TightRope: Towards Optimal Load-balancing of Paths in Anonymous Networks

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ABSTRACT

We study the problem of load-balancing in path selection in anonymous networks such as Tor. We first find that the current Tor path selection strategy can create significant imbalances. We then develop a (locally) optimal algorithm for selecting paths and show, using flow-level simulation, that it results in much better balancing of load across the network. Our initial algorithm uses the complete state of the network, which is impractical in a distributed setting and can compromise users' privacy. We therefore develop a revised algorithm that relies on a periodic, differentially private summary of the network state to approximate the optimal assignment. Our simulations show that the revised algorithm significantly outperforms the current strategy while maintaining provable privacy guarantees.

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1 INTRODUCTION

Users are increasingly turning to anonymous communication networks to protect themselves from surveillance, online tracking, or government censorship. The Tor network has several million daily active users [20] and has recently been integrated into the privacy-focused Brave browser [4].

To achieve anonymity in Tor, users' traffic is routed across a series of servers, called *relays*. There are several thousand relays, run by volunteers; each user's path through the network, called a *circuit*, typically transits three of them. This creates a load-balancing problem of assigning circuits to relays while ensuring no relay gets overloaded and all circuits receive good performance. Complicating this problem are the highly heterogeneous relay capacities—spanning some five orders of magnitude—and the privacy requirements of circuit construction. In particular, no one except the user must know the entirety of the circuit, precluding any centralized load-balancing solution.

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Currently, the Tor network uses a randomized assignment of flows circuits to relays, where each user chooses the relays for their circuits randomly weighted by their measured capacity (with some other constraints, see section 2.1). This ensures that each relay has the same *average* load; however, as we demonstrate in Section 3, this can create significant imbalances at any given point in time. We therefore consider the question of whether it is possible to provide better load balancing while satisfying the privacy requirements.

We first study a non-private load-balancing algorithm. We adapt an algorithm that calculates the max-min-fair allocation of bandwidth to circuits to select an optimal set of relays for a new path. We show that this results in significantly better load-balancing. Running this algorithm for each new flow would be impractical, therefore we create a batch version of the algorithm, which speculatively generates new circuits using the optimal algorithm and uses these circuits to induce a distribution over the relays, which is then sampled from to generate new circuits. Note that, unlike the Tor algorithm, the distribution reflects the *current* state of the network and is periodically refreshed; we show that this results in significantly better load-balancing performance than the Tor algorithm.

We then design a private version of this algorithm. Rather than working with a list of circuits, we break each hop into two 2-hop *segments*. We can then summarize the state of the network by creating a histogram of segments with one entry for each pair of relays. Our design is motivated by the fact that each entry in this histogram can be filled in by a single relay; moreover, each relay can *locally* add noise to the entry resulting in a differentially private histogram. We show that using a private histogram, we can implement a modified batch algorithm that approximates the optimal load-balancing. Our experiments show that the algorithm results in significantly better balanced circuits than the Tor randomized approach, while preserving privacy.

2 BACKGROUND

In this section we review the key properties of anonymous communication networks relevant to load balancing. We then present the max-min fair bandwidth allocation algorithm that we will use to model the load-balancing performance, and introduce the differential privacy framework that will be used to maintain users' anonymity.

2.1 Anonymous Networks and Tor

Anonymity networks provide users a way to communicate without revealing their identity, and without revealing their relationships

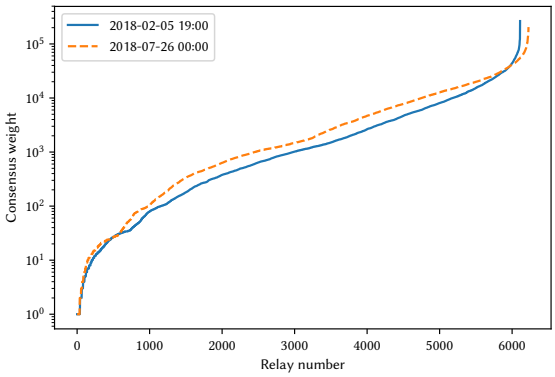


Figure 1: Measured relay bandwidths from two Tor consensus documents: February 5, 2018 at 19:00 (UTC) and July 26, 2018 at 00:00 (UTC). Note the logarithmic scale of the y-axis.

to third parties. Starting with Chaum's seminal mix network design [6], anonymity has been most frequently achieved by forwarding traffic through a series of servers in order to disguise its origin. In onion routing networks [24], each packet is multiply encrypted, with a layer of encryption being removed by each server in the path. This makes it impossible for any server to learn the entire path; rather, it knows only the preceding and following hops.

