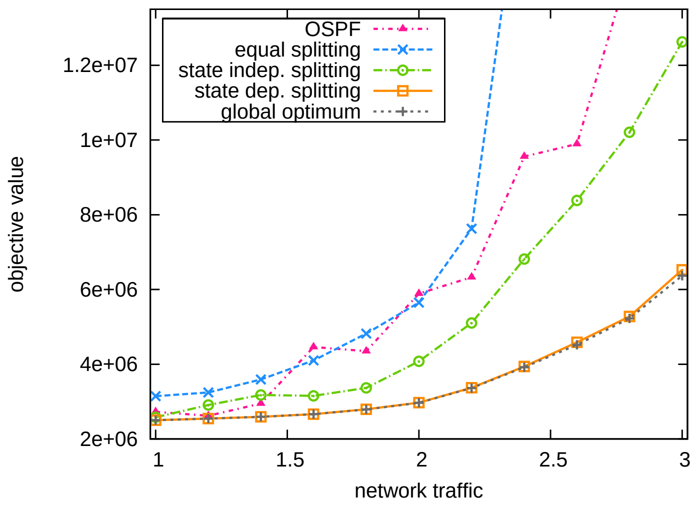
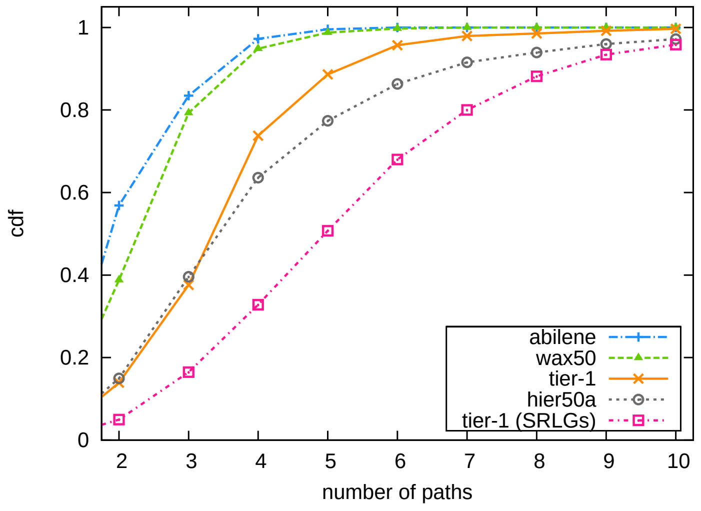


Figure 4.3: From top to bottom the traffic engineering objective as a function of an increasing traffic load in the tier-1 topology with single edge failures, and the tier-1 topology with SRLGs, respectively. The performance of the optimal solution and state-dependent splitting is nearly identical.

We observe that OSPF achieves a somewhat worse performance than state-independent and state-dependent splitting as the load increases. However, we should note that in OSPF, each router is restricted to sending all its traffic on the single path with the smallest weight, or splitting the traffic evenly if multiple smallest-weight paths exist. This approach does not allow the same flexibility in choosing routes and splitting ratios as our solution, and, therefore, OSPF should not be expected to achieve the same performance even for an optimal choice of OSPF link weights.



