

INTERDOMAIN ROUTING AND GAMES

HAGAY LEVIN*, MICHAEL SCHAPIRA†, AND AVIV ZOHAR‡

Abstract. We present a game-theoretic model that captures many of the intricacies of *interdomain routing* in today’s Internet. In this model, the strategic agents are source nodes located on a network, who aim to send traffic to a unique destination node. The interaction between the agents is dynamic and complex – asynchronous, sequential, and based on partial information. Best-reply dynamics in this model capture crucial aspects of the de facto standard interdomain routing protocol, namely the Border Gateway Protocol (BGP).

We study complexity and incentive-related issues in this model. Our main results show that in realistic and well-studied settings, BGP is incentive-compatible. I.e., not only does myopic behaviour of all players *converge* to a “stable” routing outcome, but no player has motivation to unilaterally deviate from BGP. Moreover, we show that even *coalitions* of players of *any* size cannot improve their routing outcomes by collaborating. Unlike the vast majority of works in mechanism design, our results do not require any monetary transfers (to or by the agents).

Key words. Interdomain Routing, Border Gateway Protocol (BGP), Security, Distributed Algorithmic Mechanism Design, Communication Complexity

AMS subject classifications. 91A40, 68M12

1. Introduction. The Internet is composed of smaller networks called *Autonomous Systems (ASes)*. ASes are owned by selfish, often competing, economic entities (Microsoft, AT&T, etc.). The task of establishing routes between ASes is called *interdomain routing*. Since not all ASes are directly connected, packets often have to traverse several ASes. The packets’ routes are established via complex interactions between ASes that enable them to express preferences over routes, and are affected by the nature of the network (message delays, malfunctions, etc.). The standard interdomain routing protocol *de facto* is the *Border Gateway Protocol (BGP)*.

Routing Games. The first contribution of this paper is the presentation of a game-theoretic model of interdomain routing that captures many of its intricacies (e.g., the asynchronous nature of the network). In our model (as in [16, 15]), the network is defined by an undirected graph $G = (N, L)$. The set of nodes N represents the ASes, and consists of n *source-nodes* $1, \dots, n$ (the players), and a unique *destination-node* d .¹ The set of edges L represents physical communication links between the nodes. Each source node i has a valuation function v_i that expresses a full-order of strict preferences over simple routes from i to d .

The model consists of two games: At the heart of the model is the sequential, asynchronous, and private-information CONVERGENCE GAME, which is meant to model interdomain routing dynamics. Best-reply dynamics in the CONVERGENCE GAME model crucial features of BGP dynamics, in which each AS is instructed to continuously execute the following actions:

*The Department of Economics, The Hebrew University of Jerusalem, Israel. hagayl@mscc.huji.ac.il.

†The School of Computer Science and Engineering, The Hebrew University of Jerusalem, Israel. mikesch@cs.huji.ac.il. Supported by grants from the Israel Science Foundation and the USA-Israel Bi-national Science Foundation.

‡The School of Computer Science and Engineering, The Hebrew University of Jerusalem, Israel. avivz@cs.huji.ac.il. Supported by a grant from the Israel Science Foundation.

¹The reason for this is that in today’s Internet, routes to each destination AS are computed independently.

- Receive update messages from neighbouring nodes announcing their routes to the destination.
- Choose a single neighbouring node whose route you prefer most (given v_i) to send traffic to.
- Announce your new route to all neighbouring nodes.

We also define a ONE-ROUND GAME, which will function as an analytic tool. The ONE-ROUND GAME can be regarded as the full-information non-sequential game underlying the CONVERGENCE GAME. Pure Nash equilibria in the ONE-ROUND GAME correspond to “stable solutions” in networking literature [16, 15], and are the “sinks” to which best-reply dynamics (BGP) can converge.

We study several complexity and strategic problems in this model. Most importantly, we address the issue of incentive-compatibility of best-reply dynamics in the CONVERGENCE GAME. We provide realistic settings in which the execution of best-reply dynamics (BGP) is in the best-interest of the players (ASes). We also address the following questions: How hard it is to establish whether a pure Nash exists in the ONE-ROUND GAME (the nonexistence of a pure Nash implies that best-reply dynamics will go on indefinitely)? How hard is it to get good approximations to the optimal social welfare?

Existence of Pure Nash Equilibria. Griffin and Wilfong have shown that determining whether a pure Nash equilibrium in the ONE-ROUND GAME (stable solution) exists is NP-hard [16]. We prove that this result extends to the communication model.

Theorem: *Determining whether a pure Nash equilibrium in the ONE-ROUND GAME exists requires exponential communication (in n) between the source-nodes.*

BGP Convergence and Incentives. Networking researchers, and others, invested a lot of effort into identifying sufficient conditions for the existence of a stable solution to which BGP *always* converges (see, e.g., [15, 30, 12, 11, 14, 5, 29, 4]). The most general condition known to guarantee this is “No Dispute Wheel”, proposed by Griffin, Shepherd, and Wilfong [15]. No Dispute Wheel guarantees a *unique* pure Nash in the ONE-ROUND GAME, and convergence of best-reply dynamics to it in the CONVERGENCE GAME. No Dispute Wheel allows nodes to have significantly more expressive and realistic preferences than always preferring shorter routes to longer ones. In particular, a special case of No Dispute Wheel is the celebrated *Gao-Rezford setting* [12, 11] that is said to depict the commercial structure that underlies the Internet [21] (see Section 2 for an explanation about No Dispute Wheel and interesting special cases).

Feigenbaum, Papadimitriou, Sami, and Shenker [7] initiated an economic, or mechanism design, approach to interdomain routing. While BGP was designed to guarantee connectivity between trusted and obedient parties, in the age of commercial Internet these are no longer valid assumptions (ASes are owned by different economic entities with very different, and often conflicting, commercial interests). Identifying realistic settings in which BGP is incentive-compatible has become the paradigmatic problem in Distributed Algorithmic Mechanism Design (see [10] and references therein), and is the subject of many works [28, 6, 9, 8, 3, 10, 24, 17].

Recently, a step in this direction was taken in [8, 10]. It was shown that if No Dispute Wheel and an additional condition named *Policy Consistency* hold then BGP is incentive-compatible in ex-post Nash (see [28] for a formal definition of the ex-post Nash solution concept and an explanation about its suitability for distributed computation environments). Informally, policy consistency means that *no* two neighbouring

nodes disagree over which of *any* two routes is preferable. This is obviously a very severe restriction that does not necessarily hold in practice. We take a significant step forward by removing it (in particular, we allow the Gao-Rexford commercial setting in which Policy Consistency does not necessarily hold).

Unfortunately, we prove that best-reply dynamics are *not* incentive-compatible if Policy Consistency does not hold. This is true even if No Dispute Wheel holds, and can be shown to hold even in the Gao-Rexford setting.

Theorem: *Best-reply dynamics are not incentive-compatible in ex-post Nash even if the No Dispute Wheel condition holds.*

However, there is still hope for BGP. We define a property called “*Route Verification*”. Route Verification means that a node can verify whether a route announced by a neighbouring node is indeed available to that neighbouring node (and if not simply ignore that route announcement). Unlike Policy Consistency, Route Verification *does not* restrict the preferences of ASes, but is achieved by modifying BGP (e.g., this can be achieved via cryptographic signatures). Achieving Route Verification in the Internet is an important agenda in security research². Security researchers seek ways to implement Route Verification that are not only theoretically sound, but also reasonable to deploy in the Internet (see [1]).

We note that even if announcements of non-available routes are prevented by Route Verification, nodes still have many other forms of manipulation available to them: Pretending to have different preferences (“lying”), conveying inconsistent information (e.g., displaying inconsistent preferences over routes), denying routes from neighbours by not reporting a route, and more. Hence, it still needs to be shown that Route Verification guarantees immunity of best-reply dynamics (BGP) to *all* forms of manipulation.

Our main result is the following:

Theorem: *Best-reply dynamics are incentive-compatible in (subgame perfect³) ex-post Nash if No Dispute Wheel and Route Verification hold.*

We stress that this result is achieved *without any monetary transfers* between nodes (as in [10], and unlike most prior works on interdomain routing, and in mechanism design in general).

Our result highlights an interesting connection between the two current research agendas that address the problem of disobedience and lack of trust in interdomain routing – security research and Distributed Algorithmic Mechanism Design. One of the implications of this result is that one can achieve incentive-compatibility in realistic settings (e.g., networks for which the Gao-Rexford constraints and Route Verification hold). This should further motivate security research, as it provides a strong strategic justification for modifications of BGP that guarantee Route Verification via cryptographic and other means (e.g., Secure BGP [1, 27]).