Tor [9] is the most popular anonymity network. In Tor, paths (called *circuits*) typically take three hops to transit the network. These hops are chosen from a collection of volunteer-run servers, called *relays*. These relays have vastly varying bandwidth capacity; in order to balance the load among them, their bandwidth is measured using TorFlow [17]. Relays are then allocated to circuits randomly, weighted by their measured bandwidth (also known as the consensus weight). The full Tor path selection algorithm [8] is somewhat complex because it must account for some relays not being usable in certain positions of the circuit as well as other constraints. For the purposes of this paper, we will approximate the algorithm as picking three random relays, without replacement, from the distribution induced by the measured bandwidth, leaving simulations of the full Tor algorithm for future work.

An important feature of the Tor network is that the relays, due to being supplied by volunteers, vary wildly in their bandwidth capacity. Figure 1 shows the relays and their measured capacity from two consensus documents take about five months apart. The distribution is highly skewed, with measured capacities spanning over five orders of magnitude. Note that the distribution in the two consensus documents follows a similar pattern; we therefore use the February 5, 2018 19:00 (UTC) consensus as representative for our experiments in the remainder of the paper.

2.2 Max-min Fairness

In this section we introduce our model of Tor performance using max-min fair bandwidth allocation. Our model of the Tor network includes two simplifying assumptions: (a) each user holds a single path through relays and (b) path capacities are constrained only by the relays, and not by the links between relays. The former can be easily adjusted by creating virtual users; the latter assumption is standard in analyzing Tor, and indeed central to the Tor bandwidth measurement and allocation architecture. We use the max-min fair allocation as a model because Tor schedules circuits in a round-robin fashion, which has been shown to achieve max-min fairness [12]. One further assumption is that each path is in simultaneous active use. We discuss some relaxations of this assumption in section 5.4, and defer more complex modeling and simulation of circuit usage to future work.

Notation. We introduce some notation for the rest of the paper. For any positive integer k, $[k]$ denotes the set $\{1, \ldots, k\}$. We denote the number of relays by n and the number of users (and therefore the number of paths) by m. We assign integer identifiers to the relays and users, and thus, the sets of relays and users are $[n]$ and $[m]$, respectively. The *capacity* of relay $r \in [n]$ is the positive constant $C[r]$. The *path* assigned to user $p \in [m]$ is a sequence of three relays and is denoted by $P[p]$. We identify this sequence with the the p^{th} *path*. Given an allocation of paths to all users, for any relay $r \in [n]$ we define $R[r] = \{p \in [m] \mid r \in P[p]\}$ to be the set of identifiers of the paths to which the relay r belongs.

Each relay r allocates some bandwidth to each of the paths in $R[r]$. The bandwidth allocated to the p^{th} path is the minimum bandwidth allocated for it by its three relays and is denoted by $band[p]$. For any relay r, the total allocated bandwidth to all paths in $R[r]$, must be less than the capacity of r:

$$\forall\, r \in [n], \sum_{p \in R[r]} band[p] \leq C[r]. \tag{1}$$

Allocations satisfying eq. (1) are said to be *feasible*.

A feasible allocation $band$ is *max-min fair* if and only if an increase of bandwidth allocation to any path (within the set of feasible allocations), must be at the cost of a decrease in allocation of another path with an already lower bandwidth in $band$ (See Section 6.5.2 in [3]). That is, for any other feasible allocation $band'$ and any path $p_1 \in [m]$, if $band'[p_1] > band[p_1]$, then there exists $p_2 \in [m]$ such that $band'[p_2] < band[p_2]$ and $band[p_2] \leq band[p_1]$.