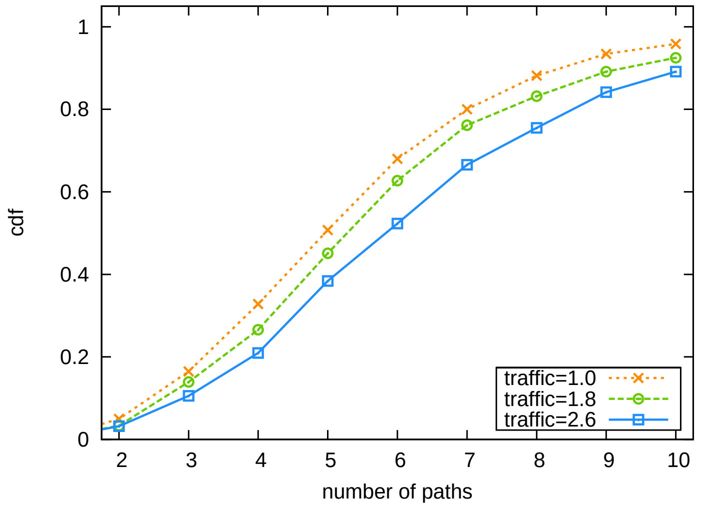


Figure 4.4: The number of paths used in various topologies at the top, and in the tier-1 topology with SRLGs at the bottom. The cumulative distribution function shows that the number of paths is almost independent of the traffic load in the network, but is larger for bigger, more well-connected topologies.

Solutions with few paths are preferred as they decrease the number of tunnels that have to be managed, and reduce the size of the router configuration. However, a sufficient number of paths must be available to avoid failures and to reduce congestion. We observe that the number of paths used by our algorithms is small. We record the number of paths used by each demand, and plot the distribution in Figure 4.4. Not surprisingly, the number of paths is greater for larger and more diverse topologies. 92% of the demands in the hierarchical topology use 7 or fewer paths, and fewer than 10 paths are needed in the tier-1 backbone topology for almost all demands. Further, Figure 4.4 shows that the number of paths only increases slightly as we scale up the amount of traffic in the networks. This small increase is caused by shifting some traffic to longer paths as the short paths become congested.

A practical solution uses few MPLS labels in order to reduce the size of routing tables in the routers. Our experimental results reveal that when we use MPLS tunnels in the tier-1 topology, a few thousand tunnels can pass through a single router. However, a simple routing table compression technique allows us to reduce the routing table size to a few hundred entries in each router. Such compression is important because it reduces the memory requirements imposed on the simple routers whose use we advocate, and it improves the route lookup time.

Routing tables can be compressed by using the same MPLS labels for routes with a common path to the destination. Specifically, if two routes to destination t pass through router r, and these routes share the same path between the router r and the destination t, the same outbound label should be used in the routing table of router r. The resulting routing table sizes as a function of the network load are depicted in Figure 4.5. The curve on the top shows the size of the largest routing table, and the curve on the bottom shows the average routing table size among all the backbone routers.

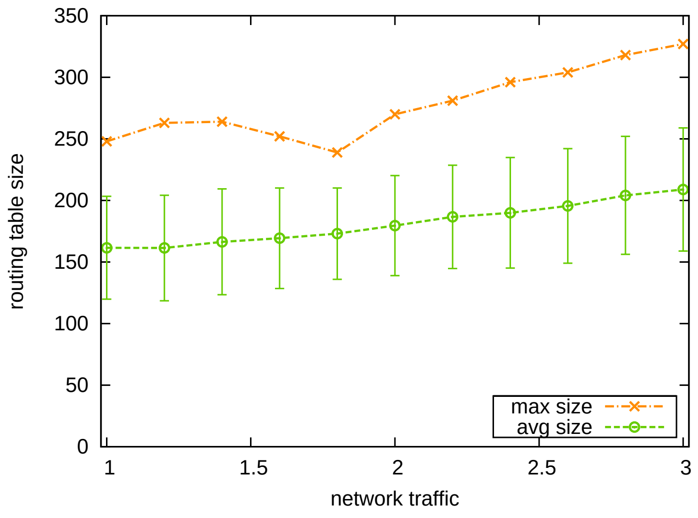


Figure 4.5: Size of the compressed routing tables in the tier-1 topology with SRLGs. The largest and average routing table sizes ($\pm$ one standard deviation) in the backbone routers are shown.