In [10] the notion of *collusion-proofness in ex-post Nash* is defined. Informally, collusion-proofness in ex-post Nash means that a *group* of agents cannot collaborate to strictly improve the outcome of any player in the group without strictly harming another player in the group. This means that some of the members of the group have

²“The US government cites BGP security as part of the national strategy for securing the Internet [Department of Homeland Security 2003]” [1]

³A strategy profile is a subgame perfect equilibrium if it represents an equilibrium of every subgame of the original game.) ex-post Nash if No Dispute Wheel and Route Verification hold.

no interest to deviate from the strategy profile (at least one member will be harmed by doing so). The previous theorem can actually be strengthened to the following one:

Theorem: *Best-reply dynamics are collusion-proof in (subgame perfect) ex-post Nash if No Dispute Wheel and Route Verification hold.*

In particular, this holds even for the coalition that contains *all* nodes in the network. This implies that, if No Dispute Wheel holds, BGP is actually socially-reasonable in the following sense (see [3]):

Corollary: *If No Dispute Wheel holds, the unique Nash equilibrium in the ONE-ROUND GAME (to which best-reply dynamics always converge) is Pareto optimal.*

Maximizing Social Welfare. Finally, we turn our attention to the objective of maximizing the social-welfare, that has also been studied in the context of interdomain routing (see [9]). Maximizing the social welfare means finding a routing tree rooted in d , $T = R_1, \dots, R_n$, in which node i is assigned route R_i , such that $\sum_i v_i(R_i)$ is maximized. In [8] it is shown that if No Dispute Wheel and Policy Consistency hold then BGP converges to a stable solution that also maximizes the social welfare. In contrast, we show that the removal of Policy Consistency can be disastrous in terms of welfare maximization. We do so by presenting two complementary bounds, one in the computational complexity model, and one in the communication complexity model.

Theorem: *Obtaining an approximation of $O(n^{\frac{1}{2}-\epsilon})$ to the optimal social welfare is impossible unless $P = NP$. Obtaining an approximation of $O(n^{1-\epsilon})$ to the social welfare is impossible unless $P = ZPP$. This holds for any $\epsilon > 0$ and even if No Dispute Wheel holds.*

Theorem: *Obtaining an approximation of $O(n^{1-\epsilon})$ to the optimal social welfare requires exponential communication (in n). This holds for any $\epsilon > 0$ and even if No Dispute Wheel holds.*

In fact, the above two inapproximability results hold even in the Gao-Rexford setting (that is a special case of No Dispute Wheel). These results should be compared with the previously known lower bound of $\Omega(n^{\frac{1}{2}-\epsilon})$ [9] (dependent on $P \neq ZPP$) for the case of general (*i.e.*, unrestricted) ASes' valuation functions. Our lower bounds show that even restrictive conditions (Gao-Rexford) that ensure existence of pure Nash in the ONE-ROUND GAME, and convergence of best-reply dynamics in the CONVERGENCE GAME, might be very far from guaranteeing good social welfare. A trivial matching upper bound of n exists even for general valuation functions (simply assign the node with the highest value for some route its most desired route).

1.1. Organization of the Paper. In Section 2 we present the model and the communication result about pure Nash equilibria. In Section 3 we present the results regarding incentives and best-reply dynamics in the CONVERGENCE GAME. In Section 4 we present inapproximability results regarding social-welfare maximization. In Section 5 we discuss open questions and directions for future research.

2. Routing Games. Here we define the game-theoretic model, and begin exploring the two games it contains.

2.1. Two Routing Games. In our model, the network is defined by an undirected graph $G = (N, L)$. N consists of n *source-nodes* $1, \dots, n$ (the players), and a unique *destination-node* d . Each source node i has a routing policy that consists of two components:

1. a *valuation function* v_i that assigns a non-negative value to every *possible* simple route from i to d (i.e., to every simple route in the complete graph over the nodes of G), including the “empty route” \emptyset .⁴ We make the standard assumption [15] that players have strict preferences: For any node i , and every two routes P, Q from i to d that do not have the same first link, it holds that $v_i(P) \neq v_i(Q)$.
2. an *export policy* that dictates which routes i is willing to make available to each neighbouring node j . Observe that nodes’ export policies can be embedded in nodes’ valuation functions (if i is not willing to export route R to j this can be modeled by making R less preferable than the empty route \emptyset to j).

The model consists of two routing games:

2.1.1. A benchmark - the One-Round Game:. The ONE-ROUND GAME is a full-information game in which a strategy of a node i is a choice of an outgoing edge (i ’s choice of an AS to forward traffic to). The *payoff* of node i for a strategy profile is $v_i(R)$ if the strategies induce a route R from i to d , and 0 otherwise.

2.1.2. The Convergence Game:. The CONVERGENCE GAME is a multi-round game with an infinite number of rounds. In each round one or more players (nodes) are chosen to participate by a *Scheduler*. The Scheduler models the asynchronous nature of the Internet, and decides which players participate in each round of the game. The *schedule* chosen must allow every player to play in an infinite number of rounds (the Scheduler cannot deny a node from playing indefinitely). In each round of the game, a player i chosen to play can perform the following actions:

- Read update messages that have arrived from neighbouring nodes. Each update message announces a simple route from the sending neighbouring node to the destination.
- Choose a *single* outgoing edge $(i, j) \in L$ (representing a choice of a neighbouring node to forward traffic to), or \emptyset (not to forward traffic at all).
- Announce simple routes (from i to d) to i ’s neighbouring nodes.

The Scheduler decides in which round sent route announcements reach their destinations. It can arbitrarily delay update messages, but cannot indefinitely prevent update messages of a node from reaching its neighbour (see [15] for a formal model).

A *strategy* of a player in the CONVERGENCE GAME specifies his actions in every round in which that player is chosen to participate. *Best-reply dynamics* is the strategy-profile in which every player continuously performs the following actions: Receive the most recent route announcements from *all* neighbours. Choose the neighbour with the *most* preferred simple route to d (according to your v_i). Announce this route to *all* neighbours.

If from some round onwards i ’s assigned route is constant then i ’s *payoff* is its value for that route. Otherwise, i ’s payoff is 0. More formally, the payoff of player i is $v_i(R)$ if R is a simple route from i to d and from some round onwards, for every link (r, s) on R , r *always* chooses s . Otherwise, i ’s payoff is 0. (All of our results hold

⁴The reason for this is that nodes are not assumed to be familiar with the topology of the network and we wish to capture nodes’ ability to express preferences over *all* routes announced by other nodes.

even if when i 's route "oscillates" i 's payoff is set to be i 's value for the highest valued route it is assigned an infinite number of times.).

2.2. Stable Solutions and Pure Nash Equilibria. Pure Nash equilibria in the ONE-ROUND GAME are known in networking literature as stable solutions. It is not hard to verify that each such stable solution forms a tree rooted in d .⁵ An important requirement from BGP is that it always converge to such a stable solution. However, this is not guaranteed in general, and definitely will not happen if a pure Nash does not exist. Griffin and Wilfong have shown that determining whether a pure Nash equilibrium in the ONE-ROUND GAME exists is NP-hard [16]. We strengthen this result by extending it to the communication model. The use of communication complexity for analyzing uncoupled Nash equilibrium procedures was recently presented in [18], following [2]. Our result can be seen as continuing this line of research.

THEOREM 2.1. *Determining whether a pure Nash equilibrium in the ONE-ROUND GAME exists requires exponential communication (in n).*

Proof. We shall prove a reduction from the communication SET DISJOINTNESS problem [22]. In this problem, there are n communication parties. Each party i holds a subset A_i of $\{1, \dots, K\}$. The parties communicate by broadcasting bits (that are observable to all other parties) and wish to jointly obtain the goal of distinguishing between the two following extreme subcases:

- $\bigcap_i A_i \neq \emptyset$ (i.e., all parties have a mutual element)
- For every $i \neq j$ $A_i \cap A_j = \emptyset$ (i.e., no two parties share an element)

It is known [22] that in order to distinguish between these two subcases the parties must exchange $\Omega(K)$ bits (if $K \gg n$). We set $K = 2^{\frac{n}{2}}$. The reduction to the problem of determining whether a pure Nash in the ONE-ROUND GAME exists is as follows: Consider a network with $2n$ source nodes and a unique destination node d . The set of nodes N consists of 2 disjoint subsets: n sending nodes, and n transit nodes. Each party $i \in [n]$ in the SET DISJOINTNESS problem is associated with a sending node s_i .