It is well-known that the allocation algorithm shown below (algorithm 1) achieves max-min fairness. It takes as input a network of relays and paths, and allocates a bandwidth to each path in an iterative fashion. Specifically, the inputs are the array or map C of all the relay capacities, the array of user paths P, and the array R. The algorithm keeps track of the *residual capacity*, $Cres$, of each relay after subtracting the bandwidths of the paths passing through it. It also keeps track of the *residual paths*, $Rres$, that is, the set of paths passing through each relay after removing those paths whose bandwidths are already allocated. At each iteration, one relay $r^* \in [n]$ is chosen and each path in $Rres[r^*]$ is allocated a bandwidth. The chosen relay r^* is the one that has the smallest ratio $Rat[r] := Cres[r]/|Rres[r]|$ at the corresponding iteration (line 7). After it is chosen, each of the paths in $Rres[r^*]$ is assigned a bandwidth of $Rat[r^*]$ (line 9). Relay r^* is called the *bottleneck relay* of these paths. Then, these paths are removed from their corresponding relays (line 12) and the capacities of these relays get subtracted by $Rat[r^*]$ (line 11). This is repeated until all paths are allocated bandwidth.

Algorithm 1 Max-min Bandwidth Allocation Algorithm

1: **input:** C, R, P
2: $Cres[r] \leftarrow C[r], \forall\, r \in [n]$
3: $Rres[r] \leftarrow R[r], \forall\, r \in [n]$
4: $band[p] \leftarrow 0, \forall\, p \in [m]$
5: **while** $\exists\ p \mid band[p] = 0$ **do**
6: $\quad Rat[r] \leftarrow \begin{cases} \frac{Cres[r]}{|Rres[r]|} & \forall\, r \in [n] \mid |Rres[r]| \neq 0 \\ \infty & otherwise \end{cases}$
7: $\quad r^* \leftarrow \underset{r \in [n]}{\text{argmin}} \quad Rat[r]$
8: $\quad$ **for** $p \in Rres[r^*]$ **do**
9: $\quad\quad band[p] \leftarrow Rat[r^*]$
10: $\quad\quad$ **for** $r \in P[p]$ **do**
11: $\quad\quad\quad Cres[r] \leftarrow Cres[r] - Rat(r^*)$
12: $\quad\quad\quad Rres[r] \leftarrow Rres[r] \setminus \{p\}$
13: **return** *band*

Remark 1. *Suppose the relay r^* chosen at line 7 of algorithm 1 belongs to a path p. Then, path p is allocated bandwidth of $Rat[r^*]$, that is $band[p] = Rat[r^*]$. Further, this allocation is not changed in subsequent iterations.*

2.3 Differential Privacy

To perform load-balancing, we would like to incorporate feedback about the state of the network into the path selection process. However, as discussed above, the state of the network is explicitly required to be private, as this is key to preserving users' anonymity. We will use differential privacy to ensure our feedback mechanism does not result in privacy loss.

Differential privacy was first proposed by Dwork [10]. It formalizes the notion that a mechanism operating over a private data set must produce an output that depends only minimally on each item in the data set. We will use the formulation given by Vadhan [25]:

Definition 1 ((Approximate) differential privacy). [25, Definition 1.4] For $\epsilon \geq 0, \delta \in [0, 1]$ we say that a randomized mechanism $\mathcal{M} : \chi^n \times \Omega \rightarrow \mathcal{Y}$ is (ϵ, δ)-differentially private if for every pair of neighboring datasets $x \sim x' \in \chi^n$ (i.e. x and x' differ in one row), and every query $q \in \Omega$, we have:

$$\forall T \subseteq \mathcal{Y}, Pr[\mathcal{M}(x, q) \in T] \leq e^{\epsilon} \cdot Pr[\mathcal{M}(x', q) \in T] + \delta,$$

where Ω is the set of possible queries. Moreover, δ should typically satisfy $\delta \leq n^{-\omega(1)}$ for this definition to be meaningful.

In our case, the dataset in question will be the complete list of circuits in the Tor network, with each circuit representing a row. As a result, differential privacy will guarantee the privacy of each individual circuit while providing aggregate traffic statistics. We note that differential private mechanisms have previously been used to study traffic properties of Tor [15].

3 ANALYZING THE RANDOMIZED ALLOCATION ALGORITHM

In this section, we will present the current method used to create paths in the Tor network. Incoming users sample without replacement three relays using the distribution over relays where each relay is weighed by its capacity.