Minimizing the delay experienced by the users is another important goal of network operators. We calculated the average round-trip propagation delays of all the evaluated algorithms. The calculated delays include delays in all failure states weighted by the corresponding likelihood of occurrence, but exclude congestion delay which is negligible. The delays are summarized in Table 4.4. We observe that the round-trip delay of all algorithms except equal splitting is almost identical at around 31 ms. These values would satisfy the 37 ms requirement specified in the SLAs of the tier-1 network. Moreover, these values are not higher than these experienced by the network users today. To demonstrate this, we repeated our simulation on the tier-1 topology using the real OSPF weights which are used by the network operator. These values are chosen to provide a tradeoff between traffic engineering and shortest delay routing. The results which appear in Table 4.4 in the row titled OSPF (current) show that the current delays are 31.38 ms for each of the two tier-1 failure sets.

Algorithm	**Single edge**	**SRLGs**
Optimal load balancing	31.75 ± 0.26	31.80 ± 0.25
State dep. splitting	31.51 ± 0.17	31.61 ± 0.16
State indep. splitting	31.76 ± 0.26	31.87 ± 0.25
Equal splitting	34.83 ± 0.33	40.85 ± 0.86
OSPF (optimized)	31.18 ± 0.40	31.23 ± 0.40
OSPF (current)	31.38 ± 0	31.38 ± 0

Table 4.4: Round-trip propagation delay in ms (average $\pm$ one standard deviation) in the tier-1 backbone network for single edge failures and SRLG failures.

4.4.3 Robust Optimization for Dynamic Traffic

Solving the optimization problems repeatedly as the traffic matrix changes is undesirable due to the need to update the router configurations with new paths and splitting ratios. We explore the possibility of using a single router configuration that is robust to diurnal changes of the demands.

To perform this study we collected hourly Netflow traffic traces in the tier-1 network on September 29, 2009. We denote the resulting 24 hourly traffic matrices $D^0, D^1, ..., D^{23}$. Figure 4.6 depicts the aggregate traffic volume, as well as example of the traffic between three ingress-egress router pairs. The aggregate traffic volume is the lowest at 9 a.m. GMT and peaks with 2.5 times as much traffic at midnight and 8 p.m. GMT. Comparison to the three depicted ingress-egress router de-

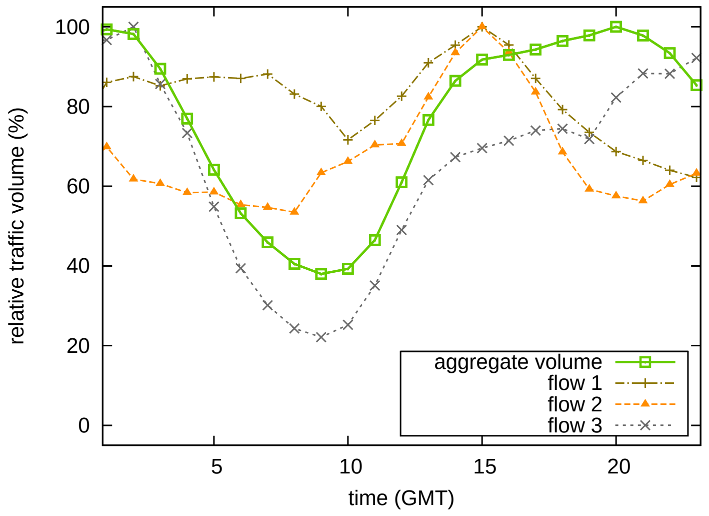


Figure 4.6: The aggregate traffic volume in the tier-1 network has peaks at midnight GMT and 8 p.m. GMT. Examples of three demands show that their peaks occur at different times of the day.

mands reveals that the traffic during a day cannot be obtained by simple scaling as the individual demands peak at different times. This makes the joint optimization challenging.