The transit nodes are divided into $\frac{n}{2}$ pairs $T_1, \dots, T_{\frac{n}{2}}$. Each such pair of nodes T_r contains a specific node we shall call a 0-node and another node we shall call a 1-node. For every $r = 1, \dots, \frac{n}{2} - 1$, each 0-node in T_r is connected to the 1-node in T_r , and to the 0-node in T_{r+1} . Similarly, each 1-node in T_r is connected to the 0-node in T_r , and to the 1-node in T_{r+1} . Both nodes in $T_{\frac{n}{2}}$ are connected to each other and directly to d . We divide the sending nodes into groups of size 3, $S_1, \dots, S_{\frac{n}{3}}$. For every S_r , the 3 nodes in S_r are connected to each other, to the 0-node in T_1 , and to d . See a description of the network in Figure 2.1. We shall refer to the 0-node in T_1 (which is the access point of all source-nodes to the transit nodes) as c .

We must now define the valuation functions of the different nodes. Let us start with the transit nodes: A 0-node in T_r u has a value of $\frac{1}{2}$ for any route to d in which the next-hop node after u is the 1-node in T_r , and a value of $\frac{1}{4}$ for routes in which the next-hop node after u is the 0-node in T_{r+1} (and very low values for all other routes to d , without ties). Similarly, a 1-node in T_r u has a value of $\frac{1}{2}$ for any route to d in which the next-hop node after u is the 0-node in T_r , and a value of $\frac{1}{4}$ for routes in which the next-hop node after u is the 1-node in T_{r+1} (and very low values for all other routes to d , without ties). Both nodes in $T_{\frac{n}{2}}$ prefer going through each other (a value of $\frac{1}{2}$) than directly to d (a value of $\frac{1}{4}$).

⁵We assume that the graph is connected.

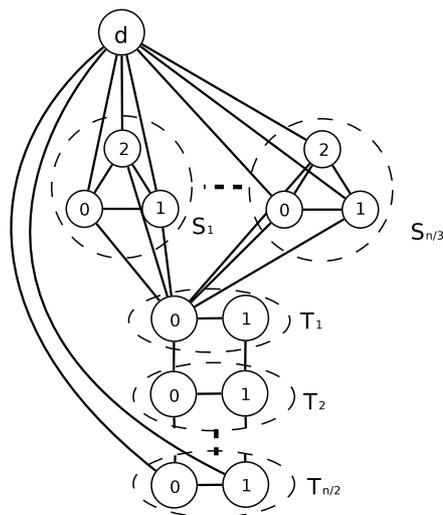


FIG. 2.1. Network constructions used in the proof of Theorem 2.1

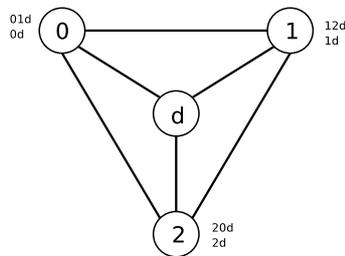


FIG. 2.2. Bad Gadget

Fix a specific triplet of nodes S_r and let $0, 1, 2$ be those nodes. Each node $i = 0, 1, 2$ will assign a value of $\frac{1}{2}$ to the route $(i, i + 1 \pmod{3}, d)$, a value of $\frac{1}{4}$ to the direct route to d , and very low values (without ties) to all other simple routes to d that only include nodes in S_r . This construction of S_r is known as BAD GADGET [16], and appears in Figure 2.2. BAD GADGET is an example of a small network in which no pure Nash equilibrium exists.

There are $2^{\frac{n}{2}}$ possible routes that go through the transit nodes and correspond to strings in $\{0, 1\}^{\frac{n}{2}}$: Consider a string of bits $b = (b_1, \dots, b_{\frac{n}{2}}) \in \{0, 1\}^{\frac{n}{2}}$. b corresponds to the route in which the b_i -node in T_i forwards traffic to the other node in T_i . Fix an arbitrary order over these routes $R_1, \dots, R_{2^{\frac{n}{2}}}$. Now, let each source node s_i assign a value of 1 (no ties) to the route $((s_i, c)R_a$ iff $a \in A_i$. s_i assigns 0 to all unmentioned routes (observe that since all these routes go through the same next-hop node c breaking ties is not a problem).

The key point is that once the route of c is fixed, all source-nodes have no other choice of route through the transit nodes. Hence, either all source nodes agree on a route that goes through the transit nodes, or some choose not to route through the transit nodes at all. The reader can verify that if there is some $a \in \bigcap_i A_i$ then assigning every node i the route $(s_i, c)R_a$ is a pure Nash equilibrium. If, on the other hand, for every $i \neq j$, $A_i \cap A_j = \emptyset$, then there is no pure Nash (because of the BAD

GADGET construction for every triple S_r).

Hence, we have a network with $O(n)$ nodes, in which determining whether a pure Nash equilibrium exists is equivalent to solving the SET DISJOINTNESS problem with n players (each player i corresponds to node s_i), each holding a subset of $1, \dots, 2^{\frac{n}{2}}$. It therefore requires at least $\Omega(2^{\frac{n}{2}})$ bits of communication. This concludes the proof of the theorem. \square

We note that the construction we use in our proof is unlikely to appear in real-life networks; in particular, we require nodes to have unreasonable preferences, according to which very long routes (linear in n , the number of all ASes in the entire Internet) are sometimes very desirable. However, the machinery we use in the proof is sufficient to show the following:

THEOREM 2.2. *Let the valuation functions of the nodes be such that nodes always strictly prefer routes of length at most x to routes of length greater than x . There exists a network and preference setting in which determining whether a pure Nash equilibrium in the ONE-ROUND GAME exists requires $\Omega(2^x)$ bits.*

The theorem is proven by using the same construction as in the proof of Theorem 2.1, only with a transit node hierarchy of depth x instead of $\frac{n}{2}$. This result has interesting implications for BGP convergence, as it shows a lower bound on the worst-case amount of communication that BGP (or any other protocol) requires in order to reach a stable solution, as a function of the nodes' valuations.

2.3. BGP Convergence and Best-Reply Dynamics.

2.3.1. No Dispute Wheel. No Dispute Wheel is the broadest condition known, to date, to guarantee BGP convergence to a stable solution. In our terms, this translates to convergence of best-reply dynamics in the CONVERGENCE GAME. A Dispute Wheel, defined by Griffin et al. [15], is an abstract mathematical structure that can be induced by the network topology and the valuation functions. Formally, a dispute wheel is defined as the 3-tuple $(\mathcal{U}, \mathcal{R}, \mathcal{Q})$ where $\mathcal{U} = (u_0, u_1, \dots, u_{k-1})$ is a sequence of k nodes in the network and $\mathcal{R} = (R_0, R_1, \dots, R_{k-1})$, $\mathcal{Q} = (Q_0, Q_1, \dots, Q_{k-1})$ are sequences of routes that exist in G (indices for these nodes and routes should be considered modulo k). We shall call u_0, \dots, u_{k-1} the *pivot-nodes*. It must hold that:

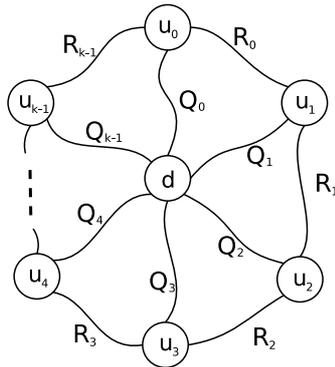
- Each route Q_i starts at u_i and ends at the destination node d .
- Each route R_i starts at node u_i and ends at node u_{i+1} .
- $v_i(Q_i) \leq v_i(R_i Q_{i+1})$ (where $R_i Q_{i+1}$ is the concatenation of the routes R_i and Q_{i+1})

The term ‘‘Dispute Wheel’’ is due to the resemblance of its form to that of a wheel. Figure 2.3 depicts such a structure. It is known that No Dispute Wheel guarantees that there is a unique stable solution [15].

No Dispute Wheel is known to hold for several interesting special cases. One special case is *Metric-Based Routing* (see, e.g., [8]) which is a generalization of Shortest-Path Routing (in which the length of every link is 1).