Algorithm 2 Random Path Allocation Algorithm

1: **input:** C
2: Sample three relays $\{r_1, r_2, r_3\}$ without replacement from the set of relay where each relay is weighed by its capacity.
3: **return:** $\{r_1, r_2, r_3\}$

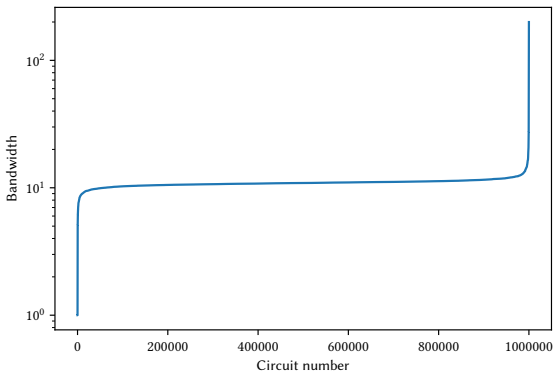


Figure 2: Bandwidth allocation of 1 million paths using the random algorithm 2.

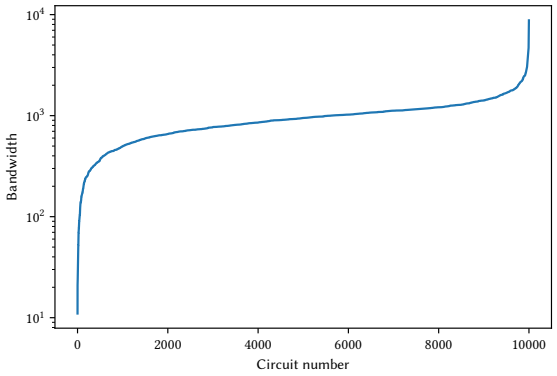


Figure 3: Bandwidth allocation of 10 000 paths using the random algorithm 2.

We evaluate the performance of this algorithm with respect to the parameters of the Tor network. Jansen and Johnson [15] estimated that there were approximately 1.2 million active circuits (95% CI +/- 500,000) in the network. We therefore created 1 million paths using the random algorithm 2 and then computed their bandwidth allocation using max-min fair algorithm 1. The results are in fig. 2. We can see that the majority of circuits receive an allocation close to the average of 10.9, but there are significant tails at both ends of the distribution: the minimum circuit has allocation of 1.0 and the maximum of 190. (The standard deviation is 1.28.)

Jansen and Johnson define an active circuit as one that has *ever* been used to forward data; however, at any given moment, most

circuits are idle. This can be seen by comparing the data about the aggregate Tor network traffic of approximately 100 Gbps [19] to the performance of a sustained download, which is roughly 10 s for 5 MiB [18]. Given that each circuit gets carried by 3 relays, this suggests that 100 Gbps/3/(5 MiB/10 s) ≈ 10 000 circuits are active at any given time. Figure 3 shows the bandwidth allocation of 10 000 circuits generated by algorithm 2. Observe that the imbalances in this case are much more significant: the minimum allocation is 14 and the maximum is 6419, with a standard deviation of 475.8. 932 out of 10 000 flows receive less than half of the average bandwidth of 980, and 37 receive less than 10%.

4 LOCALLY OPTIMAL INCREMENTAL PATH ALLOCATION

In this section, we modify the max-min fair allocation algorithm 1 to design an algorithm that, given the state of the Tor network, returns three relays that would result in an optimal allocated bandwidth for a new path. The result is algorithm 3. This algorithm assumes that the bandwidths of the existing paths in the network are allocated using max-min fairness (algorithm 1).

4.1 Algorithm Description

The algorithm iteratively creates a list B of $(bandwidth, relay)$-pairs. This list determines how much bandwidth would be allocated to a newly added path to the network. That is, of the relays appearing in a new path, the one that appears earliest in B determines the bandwidth of the path.

The idea of algorithm 3 is to simulate the behaviour of the max-min fairness algoirthm algorithm 1 on the network when an arbitrary new path is added. This simulation allows us to know how much bandwidth it would get allocated. A trivial but key observation is that when a new path is added, the number of paths passing through each of its three relays will be incremented by one. For all other relays, the number of paths will remain unchanged. Moreover, as per remark 1, the relay that is chosen first from a path determines the path's bandwidth allocation. Since we are searching for the relays that would maximize the bandwidth allocation for a newly added path, algorithm 3 computes the different possible allocations based on the different possible bottlenecks.