The first step in the joint optimization is to calculate a single set of paths that guarantee failure resilience and load balancing for each of the 24 traffic matrices. There are several approaches we can take. In the first approach, we solve the linear program for (2) for each traffic matrix D^i separately and use the union of the paths obtained for each matrix. The second approach is to calculate the average traffic matrix $D = \frac{1}{24}\sum_i D^i$. The linear program for (2) is then solved for the average traffic matrix. In the third approach we use the envelope of the 24 traffic matrices instead of the average, i.e., we let $D_{jk} = max_i D^i_{jk}$.

In our simulations we chose the last method. Compared to the first method, it results in fewer paths. Compared to the second method, it allows better load balancing because demands between ingress-egress pairs with high traffic variability throughout the day are represented by the peak traffic.

The second step is to calculate router configuration robust to traffic changes. We again use the envelope $D_{jk} = max_i D^i_{jk}$ as the input traffic matrix and repeat the optimizations from the previous section. Then we test the solution by simulating the varying traffic demand during one day period. The resulting objective value of state dependent splitting and state independent

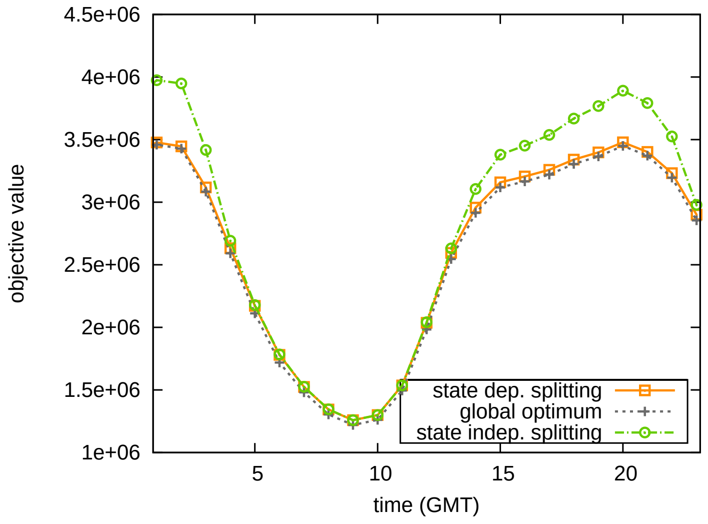


Figure 4.7: The traffic engineering objective in the tier-1 topology with SRLGs. The state dependent and state independent splitting algorithms use a single configuration throughout the day. The optimal solution uses a custom configuration for each hour.

splitting is depicted in Figure 4.7. The optimal objective in Figure 4.7 represents the performance of the best possible solution that uses custom configuration updated hourly. We observe that state dependent splitting with a single configuration is robust to diurnal traffic changes and the value of its objective closely tracks the optimum. State independent splitting is also close to optimal during low congestion periods, but becomes suboptimal during the peak hours.

4.5 Deployment Scenarios

Although our architecture enables the use of new simpler routers, we can readily deploy our solutions using existing protocols and equipment, as summarized in Table 4.5. An ISP can deploy our architecture using Multi-Protocol Label Switching (MPLS) [82]. Data centers could use the same solution, or leverage existing Ethernet switches and move some functionality into the end-host machines.

4.5.1 ISP Backbone Using MPLS

Installing MPLS paths with RSVP: MPLS is particularly suitable because ingress routers encapsulate packets with labels and direct them over pre-established Label-Switched Paths (LSPs).

	ISP Backbone	**Data Center**
Network element	MPLS router	Ethernet switch
Path installation	RSVP	VLAN trunking
Traffic splitting	Ingress router	End host
Failure detection	BFD	Host probing
Fast recovery	Ingress router	End host
Traffic demand	MPLS MIB	Host/VLAN counter

Table 4.5: Existing tools and protocols that can be used to deploy our architecture.

This enables flexible routing when multiple LSPs are established between each ingress-egress router pair. Our solution, then, could be viewed as a particular application of MPLS, where the management system computes the LSPs, instructs the ingress routers to establish the paths (say, using RSVP), and disables any dynamic recalculation of alternate paths when primary paths fail.