In practice, there is no objective metric according to which all route choices are made. ASes' preferences are influenced by many economic, and other, considerations, like the business relationships between them. The Gao-Rexford setting is said to accurately depict the underlying commercial structure of today's Internet [21], and is a special case of No Dispute Wheel [11].

2.3.2. The Gao-Rexford Framework. Studies of the commercial Internet [21] suggest two types of business relationships that characterize AS inter-connections:



Each node u_i would rather route clockwise through node u_{i+1} than through the path Q_i

FIG. 2.3. A Dispute Wheel

Pairs of neighbouring ASes have either a *customer-provider* or a *peering* relationship. Customer ASes pay their provider ASes for connectivity – access to Internet destinations through the provider’s links and advertisement of customer destinations to the rest of the Internet. Peers are ASes that find it mutually advantageous to exchange traffic for free among their respective customers, *e.g.*, to shortcut routes through providers. An AS can be in many different relationships simultaneously: It can be a customer of one or more ASes, a provider to others, and a peer to yet other ASes. These agreements are assumed to be longer-term contracts that are formed because of various factors, *e.g.*, the traffic pattern between two nodes.

In a seminal paper Gao and Rexford [12] suggest constraints on routing policies that are naturally induced by the business relationships between ASes.

(i) **No customer-provider cycles:** Let G_{CP} be the directed graph with the same set of nodes as G and with a directed edge from every customer to its direct provider. There must be no directed cycles in this graph. This requirement has a natural economic justification as it means that no AS is indirectly its own provider.

(ii) **Prefer customers to peers and providers:** A *customer route* is a route in which the *next-hop* AS (the first AS to which packets are forwarded on that route) is a customer. *Provider* and *peer routes* are defined similarly. We require that nodes always prefer customer routes over peer routes and provider routes. This constraint is on the valuation functions of the nodes – it demands every node assign customer routes higher values than peer routes and provider routes.

(iii) **Provide transit services only to customers:** Nodes do not always carry *transit traffic*—traffic that originates and terminates at hosts outside the node. ASes are obligated (by financial agreements) to carry transit traffic to and from their customers. However, ASes do not carry transit traffic between their providers and peers. Therefore, ASes should share only customer routes with their providers and peers but should share *all* of their routes with their customers. This constraint is on the filtering policy of the nodes – it requires that nodes only export peer and provider routes to their customers (customer routes are exported to all neighbouring nodes).

3. Best-Reply Dynamics and Incentives. In this section we discuss several results regarding the incentive-compatibility of BGP and best-reply dynamics in the CONVERGENCE GAME.

3.1. BGP Is Not Incentive-Compatible. We first prove the following theorem:

THEOREM 3.1. *BGP is not incentive-compatible in ex-post Nash even if the No Dispute Wheel condition holds.*

Proof. Consider the network in Figure 3.1. There are 3 source nodes, 1, 2, m and a destination node d . Each nodes' two most preferred routes are listed next to it (where the higher route is more preferred than the lower route). It is easy to show that the valuation functions and topology of the network do not induce a Dispute Wheel (because route $2md$ does not actually exist). Hence, there is a unique pure Nash in the ONE-ROUND GAME, and best-reply dynamics always converge to this Nash in the CONVERGENCE GAME. The unique pure Nash is assigning $12d, 2d, m12d$ to 1, 2, m , respectively. However, if m announces to 2 repeatedly that its route is md then it is easy to see that BGP dynamics will always assign 1 the direct route to d , $1d$, thus enabling m to get its most preferred route $m1d$. Hence, deviating from BGP is beneficial.

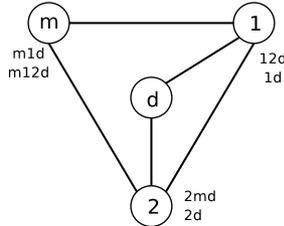


FIG. 3.1. *Best-Reply Dynamics Are Not Incentive-Compatible*

We note that it is possible to define business relationships between the nodes in this example that are consistent with the Gao-Rexford constraints; for instance, assume that m is a customer of 1, 2, who are both customers of d , and that 1 is a customer of 2. Observe that this is consistent with the first two Gao-Rexford constraints – it does not create customer-provider cycles, and the preferences of all nodes are such that customer routes are always preferred to provider routes (there are no peer routes). Even if we assume that 1, 2 (and d) execute BGP and uphold the third Gao-Rexford constraint (they only announce peer- and provider-routes to their customers), m can still gain by lying to 2, as explained above. \square

3.2. BGP With Route Verification Is Incentive-Compatible. As we have said before, the fact that BGP, as is, is not incentive compatible, can be rectified if Route Verification holds.

THEOREM 3.2. *If No Dispute Wheel and Route Verification hold, then best-reply dynamics are incentive-compatible in (subgame perfect) ex-post Nash*

Proof. Consider a network graph $G = (N, L)$ for which No Dispute Wheel holds. There is a unique stable solution T to which all best-reply dynamics are bound to converge in the CONVERGENCE GAME. We denote the route of every source node r in T by T_r .

Assume, by contradiction, that some manipulating node r_m manages to reach a different outcome M by unilaterally deviating from best-reply dynamics (not execut-

ing BGP), and gains by doing so. We shall show that this implies the existence of a Dispute Wheel. The proof shall proceed in steps, pointing out a sequence of routes in the graph that will eventually form a Dispute Wheel.

We define the route M_r to be the route node r believes it is assigned in M (i.e., the route that r 's next selected hop has advertised to r). That is, it could be that the manipulator tricked nodes that send traffic through it in M to believe that their traffic is forwarded along a route not used in practice.

REMARK 3.3. *We note that it could be the case that node r_m intentionally causes a protocol divergence that does not affect it in order to improve its routing outcome (that is, remote persistent route oscillations). Our proof also works for this case simply by defining the route M_r of a node r whose route oscillates to be r 's most preferred route, out of the routes assigned to r in the oscillation. Observe that, in this case, there is no well-defined outcome M because nodes' routes (the M_r 's) can be inconsistent. Luckily, this is not needed for our proof to work. For ease of exposition, we prove our result below for the case that there are no persistent route oscillations following the manipulation. The reader can verify that all of our arguments also hold also for the above definition of the M_r 's.*

Since we assumed that r_m gained from its manipulation we deduce that:

$$(3.1) \quad v_{r_m}(T_{r_m}) < v_{r_m}(M_{r_m})$$

Because r_m strictly prefers M_{r_m} to T_{r_m} , but did not choose it in the routing tree T , we must conclude that the route M_{r_m} is not available to r_m in T . This means that there must exist some node r (other than r_m) that is on the route M_{r_m} and that does not have the same route in M as it has in T . Let r_1 be the node on the path M_{r_m} that is closest to d on M_{r_m} , such that $M_{r_1} \neq T_{r_1}$.

By definition, all nodes that follow r_1 on the route M_{r_m} have exactly the same routes in T and in M . This means that the node r_1 could choose the route M_{r_1} in T . Since it did not choose that route we must conclude⁶ that:

$$(3.2) \quad v_{r_1}(M_{r_1}) < v_{r_1}(T_{r_1})$$

We can now proceed to the next step in the proof. Since T_{r_1} is preferred by r_1 , and was not chosen by r_1 in the routing tree M , it must be that it was not an available option. Therefore, there is some node r on the route T_{r_1} , that is not r_1 , such that $T_r \neq M_r$. We select r_2 to be the node r closest to d on the path T_{r_1} for which $T_r \neq M_r$. As before, all nodes closer to d than r_2 on the route T_{r_1} send traffic along identical routes in both T and M . Hence, the route T_{r_2} must be available to r_2 even in M . The fact that it was not chosen in M implies that r_2 prefers M_{r_2} over it. Thus, we have that:

$$(3.3) \quad v_{r_2}(T_{r_2}) < v_{r_2}(M_{r_2})$$

We can continue these steps, alternating between the routing trees T and M and creating a sequence of nodes as follows:

- $r_0 = r_m$

⁶The reason that the inequality is strict is that equality can exist only if the two routes go through the same neighbouring node. This cannot be the case as $M_{r_1} \neq T_{r_1}$.

for $n = 0, 1, 2, \dots$ we perform the following steps:

- **M step:** Let r_{2n+1} be the node r on the route $M_{r_{2n}}$ such that $M_r \neq T_r$, and r is closest among all such nodes to d on $M_{r_{2n}}$.
- **T step:** Let r_{2n+2} be the node r on the route $T_{r_{2n+1}}$ such that $M_r \neq T_r$, and r is closest among all such nodes to d on $T_{r_{2n+1}}$.