In addition to the ratio $Cres[r]/|Rres[r]|$ tracked in algorithm 1 (line 7), algorithm 3 also tracks $Cres[r]/(|Rres[r]|+1)$, for each relay r (line 8). In other words, Rat is now a $2 \times n$ matrix: row 1 stores $Cres[r]/|Rres[r]|$ and row 2 stores $Cres[r]/(|Rres[r]|+1)$.

At each iteration, a minimum of all the $2n$ ratios is chosen. We denote the minimizing row by $t^* \in \{1,2\}$ and the corresponding relay by $r^* \in [n]$. If the minimizing row $t^* = 1$, the algorithm proceeds as max-min fair allocation algorithm 1 by allocating bandwidth of $Rat[1,r^*]$ to each of the paths in $Rres[r^*]$. Then, it removes them from the other relays in which they pass (lines 17 and 18). Both ratios in the same column of Rat, i.e., ratios corresponding to the same relay, are updated in the same way unless the ratio in the second row is already added to B (line 8). That is because both ratios use the same arrays $Cres$ and $Rres$ for their numerator and denominator. Otherwise, if $t^* = 2$, the pair $(Rat[2,r^*], r^*)$ is added to the end of B. It is added at the end since the r^* will be the bottleneck of the added path only when its other relays are not in already in B at this iteration. If one of its relays is already in B at this iteration, that would be its bottleneck instead of r^*. This will be proved in the next section. The algorithm iterates until all n relays have entries in B.

Algorithm 3 Locally Optimal Path Allocation Algorithm

1: **input:** C, R, P
2: $B \leftarrow \emptyset$
3: $Cres[r] \leftarrow C[r], \forall\, r \in [n]$
4: $Rres[r] \leftarrow R[r], \forall\, r \in [n]$
5: $band[p] \leftarrow 0, \forall\, p \in [m]$
6: **while** $\exists\ i \notin B$ **do**
7: $\quad Rat[1,r] \leftarrow \begin{cases} \frac{Cres[r]}{|Rres[r]|} & \text{if } |Rres[r]| \neq 0 \\ \infty & \text{otherwise} \end{cases}$
8: $\quad Rat[2,r] \leftarrow \begin{cases} \frac{Cres[r]}{|Rres[r]|+1} & \text{if } r \notin B \\ \infty & \text{otherwise} \end{cases}$
9: $\quad (t^*, r^*) \leftarrow \underset{t \in \{1,2\}, r \in [n]}{\text{argmin}} Rat[t,r]$
10: $\quad$ **if** $t^* == 2$ **then**
11: $\quad\quad B.push([Rat[t^*, r^*], r^*])$
12: $\quad\quad Rat[t^*, r^*] \leftarrow \infty$
13: $\quad$ **else**
14: $\quad\quad$ **for** $p \in Rres[r^*]$ **do**
15: $\quad\quad\quad band[p] \leftarrow Rat[t^*, r^*]$
16: $\quad\quad\quad$ **for** $r \in P[p]$ **do**
17: $\quad\quad\quad\quad Cres[r] \leftarrow Cres[r] - Rat[t*, r*]$
18: $\quad\quad\quad\quad Rres[r] \leftarrow Rres[r] \setminus \{p\}$
19: Let $((b_1, r_1), (b_2, r_2), (b_3, r_3))$ be the last three elements of B
20: **return:** $\{r_1, r_2, r_3\}$

4.2 Algorithm Correctness

In this section, we will prove that the output of algorithm 3 is a path with maximum bandwidth allocation possible, for a new path that is to be added to the given network. In this analysis, we will compare the state of algorithm 3 with the state of max-min fair allocation algorithm 1, in the same iteration. We will add bars on top of the variable names of algorithm 3 to distinguish them from the variables with the same names in algorithm 1. A subscript of zero refers to the initial values of the variables. A subscript $l > 0$ denotes the value of the variable at the end of the l^{th} while-loop iteration, for the corresponding program[1]. For example, $\overline{Cres}_2$ is the value of $Cres$ of algorithm 3 at the end of the second iteration of its while loop. We will use $\oplus$ to represent sum of sets.

The following key lemma is used to prove an equivalence between the behaviors of algorithm 1 and algorithm 3: given any new path p, the bandwidth allocated to p by algorithm 1 equals the bandwidth associated with the relay in p that appears earliest in $\overline{B}$ (computed by algorithm 3).