Hash-based splitting at ingress routers: Multipath forwarding is supported by commercial routers of both major vendors [3, 76]. The routers can be configured to hash packets based on port and address information in the headers into several groups and forward each group on a separate path. This provides path splitting with relatively fine granularity (e.g., at the 1/16th level), while preventing out-of-order packet delivery by ensuring that packets belonging to the same TCP or UDP flow traverse the same path.

Path-level failure detection using BFD: Fast failure detection can be done using Bidirectional Forwarding Detection (BFD) [56]. A BFD session can monitor each path between two routers, by piggybacking on the existing data traffic. (Backbones covering a large geographic region may also use existing link-level detection mechanisms for even faster recovery. For example, *local path protection* [77] installs a short alternate path between two adjacent routers, for temporary use after the direct link fails. However, local protection cannot fully exploit the available path diversity, leading to suboptimal load balancing; instead, local protection can be used in conjunction with our design.)

Failure recovery at ingress router: The ingress router adapts to path failures by splitting traffic over the remaining paths. In state-independent splitting, the ingress router has a single set of traffic-splitting weights, and automatically renormalizes to direct traffic over the working paths. State-dependent splitting requires modification to the router software to switch to alternate traffic-splitting weights in the data plane; no hardware modifications are required.

Measuring traffic demands using SNMP: MPLS has SNMP counters (called Management Information Bases) that measure the total traffic traversing each Label-Switched Path. The management system can poll these counters to measure the traffic demands. Alternative measurement techniques, such as Netflow or tomography, may also be used.

4.5.2 Data Center Using Hosts and Switches

While a data center could easily use the same MPLS-based solution, control over the end host and the availability of cheaper commodity switches enable another solution.

End-host support for monitoring and traffic splitting: The server machines in data centers can perform many of the path-level operations in our architecture. As in the VL2 [40] and SPAIN [70] architectures, the end host can encapsulate the packets (say, using a VLAN tag) to direct them over a specific path. This enables much finer-grain traffic splitting. In addition, the end host can perform path-level probing in the data plane, by piggybacking on existing data traffic and sending additional active probes when needed. Upon detecting path failures, the end host can change to new path-splitting percentages based on the precomputed configuration installed by the controller. The end host could also measure the traffic demands by keeping counts of the traffic destined to each egress switch. These functions can be implemented in the hypervisor, such as the virtual switch that often runs on server machines in data centers.

Multiple VLANs or OpenFlow rules for forwarding: The remaining functions can be performed by the underlying switches. For example, the management system can configure multiple paths by merging these paths into a set of trees, where each tree corresponds to a different VLAN [70]. Or, if the switches support the emerging OpenFlow standard [1,67], the management system could install a forwarding-table rule for each hop in each path, where the rule matches on the VLAN tag, and forwards the packet to the appropriate output port. Since OpenFlow switches maintain traffic counters for each rule, the management system can measure the traffic demands by polling the switches, in lieu of the end hosts collecting these measurements.

4.6 Related Work

Traffic engineering: Most of the related work treats failure recovery and traffic engineering independently. Traffic engineering without failure recovery in the context of MPLS is studied in [26, 29, 63, 87, 98]. The work in [26] utilizes traffic splitting to minimize end-to-end delay and loss rates; however, an algorithm for optimal path selection is not provided. The works in [63] and [87] minimize the maximum link utilization while satisfying the requested traffic demands. Other papers [29, 52, 60, 98] prevent congestion by adaptively balancing the load among multiple paths based on measurements of congestion, whereas our solution *precomputes* traffic-splitting configurations based on both the offered traffic and the likely failures.