Note that the destination node d cannot appear in this sequence because the route $M_d = T_d$ is the empty set. Due to our construction, and to arguments similar to the ones presented before, the preferences over routes are as follows:

$$(3.4) \quad \text{for } i = 0, 2, 4, \dots \quad v_{r_i}(T_{r_i}) < v_{r_i}(M_{r_i})$$

$$(3.5) \quad \text{for } i = 1, 3, 5, \dots \quad v_{r_i}(M_{r_i}) < v_{r_i}(T_{r_i})$$

Since there is only a finite number of nodes, at some point a node will appear in this sequence for the second time. We denote the first node that appears two times in the sequence by u_0 . Let u_0, \dots, u_{k-1}, u_0 be the substring of r_0, r_1, \dots that begins in the first appearance of u_0 and ends in its second appearance. We shall examine two distinct cases.

CASE I: The manipulator r_m does not appear in the substring u_0, \dots, u_{k-1}, u_0 .

PROPOSITION 3.4. *If for all $i \in \{0, \dots, k-1\}$ $r_m \neq u_i$ (the manipulator is not one of the nodes in the substring) then k must be even.*

Proof. If k is odd, then it must be that both u_1 (in its first appearance in the substring) and u_0 (in its second appearance in the substring) were both selected in M steps, or were both selected in T steps. However, if this is the case we reach a contradiction as both nodes were supposed to be the node r closest to d on a certain route, such that $T_r \neq M_r$. Since $u_{k-1} \neq u_0$ this cannot be. \square

If k is even then the substring of nodes u_0, \dots, u_{k-1}, u_0 , along with the T_{u_i} and M_{u_i} route, and the preferences over these routes (expressed before) form a dispute wheel (as in Figure 3.2).

CASE II: The manipulator r_m appears in the substring (that is, $u_0 = r_m$). We now need to handle two subcases: The subcase in which k is even and the subcase in which k is odd. If k is even then the second appearance of the manipulator (u_0) in the substring is due to a T step. If so, a Dispute Wheel is formed, as in the example in Figure 3.2⁷.

We are left with the subcase in which k is odd. In this case the second appearance of r_m was chosen in an M step. If so, it must be that $M_{u_{k-1}}$ (that goes through r_m) is not used in practice (otherwise, both u_{k-1} and that the second appearance of $r_m = u_0$ was chosen in M steps, and arguments similar to those of Proposition 3.4 would result in a contradiction). This must be the result of a manipulation by r_m . Let L_{r_m} be the false route reported by the manipulator to the node that comes before it on $M_{u_{k-1}}$. Recall that the manipulator can only announce a route L_{r_m} that exists and is available to it in M ⁸. Recall, that the second appearance of the manipulator was chosen due to an M step. Therefore, all nodes that follow it on $M_{u_{k-1}}$ (which are the same nodes as in L_{r_m}) are assigned the same routes in T and M . Hence, L_{r_m} was available to r_m in T . It must be that $v_{r_m}(L_{r_m}) \leq v_{r_m}(T_{r_m})$, for otherwise r_m would

⁷The route $R \setminus S$ (where S is a sub-route of R) is route R truncated before the beginning of S

⁸Notice that this is the crucial step in which route verification is used within the proof.

have chosen L_{r_m} as its route in T (a contradiction to the stability of T). We know that $v_{r_m}(T_{r_m}) \leq v_{r_m}(M_{r_m})$ because we assumed that the manipulation performed by r_m was beneficial to it. We get:

$$(3.6) \quad v_{r_m}(L_{r_m}) \leq v_{r_m}(T_{r_m}) \leq v_{r_m}(M_{r_m})$$

Thus, we form a Dispute Wheel with L_{r_m} as shown in Figure 3.3.

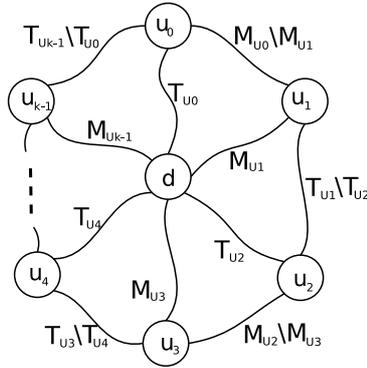


FIG. 3.2. A Dispute Wheel constructed during the proof of Theorem 3.2 (even number of nodes).

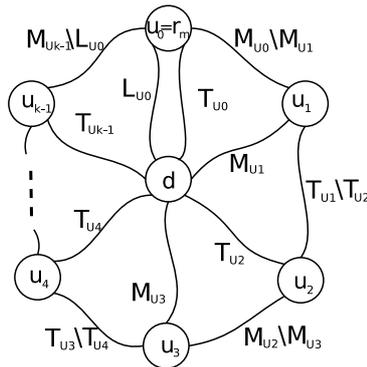


FIG. 3.3. A Dispute Wheel constructed during the proof of Theorem 3.2 (odd number of nodes).

□

Notice that route verification is required only for CASE II of the proof above, specifically when there is an odd number of nodes in the cycle.

3.3. BGP With Route Verification Is Collision-Proof. In our context, collusion proofness in ex-post Nash [10] means that even a coalition of ASes of any size cannot *strictly*⁹ better the routing outcomes (*i.e.*, their actual routes) of *all* members in the coalition by deviating from BGP. A construction similar to that in the proof of 3.2 shows that Theorem 3.2 can be strengthened as follows:

THEOREM 3.5. *If No Dispute Wheel and Route Verification hold, then the best-reply dynamics are collusion-proof in (subgame perfect) ex-post Nash.*

⁹The definition of collusion-proofness in [10] also allows weak inequalities. Here, we shall only deal with cases in which all members of the coalition are strictly better off.

Proof. We shall assume, by contradiction, that a group of manipulators colludes in an interdomain routing instance with no Dispute Wheel in order to improve their routing outcomes. We define T_r and M_r as in the proof of Theorem 3.2. We assume, by contradiction, that all manipulators are not harmed by this manipulation:

$$(3.7) \quad \forall r \in \text{Manipulators} \quad v_r(T_r) < v_r(M_r)$$

We shall arrive at a contradiction by showing the existence of a Dispute Wheel in a similar manner to that demonstrated in the proof of Theorem 3.2.

We begin the construction by selecting one of the manipulators that strictly gained from the collusion. We shall denote this manipulator by r_m (it must be that $T_{r_m} \neq M_{r_m}$). We then construct a sequence of nodes in the following way (This is a slightly modified version of the construction appearing in Theorem 3.2):

- $r_0 = r_m$
- **M-manip step:** For a node r_n which is a manipulator we define r_{n+1} to be the node r on the route M_{r_n} , such that $M_r \neq T_r$, and r is the closest to d on M_{r_n} among all such nodes.
- **T step:** For a node r_n that is not a manipulator, and was chosen in an M step, we define r_{n+1} to be the node r on the route T_{r_n} , such that $M_r \neq T_r$, and r is the closest to d on T_{r_n} among all such nodes.
- **M-honest step:** For a node r_n that is not a manipulator, and was chosen in a T step, we define r_{n+1} to be the node r on M_{r_n} , such that $M_r \neq T_r$, and r is the closest to d on M_{r_n} among all such nodes.

The nodes in the sequence r_0, r_1, r_2, \dots must start repeating at some point. We denote by u_0 the first node that repeats in the sequence, and by u_0, \dots, u_{k-1}, u_0 the sub-sequence of nodes between those repetitions. As in Theorem 3.2, we will reach a contradiction by showing that a dispute wheel exists in the graph. The nodes u_0, \dots, u_{k-1}, u_0 will be the pivot-nodes. First, we prove the following proposition that will aid our construction.

PROPOSITION 3.6. *There is no $i \in \{0, \dots, k-1\}$ such that both u_i and u_{i+1} (modulo k) are manipulators (no two manipulators come one after the other in the substring u_0, \dots, u_{k-1}, u_0).*

Proof. By contradiction, let u_i and u_{i+1} be two consecutive manipulators. u_{i+1} was chosen in an M step. u_{i+1} is therefore the node r closest to d on M_{u_i} such that $M_r \neq T_r$. Hence, $M_{u_{i+1}}$ must be available to u_{i+1} in both M and T . We know that $v_{u_{i+1}}(T_{u_{i+1}}) \leq v_{u_{i+1}}(M_{u_{i+1}})$, as u_{i+1} is a manipulator. Since u_{i+1} chose $T_{u_{i+1}}$ over $M_{u_{i+1}}$ in T it must also be that $v_{u_{i+1}}(T_{u_{i+1}}) \geq v_{u_{i+1}}(M_{u_{i+1}})$. We conclude that $v_{u_{i+1}}(T_{u_{i+1}}) = v_{u_{i+1}}(M_{u_{i+1}})$. However, equality of the values of routes assigned by u_{i+1} is only possible if u_{i+1} forwards traffic to the same node in both routes. Since both routes are available in T , this means that $T_{u_{i+1}} = M_{u_{i+1}}$. This contradicts the reason for which u_{i+1} was selected ($M_{u_{i+1}} \neq T_{u_{i+1}}$). \square

We will now need to demonstrate that each node u_i in the repeating sequence has two paths to its destination: a direct path Q_i , and one that passes through u_{i+1} that we will denote as $R_i Q_{i+1}$. We will show that each node prefers the indirect path over the direct one. We will examine three possible cases that match the three types of steps in the selection of the series of nodes.

(i) **M-manip step:** For a node u_i which is a manipulator we select the route $R_i Q_{i+1}$ as the indirect route M_{u_i} that passes through u_{i+1} . This path must exist since it is the one the manipulator uses in the manipulation (and gains by from its use).

From proposition 3.6, we know that the node u_{i-1} is not a manipulator, since during the construction we picked node u_i after node u_{i-1} . Node u_{i-1} may have been picked during an M-honest step or a T-honest step (as will be shown below). Because honest nodes use Route Verification, it must be that the path from u_{i-1} to the destination through node u_i exists. We set the path Q_i to be the suffix of the path node u_{i-1} believes it is using. Node u_i was selected as the last node on this path that routes differently in M and T , and so this path must be available to node u_i in T . From this we learn that

$$(3.8) \quad v_{u_i}(Q_i) \leq v_{u_i}(T_{u_i}) < v_{u_i}(M_{u_i}) = v_{u_i}(R_i Q_{i+1})$$

which is the preference relation we require to construct a dispute wheel.

(ii) **T step:** For a node u_i that is not a manipulator, and was chosen in an M step, we define Q_i to be the path that this node has in M . In this case we know that Q_i exists and is not just a perceived route, as the node u_i is honest, and all nodes below it route the same in M and T . We define its indirect route $R_i Q_{i+1}$ as its path in T which must exist in the network, as it is used when all nodes play honestly, and which by construction passes through u_{i+1} . Note, that node u_i is the closest node to d on $M_{u_{i-1}}$, and all following nodes on this path route the same in M, T . However, this path is not chosen by u_i in T which means it is less preferred:

$$(3.9) \quad v_{u_i}(Q_i) < v_{u_i}(R_i Q_{i+1})$$

(iii) **M-honest step:** For a node u_i that is not a manipulator, and was chosen in a T step, we define Q_i to be the path that this node has in T . We define its indirect route $R_i Q_{i+1}$ as the path it believes it has in M (note that here we are consistent with the definition of Q_{i+1} for the case in which node u_{i+1} is a manipulator). This path must exist in the network due to Route Verification and by construction passes through node u_{i+1} . Note, that node u_i is the closest node to d on $T_{u_{i-1}}$, and all following nodes on this path route the same in M, T . However, this path is not chosen by u_i in M which means it is less preferred:

$$(3.10) \quad v_{u_i}(Q_i) < v_{u_i}(R_i Q_{i+1})$$

The reader may check that the direct paths defined as Q_{i+1} in each step is consistent with the suffix of $R_i Q_{i+1}$ that is defined in the previous step. We have therefore completed our construction of a dispute wheel from the topology of the graph, and preferences of nodes which proves the theorem. \square

Again, notice that the crucial application of route verification in the proof is during an M-honest step where an honest node is assigned a path $R_i Q_{i+1}$ which is the path it believes it is routing through in M . Even if a manipulator is present on this path, route verification allows us to conclude that every link on this route actually exists (even if it is not really used for routing), and so can be a part of the dispute wheel we construct.

3.4. Discussion: On Dispute Wheels and Filtering Policies. The results presented in this section imply that when No Dispute Wheel holds, BGP with Route

Verification has surprisingly good properties from a mechanism design perspective. This means that No Dispute Wheel is not only the most general known condition, to date, that guarantees BGP convergence to a stable solution, but also ensures that this convergence is properly incentivized (in the sense that the individual nodes have no motivation not to execute BGP).

What if the network and the nodes' preferences do induce a Dispute Wheel? Observe that the BGP strategy profile considered in this paper is such that nodes announce routes they use to *all* neighbours. However, a natural way to avoid Dispute Wheels is to make sure that certain routes (that could potentially appear in a Dispute Wheel) are removed from consideration by ensuring that they are not announced to certain nodes. In fact, this is precisely what is achieved by the third Gao-Rexford constraint, that requires that ASes only announce certain routes to their customers (and not to their providers or peers).

To enable us to discuss such scenarios, in which nodes are requested to “filter” certain routes, we can now think of BGP (best-reply dynamics) as a strategy-profile in which every player continuously performs the following actions: Receive the most recent route announcements from *all* neighbours. Choose the neighbour with the *most* preferred simple route to d (according to your v_i). Announce this route to a *subset your neighbours*, given your prescribed *filtering policy*. For instance, in the Gao-Rexford setting, we can think of BGP as the strategy-profile in which ASes have filtering policies that uphold the third Gao-Rexford constraint.

All of the results presented in this section apply to settings like the Gao-Rexford setting, in which nodes filter routes based on prescribed filtering policies, as long as the two following conditions hold:

1. The prescribed filtering policies imply No Dispute Wheel. That is, every Dispute Wheel that is induced by the network and the nodes' preferences is such that some route in that Dispute Wheel should be filtered by one of the nodes on it. The Gao-Rexford setting is one example in which the prescribed filtering policies indeed prevent Dispute Wheel formation [12].

2. Nodes cannot announce routes that violate their prescribed filtering policies without being caught. That is, if a node j announces a route to node i that it should not announce according to its prescribed filtering policy, node i should be able to detect this “illegitimate” announcement and ignore it. In the Gao-Rexford setting, for instance, this can be accomplished via very local security checks, because in order to figure out whether a route R_j announced by j is “legitimate”, i need only verify the business relationship between j and the next-hop node on R_j .¹⁰¹¹

4. Maximizing Social Welfare. We prove that obtaining an approximation ratio better than n to the optimal social welfare is hard even if No Dispute Wheel holds. In fact, this can be shown even for the Gao-Rexford setting. We present two lower bounds, one in the computational complexity model, and one in the communication complexity model.

THEOREM 4.1. *Obtaining an approximation of $O(n^{\frac{1}{2}-\epsilon})$ to the social welfare is impossible unless $P = NP$. Obtaining an approximation of $O(n^{1-\epsilon})$ to the social*

¹⁰Specifically, if j is i 's customer or peer then i should verify (e.g., via cryptographic signatures) that k is j 's customer.

¹¹In reality, business relations between the ASes are sometimes secret and may be difficult to verify. However, some information is already revealed by the fact that an AS advertises a route through its neighbor. All that remains is to verify this implicit claim.

welfare is impossible unless $P = ZPP$. This holds for any $\epsilon > 0$ and even in the Gao-Rexford setting.

Proof. Our proof will be by reduction from CLIQUE. Assume a graph $G = \langle V, E \rangle$, we construct a network with N nodes and L links. In this network, N consists of $2|V| + 1$ source-nodes and a unique destination node d . The source nodes are divided into 3 disjoint sets: Two sets N_1, N_2 , such that $|N_1| = |N_2| = |V|$ and a *connection node* c . We create a node $v_{N_1} \in N_1$ and a node $v_{N_2} \in N_2$ in the AS graph for every node $v \in V$ from the given instance of the CLIQUE problem. All nodes in N_1 are connected to the connection node c . All nodes in N_2 are connected to each other, to the connection node, and to d . See Figure 4.1.