Failure recovery: Local and global path protection are popular failure recovery mechanisms in MPLS. In local protection the backup path takes the shortest path that avoids the outage location from a point of local repair to the merge point with the primary path. The IETF RFC 4090 [77] focuses on defining signaling extensions to establish the backup paths, but leaves the issues of bandwidth reservation and optimal route selection open. In [95] the shortest path that avoids the failure is used. While [85] and [97] attempt to find optimal backup paths with the goal of reducing congestion, local path protection is less suitable for traffic engineering than global path protection, which allows rerouting on end-to-end paths [89]. Other work describes how to manage restoration bandwidth and select optimal paths [59, 64]. While our solution also uses global protection to reroute around failures, the biggest difference is that most of the related work distinguishes primary and backup paths and only uses a backup path when the primary path fails. In contrast, our solution balances the load across multiple paths even before failures occur, and simply adjusts the splitting ratios in response to failures.

Integrated failure recovery and TE: Some proposals only use alternate paths when primary routes fail [84], or they require explicit congestion feedback and do not provide algorithms to find the optimal paths [62, 68]. YAMR [36] constructs a set of diverse paths in the interdomain routing setting that are resilient against a specified set of failures, but without regard to load balancing. In [103] they integrate failure recovery with load balancing, but their focus is different—they guarantee delivery of a certain fraction of the traffic after a single edge failure, whereas our goal is to deliver *all* traffic for a known set of multi-edge failures. In [99] they propose an architecture that handles up to F link failures by using local rerouting, subject to link capacity constraints. Unlike [99], our work uses end-to-end routing, and does not require link state flooding and dynamic router reconfigurations. Proposals that optimize OSPF or IS-IS link weights with failures in mind, such as [34] and [72], must rely on shortest path IGP routing and therefore cannot fully utilize the path diversity in the network.

Failure recovery and TE with multiple spanning trees: Enterprise and data-center networks often use Ethernet switches, which do not scale well because all traffic flows over a single spanning tree, even if multiple paths exist. Several papers propose more scalable Ethernet designs that use multiple paths. The work of Sharma et al. uses VLANs to exploit multiple spanning trees to improve link utilization, and achieve improved fault recovery [88]. Most of the designs such as VL2 [40], and PortLand [71] rely on equal splitting of traffic on paths with the same cost. SPAIN [70] supports multipath routing through multiple spanning trees, with end hosts splitting traffic over the multiple paths. However, the algorithm for computing the paths does not consider the traffic demands, and the end hosts must play a stronger role in deciding which path to use for each individual flow based on the observed performance.

NP-hardness: Hardness proofs of optimization problems related to failure recovery appear, e.g., in [94] and [25].

4.7 Proofs

This section shows that two problems are NP-hard:

Failure State Distinguishing

instance: A directed graph $G = (V, E)$, source and destination vertices $u, v \in V$, and a sets $s \subseteq E$.

question: Is there a simple directed path P from u to v that is up if the edges in s do not fail and down if they fail?

Bounded Path Load Balancing

instance: A directed graph $G = (V, E)$ with a positive rational capacity c_e for each edge $e \in E$, a collection S of subsets $s \subseteq E$ of *failure states* with a rational weight w^s for each $s \in S$, a set of triples (u_d, v_d, h_d), $1 \leq d \leq k$, corresponding to *demands*, where h_d units of demand d need to be sent from source vertex $u_d \in V$ to destination vertex $v_d \in V$, an integer bound J on the number of paths that can be used between any source-destination pair, a piecewise-linear increasing cost function $\Phi(\ell)$ mapping edge loads ℓ to rationals, and an overall cost bound B.

QUESTION: Are there J (or fewer) paths between each source-destination pair such that the given demands can be assigned to the paths so that the cost (sum of $\Phi(\ell)$ over all edges and weighted failure states as described in the text) is B or less?

To prove that a problem X is NP-hard, we must show that for some known NP-hard problem Y, any instance y of Y can be transformed into an instance x of X in polynomial time, with the property that the answer for y is yes if and only if the answer for x is yes. Both our problems can be proved NP-hard by transformations from the following problem, proved NP-hard by Fortune, Hopcroft, and Wyllie [33].