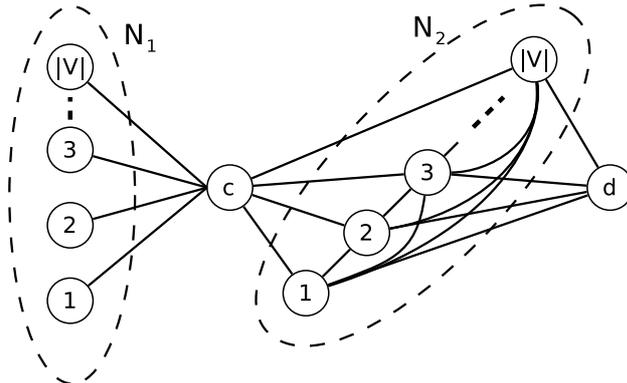


FIG. 4.1. The network in the proof of Theorem 4.1

All nodes in N_2 , and c , have valuation functions that assigns a value close to 0 to all routes (no ties). Fix some order O on the nodes in N_2 (for now, this order shall play no part in the proof, we shall only need it to show how to make the proof work in the Gao-Rexford framework). A node $v_{N_1} \in N_1$ assigns a value close to 1 (no ties) to a route R iff *all* the following conditions hold:

- (v_{N_1}, c) is the first link on R .
- The order of appearance of the nodes from N_2 in R is consistent with O .
- v_{N_2} is on R .
- For every node $u_{N_2} \neq v_{N_2} \in N_2$ on R there is an edge (v, u) in G .

Observe that since c is the only connection between N_1 and $N_2 \cup \{d\}$, c 's route determines the routes of all nodes in N_1 . The reader can verify that every clique in G corresponds to a routing tree with a social welfare that equals (almost) the size of the clique (assign every node in N_1 that is in the clique the route that goes through c and then all the nodes in N_2 that are associated with nodes in the clique). In addition, the social welfare of every routing tree corresponds to a clique in the original graph G (the route from c to d through N_2 determines the identity of the clique). The theorem follows from the known inapproximability results for CLIQUE [19].

We note that this result can be made to hold in the Gao-Rexford setting by defining business relationships as follows: c is a customer of all nodes in N_1 . All nodes in N_2 are customers of c . For every two nodes $r, s \in N_2$, such that s comes after r in O , s is r 's customer. d is a customer of all nodes in N_2 . \square

THEOREM 4.2. *Obtaining an approximation of $O(n^{1-\epsilon})$ to the social welfare requires exponential communication (in n). This holds for any $\epsilon > 0$ and even in the Gao-Rexford setting.*

Proof. The proof is similar to the proof of Theorem 2.1, and is by reduction from SET DISJOINTNESS. There are n parties, and each party i holds a subset A_i of $1, \dots, K$. The goal is to distinguish between the two following extreme subcases:

- $\bigcap_i A_i \neq \emptyset$
- For every $i \neq j$ $A_i \cap A_j = \emptyset$

It is known that in order to distinguish between these two subcases the parties must exchange $\Omega(K)$ bits. We set $K = 2^{\frac{n}{2}}$.

Now, consider a network with $2n + 1$ source nodes and a unique destination node d . The set of nodes N consists of 3 disjoint subsets: n sending nodes, a connecting node c , and n transit nodes. Each party $i \in [n]$ in the SET DISJOINTNESS problem is associated with a sending node s_i .

The transit nodes are divided into $\frac{n}{2}$ pairs $T_1, \dots, T_{\frac{n}{2}}$. All sending nodes are connected to the connecting node, which, in turn, is connected to both nodes in T_1 . For every $r = 1, \dots, \frac{n}{2} - 1$, each node in T_r is connected to both nodes in T_{r+1} . Both nodes in $T_{\frac{n}{2}}$ are connected directly to d . See a description of the network in Figure 4.2.

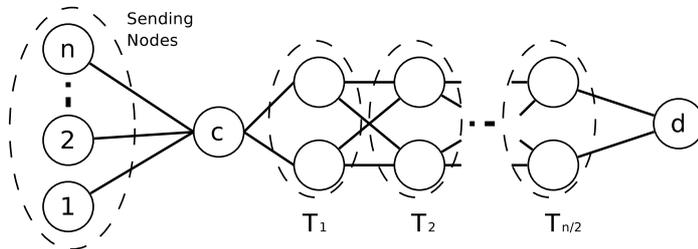


FIG. 4.2. The network in the proof of Theorem 4.2

All transit nodes and the connecting node have a value close to 0 (no ties) for all routes. There are $2^{\frac{n}{2}}$ possible routes from c to d that go through the transit nodes (see proof of Theorem 2.1). Fix an arbitrary order on these routes $R_1, \dots, R_{2^{\frac{n}{2}}}$. The valuation function of each s_i assigns a value close to 1 (no ties) to a route $(s_i, c)R_a$ iff $a \in A_i$. It assigns a value close to 0 to all other routes (no ties). The reader can verify that there is a route assignment with a social welfare-value close to n if there is some $a \in \bigcap_i A_i \neq \emptyset$ (assign every node i the route $(s_i, c)R_a$). If, on the other hand, for every $i \neq j$ $A_i \cap A_j = \emptyset$, then any route assignment cannot have a social-welfare value better than $1 + \epsilon$ (this is because c 's route determines the routes of the sending nodes through the transit nodes).

Hence, we have a network with $O(n)$ nodes, in which determining whether the social-welfare is n or 1 is equivalent to solving the SET DISJOINTNESS problem with n players (each player i simulates node s_i), each holding a subset of $1, \dots, 2^{\frac{n}{2}}$. It therefore requires at least $\Omega(2^{\frac{n}{2}})$ bits of communication. This concludes the proof of the theorem.

This result too can be made to hold for the Gao-Rexford setting if we define business relationships as follows: For every $i = 2, \dots, \frac{n}{2}$, the nodes in T_i are customers of the nodes in T_{i-1} . d is a customer of the nodes in $T_{\frac{n}{2}}$, the nodes in T_1 are customers of c , and c is a customer of all nodes in N_1 . \square

REMARK 4.3. As in the case of Theorem 2.1, the proof of Theorem 4.2 can easily be modified to show that even if we assume that nodes are only interested in routes of length at most x , there exists a network setting in which obtaining an approximation of $O(x^{1-\epsilon})$ requires communicating $O(2^x)$ bits.

A trivial upper bound of n can be achieved by finding the node with the highest value for some route, assigning that route to that node (thus getting an n -approximation)¹², and then assigning routes to all other nodes in a way that forms a tree rooted in d .

5. Open Questions.

(i) No Dispute Wheel is sufficient, but not necessary, for guaranteeing BGP convergence. Recall, that when we say BGP convergence we are referring to BGP convergence for *all* possible timings of update messages, a property known as “*BGP safety*” in networking literature [16]. Very recently (subsequently to our work), Sami et al. [26] presented a more general sufficient condition for BGP convergence than No Dispute Wheel called, “*Iterated Dominance Tree*”. Sami et al. [25] have also shown that the existence of two or more stable solutions always implies that BGP safety does not hold, thus providing the first non-trivial *necessary* condition for BGP convergence (uniqueness of the stable solution). However, there is still a big gap between the known sufficient conditions, and this new necessary condition. **An important open question is to find non-trivial *characterizations* (sufficient and necessary conditions) of networks for which BGP safety is guaranteed.**

(ii) **What more general conditions than No Dispute Wheel and Route Verification guarantee incentive-compatible convergence of best-reply dynamics in our interdomain routing model?** One can try to identify such more general conditions by relaxing the No Dispute Wheel assumption, and/or the Route Verification requirement. Indeed, Sami et al. [26] have been able to strengthen our results, by showing that similar results hold for a condition broader than No Dispute Wheel (called Iterated Dominance Tree), and a security requirement weaker than Route Verification (called Topology Validation). However, we still do not have a good understanding of which conditions imply incentive-compatibility. No Dispute Wheel does not only guarantee BGP convergence, but even *robust BGP convergence*. That is, not only is BGP guaranteed to converge, but it will converge even if nodes and links are arbitrarily removed from the network (because for every subgraph of the AS graph No Dispute Wheel holds). This is very important from a practical perspective as it implies that BGP convergence is achieved even in the presence of link and node failures. We present the following conjecture:

Conjecture: *Robust BGP convergence always implies incentive-compatibility of BGP with Route Verification.*

That is, whenever the network and routing policies are such that robust BGP convergence is guaranteed, BGP with Route Verification is incentive-compatible.

(iii) **Other Models.** The questions we consider in this paper, and our formal framework, can be applied to variations of our model. For instance, subsequently to our work Goldberg et al. [13] have presented several impossibility results for the case that utility functions in our model are also influenced by incoming traffic (i.e., nodes do not only care about the routes they get, but also about which nodes are routing through them). Similarly, Kintali [31] extended our game-theoretic model to the case of Fractional BGP [20]. It will be interesting to study incentive-compatibility

¹²this bound naturally assumes a representation in which finding such a maximal value route is easy

of best-reply dynamics in various other interdomain routing settings.

(iv) **Identify realistic interdomain routing settings in which good approximations to the social-welfare are possible.** It has been shown [8] that No Dispute Wheel and Policy Consistency imply that BGP always converges to the socially-optimal routing tree. It was also shown [8] that relaxations of Policy Consistency lead to arbitrarily bad solutions, even in very small and realistic networks. We are interested in identifying settings in which BGP is guaranteed to reach solutions that have a good social-welfare value. In our game-theoretic model, this question translates to a *price of anarchy* [23] problem: We wish to identify restrictions on the topology of the network, and the routing policies of nodes, for which the worst-case ratio (with respect to social welfare) between pure Nash equilibria in the ONE-ROUND GAME (stable solutions), and the optimal solution, is small.

There is also the computational question: Our results in this paper show that even in the Gao-Rexford setting one cannot find a good approximation to the optimal social welfare (via *any* algorithm) *in polynomial time*. In which settings is it possible to obtain a good approximation in polynomial time (via BGP, or another protocol)? Furthermore, will it be possible to achieve some approximation via a non-centralized computation through a network protocol?

Acknowledgements. We thank Sharon Goldberg and Noam Nisan for significant contributions to this paper. We thank Shahar Dobzinski, Joan Feigenbaum, Felix Fischer, Aaron Jaggar, Ehud Kalai, Ahuva Mu'alem, Motty Perry, Vijay Ramachandran, Jeff Rosenschein, Rahul Sami, and Scott Shenker for helpful discussions. We also thank the anonymous reviewers for valuable comments.

REFERENCES

- [1] Kevin Butler, Toni Farley, Patrick McDaniel, and Jennifer Rexford. A survey of BGP security issues and solutions. Technical Report TD-5UGJ33, AT&T Labs, 2004.
- [2] Vincent Conitzer and Tuomas Sandholm. Communication complexity as a lower bound for learning in games. In *In International Conference on Machine Learning (ICML)*, pages 185–192, 2004.
- [3] Ronny R. Dakdouk, Semih Salihoglu, Hao Wang, Haiyong Xie, and Yang Richard Yang. Interdomain routing as social choice. In *Proceedings the of Second International Workshop on Incentive-Based Computing*, Lisboa, Portugal, 2006.
- [4] Alex Fabrikant and Christos Papadimitriou. The complexity of game dynamics: BGP oscillations, sink equilibria, and beyond. In *Proceedings of ACM-SIAM Symposium on Discrete Algorithms*, 2008.
- [5] Nick Feamster, Ramesh Johari, and Hari Balakrishnan. Implications of autonomy for the expressiveness of policy routing. *SIGCOMM Computer Communication Review*, 35(4):25–36, 2005.
- [6] Joan Feigenbaum, David R. Karger, Vahab S. Mirrokni, and Rahul Sami. Subjective-cost policy routing. *Theoretical Computer Science*, 378(2):175–189, 2007.
- [7] Joan Feigenbaum, Christos H. Papadimitriou, Rahul Sami, and Scott Shenker. A BGP-based mechanism for lowest-cost routing. *Distributed Computing*, 18(1):61–72, 2005.
- [8] Joan Feigenbaum, Vijay Ramachandran, and Michael Schapira. Incentive-compatible interdomain routing. In *Proceedings of the 7th ACM conference on Electronic commerce*, pages 130–139, 2006.
- [9] Joan Feigenbaum, Rahul Sami, and Scott Shenker. Mechanism design for policy routing. *Distributed Computing*, 18(4):293–305, 2006.
- [10] Joan Feigenbaum, Michael Schapira, and Scott Shenker. *Algorithmic Game Theory*, chapter 14, pages 363–383. Cambridge University Press, 2007.
- [11] Lixin Gao, Timothy G. Griffin, and Jennifer Rexford. Inherently safe backup routing with

- BGP. In *Proceedings of the Twentieth Annual Joint Conference of the IEEE Computer and Communications Societies (INFOCOM 2001)*, pages 547–556, 2001.
- [12] Lixin Gao and Jennifer Rexford. Stable Internet routing without global coordination. *IEEE/ACM Transactions on Networking*, 9(6):681–692, 2001.
 - [13] Sharon Goldberg, Shai Halevi, Aaron D. Jaggard, Vijay Ramachandran, and Rebecca N. Wright. Rationality and traffic attraction: incentives for honest path announcements in bgp. In Victor Bahl, David Wetherall, Stefan Savage, and Ion Stoica, editors, *SIGCOMM*, pages 267–278. ACM, 2008.
 - [14] Timothy G. Griffin, Aaron D. Jaggard, and Vijay Ramachandran. Design principles of policy languages for path vector protocols. In *Proceedings of the 2003 conference on Applications, technologies, architectures, and protocols for computer communications*, pages 61–72. ACM, 2003.
 - [15] Timothy G. Griffin, F. Bruce Shepherd, and Gordon Wilfong. Policy disputes in path-vector protocols. In *Proceedings of the Seventh Annual International Conference on Network Protocols*, pages 21–30. IEEE Computer Society, 1999.
 - [16] Timothy G. Griffin and Gordon Wilfong. An analysis of BGP convergence properties. *SIGCOMM Computer Communication Review*, 29(4):277–288, 1999.
 - [17] Alex Hall, Evdokia Nikolova, and Christos Papadimitriou. Incentive-compatible interdomain routing with linear utilities. In *Proceedings of the 3rd International Workshop on Internet and Network Economics, WINE 2007*, volume 4858 of *Lecture Notes in Computer Science*, pages 232–244. Springer, 2007.
 - [18] Sergiu Hart and Yishay Mansour. The communication complexity of uncoupled nash equilibrium procedures. In *Proceedings of the thirty-ninth annual ACM Symposium on Theory of Computing (STOC 2007)*, pages 345–353. ACM, 2007.
 - [19] Johan. Håstad. Clique is hard to approximate within $n^{1-\epsilon}$. *Acta Mathematica*, 182:105–142, 1999.
 - [20] P. E. Haxell and G. T. Wilfong. A fractional model of the border gateway protocol (bgp). In *SODA '08: Proceedings of the nineteenth annual ACM-SIAM symposium on Discrete algorithms*, pages 193–199, Philadelphia, PA, USA, 2008. Society for Industrial and Applied Mathematics.
 - [21] Geoff Huston. Interconnection, peering, and settlements. In *Internet Global Summit (INET)*. The Internet Society, 1999.
 - [22] Eyal Kushilevitz and Noam Nisan. *Communication Complexity*. Cambridge University Press, 1997.
 - [23] Elias Koutsoupias and Christos H. Papadimitriou. Worst-case equilibria. In *the Proceedings of STACS*, 1999.
 - [24] Ahuva Mu'alem and Michael Schapira. Setting lower bounds on truthfulness. In *Proceedings of the eighteenth annual ACM-SIAM symposium on Discrete algorithms (SODA 2007)*, pages 1143–1152. Society for Industrial and Applied Mathematics, 2007.
 - [25] Rahul Sami, Michael Schapira, and Aviv Zohar. Searching for stability in interdomain routing. Working paper, 2008.
 - [26] Rahul Sami, Michael Schapira, and Aviv Zohar. Security and selfishness in interdomain routing. Working paper, 2008.
 - [27] Secure BGP Project Page. <http://www.ir.bbn.com/sbgp/>.
 - [28] Jeffrey Shneidman and David C. Parkes. Specification faithfulness in networks with rational nodes. In *Proceedings of the twenty-third annual ACM symposium on Principles of distributed computing (PODC 2004)*, pages 88–97. ACM, 2004.
 - [29] João L. Sobrinho. An algebraic theory of dynamic network routing. *IEEE/ACM Transactions on Networking*, 13(5):1160–1173, 2005.
 - [30] Kannan Varadhan, Ramesh Govindan, and Deborah Estrin. Persistent route oscillations in inter-domain routing. *Computer Networks*, 32(1):1–16, 2000.
 - [31] Shiva Kintali. A Distributed Protocol for Fractional Stable Paths Problem. Working paper, 2008.