DISJOINT DIRECTED PATHS

INSTANCE: A directed graph $G(V, E)$ and distinguished vertices $u_1, v_1, u_2, v_2 \in V$.

QUESTION: Are there directed paths P_1 from u_1 to v_1 and P_2 from u_2 to v_2 such that P_1 and P_2 are vertex-disjoint?

Theorem 1. The FAILURE STATE DISTINGUISHING problem is NP-hard.

Proof. Suppose we are given an instance $G = (V, E), u_1, v_1, u_2, v_2$ of DISJOINT DIRECTED PATHS. Our constructed instance of FAILURE STATE DISTINGUISHING consists of the graph $G' = (V, E')$, where $E' = E \cup \{(v_1, u_2)\}$, with $u = u_1$, $v = v_2$, and $s = \{(v_1, u_2)\}$.

Given this choice of s, a simple directed path from u to v that is up only if the edge (v_1, u_2) is up must contain that edge. We claim that such a path exists if and only if there are vertex-disjoint directed paths P_1 from u_1 to v_1 and P_2 from u_2 to v_2. Suppose a distinguishing path P exists. Then it must consist of of three segments: a path P_1 from $u = u_1$ to v_1, the edge (v_1, u_2), and then a path P_2 from u_2 to $v = v_2$. Since it is a simple path, P_1 and P_2 must be vertex-disjoint. Conversely, if vertex-disjoint paths P_1 from u_1 to v_1 and P_2 from u_2 to v_2 exist, then the path P that concatenates P_1 followed by (v_1, u_2) followed by P_2 is our desired distinguishing path. ■

Theorem 2. The BOUNDED PATH LOAD BALANCING problem is NP-hard even if there are only two commodities ($k = 2$), only one path is allowed for each ($J = 1$), and there is only one failure state s.

Proof. For this result we use the variant of DISJOINT DIRECTED PATHS in which we ask for edge-disjoint rather than vertex-disjoint paths. The NP-hardness of this variant is easy to prove, using a construction in which each vertex x of G is replaced by a pair of new vertices in_x and out_x connected by the edge (in_x, out_x), and each edge (x, y) of G is replaced by the edge (out_x, in_y).

Suppose we are given an instance $G = (V, E), u_1, v_1, u_2, v_2$ of the edge-disjoint variant of DISJOINT DIRECTED PATHS. Our constructed instance of BOUNDED PATH LOAD BALANCING is based on the same graph, with each edge e given capacity $c_e = 1$, with the single failure state $s = \phi$ (i.e., the state with no failures), with $w^s = 1$, and with demands represented by $(u_1, v_1, 1)$ and $(u_2, v_2, 1)$. The cost function Φ has derivative $\Phi'(\ell) = 1$, $0 \leq \ell \leq 1$, and $\Phi'(\ell) = |E| + 1$, $\ell > 1$. Our target overall cost bound is $B = |E|$.

If the desired disjoint paths exist, we can use P_1 to send the required unit of traffic from u_1 to v_1, and P_2 to send the traffic from u_2 to v_2. Since the paths are edge-disjoint, no edge will carry more than one unit of traffic, so the cost per edge used is 1, and the total number of edges used is at most $|E|$. Thus the specified cost bound $B = |E|$ is met. On the other hand, if no such pair of paths exist, then we must choose paths P_1 and P_2 that share at least one edge, which will carry two units of flow, for a cost of at least $|E| + 1$, just for that edge. Thus if there is a solution with cost $|E|$ or less, the desired disjoint paths must exist. ■

Adding more paths, failure states, or commodities cannot make the problem easier. Note, however, that this does not imply that the problem for the precise cost function Φ presented in the text is NP-hard. It does, however, mean that, assuming P $\neq$ NP, any efficient algorithm for that Φ would have to exploit the particular features of that function.

4.8 Summary

In this chapter we propose a mechanism that combines path protection and traffic engineering to enable reliable data delivery in the presence of link failures. We formalize the problem by providing several optimization-theoretic formulations that differ in the capabilities they require of the network routers. For each of the formulations, we present algorithms and heuristics that allow the network operator to find a set of optimal end-to-end paths and load balancing rules.

Our extensive simulations on the IP backbone of a tier-1 ISP and on a range of synthetic topologies demonstrate the attractive properties of our solutions. First, state-dependent splitting achieves load balancing performance close to the theoretical optimum, while state-independent splitting often offers comparable performance and a very simple setup. Second, using our solutions does not significantly increase propagation delay compared to the shortest path routing of OSPF. Finally, our solution is robust to diurnal traffic changes and a single configuration suffices to provide good performance.

In addition to failure resilience and favorable traffic engineering properties, our architecture has the potential to simplify router design and reduce operation costs for ISPs as well as operators of data centers and enterprise networks.

Chapter 5

Conclusion

This dissertation suggests three methods to improve reliability of the Internet by analyzing three major weaknesses of state-of-the-art network routing protocols.

First, we introduced a new model, the Dynamic Path Vector Protocol (DPVP), that provides a convenient and accurate framework for understanding the dynamic properties of interdomain routing. Using this model, we resolved a decade-long open question that asks for the necessary and sufficient conditions under which routing policies are safe. This in turn allowed us to design the DeCoy algorithm that verifies safety of routing policies used in practice in polynomial time. Policy induced oscillations are responsible for temporary losses of connectivity, creation of transient loops leading to packet losses, and frequent route changes have negative impact on TCP performance. Therefore, deployment of the DeCoy algorithm in the Internet has the potential to significantly improve user perceived performance.

Second, we studied the security of interdomain routing. We started with the observation that current proposals of secure interdomain routing protocols cannot yield measurable security benefits in small scale deployments, and network operators do noth have incentives for early adoption. We identified a combination of mechanisms that provides significant and measurable security benefits to participants and non-participants alike even if the proposed solution is only deployed by a small group of participants. If deployed in the Internet, our solution can significantly reduce the occurrence of route blackholing and data interception caused either by malicious attacks or accidental misconfigurations.

Third, we introduced a new simple network architecture that allows load balancing in the presence of equipment failures in the networks of ISPs. Our solution is attractive because a simple static preconfiguration of the network routers is sufficient to provide them with enough information to perform well under a wide set of failures. Static configuration is also sufficient even for varying traffic loads. The ability of our solution to select optimal routes without the need to perform time-consuming reoptimizations can be translated to lower network management complexity, and improved performance during the transient periods after each failure.

Several questions remain open:

- Our analysis of BGP safety is restricted to cases where the route preference function follows the rule of independent ranking. However, as Griffin and Wilfong observed [41], iBGP configurations that use the Multi Exit Discriminator (MED) violate the rule of independent ranking, i.e., a route's ranking can vary depending on the presence or absence of other routes. How can we extend the DPVP model to analyze such cases?

- The DeCoy algorithm is a convenient tool that uses AS policies on the input and provides policy conflicts on the output. However, routing policies are kept secret and ASes are not willing to share them. What should a privacy preserving version of the DeCoy algorithm look like? While it is certainly possible to apply standard techniques to make the algorithm privacy preserving, what is the communication overhead, and would ASes be willing to use the algorithm if some information about their policies can be inferred from the output of the algorithm itself?

- Another open question concerns the action that ASes should take once, through the use of the DeCoy algorithm, they detect policy conflicts. If a policy conflict can be resolved if one of several ASes selects a higher cost route, how can the ASes agree who should adopt the change?

In conclusion, we believe that deployment of the algorithms and techniques proposed in this dissertation would result in a significant improvement of the reliability of the Internet. Moreover, most of our techniques are general enough to be used as building blocks, and we hope that they will be used to make future routing protocols more safe, secure, and reliable.

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