Lemma 1. *Let the new path be $H = \{h_1, h_2, h_3\} \in [n]^3$. Assume (a) $C = \overline{C}$, (b) $R[r] = \overline{R}[r]$ for all $r \in [n] \setminus H$ and $R[r] = \overline{R}[r] \cup \{m+1\}$ for $r \in H$, and (c) $P[p] = \overline{P}[p]$ for all $p \in [m]$, and $P[m+1] = H$.*

[1]For algorithm 1, it is the value of the variable after the execution of line 12 and for algorithm 3 it is the value after the execution of line 18

Assume w.l.o.g that h_1 is the first relay to be added by algorithm 3 to $\overline{B}$. Let $l_1, l_2, \dots l_k$ be the iterations in algorithm 3 at which $\overline{t}^* = 1$ before h_1 is added to $\overline{B}$. Then, for all $s \leq k$, $Cres_s = \overline{Cres}_{l_s}$ and $Rres_s[r] = \overline{Rres}_{l_s}[r]$ for all $r \in [n] \setminus H$ and $Rres_s[r] = \overline{Rres}_{l_s}[r] \cup \{m+1\}$ for $r \in H$.

Proof. First, $Cres_0 = C_1 = C_2 = \overline{Cres}_0$. Second, note that $Rat_1[r] = \overline{Rat}_1[1, r]$ for all $r \in [n] \setminus H$ and $Rat_1[r] = \overline{Rat}_1[2, r]$ for all $r \in H$. Hence, if the minimum ratio at the first iteration of algorithm 3 exists in Rat_1, that same ratio will be chosen by algorithm 1 at its first iteration. Thus, if $\overline{t}_1^* = 2$ and $\overline{r}_1^* = h_1$ ($\overline{r}_1^*$ cannot be h_2, h_3 as we assumed that h_1 is chosen first), then the lemma would hold with $k = 0$ and $r_1^* = h_1$.

If $\overline{t}_1^* = 2$ and $\overline{r}_1^* \neq h_1$, then the minimum ratio of algorithm 3 does not belong to Rat_1 and neither $\overline{Cres}$ nor $\overline{Rres}$ would be changed in this iteration, i.e. $\overline{Cres}_1 = \overline{Cres}_0$ and $\overline{Rres}_1 = \overline{Rres}_0$. In that case, Rat_2 would still be a subset of $\overline{Rat}_2$.

The case where $\overline{t}_1^* = 1$ and $\overline{r}_1^* \in H$ cannot happen since $\overline{Rat}_1[2, r] \leq \overline{Rat}_1[1, r]$ for all $r \in [n]$.

Finally, if $\overline{t}_1^* = 1$ and $\overline{r}_1^* \notin H$, then $l_1 = 1$ and both algorithms will have the same minimum ratio, the else branch would be taken in algorithm 3 and $\overline{Cres}$ and $\overline{Rres}$ would be updated in the same manner as those $Cres$ and $Rres$. Thus, the property in the lemma would be preserved.

Hence, the above analogy can be repeated to shown that the l^{th} iteration of algorithm 3 at which $\overline{t}^* = 1$ will run the same updates as that of the l^{th} iteration of algorithm 1 till h_1 is added to $\overline{B}$. Iterations of the while loop in algorithm 3 at which $\overline{t}^* = 2$ does not affect $\overline{Cres}$, $\overline{Rres}$, and the ratios common with algorithm 1 until h_1 is chosen. Once h_1 is added to $\overline{B}$, that corresponds to the $k+1^{th}$ iteration of algorithm 1 where h_1 will be chosen as the minimizing relay too and the bandwidth of the new path would be determined. □

Corollary 1. *For any new path $H = \{h_1, h_2, h_3\}$. The bandwidth associated with h_1 in $\overline{B}$ (algorithm 3) equals the bandwidth allocated for H by algorithm 1.*

Since in the following lemma we will be only analyzing algorithm 3, there will be no confusion with the variables of algorithm 1 so we drop the bars.

Lemma 2. *The ratios in B appear in increasing order.*

Proof. Consider the l^{th} iteration of algorithm 3 at which a ratio-relay pair (b, r_l^*) is added to B, i.e. $t_l^* = 2$. If in the preceding iteration $l-1$, $t_{l-1}^* = 2$, another ratio-relay (b_1, r_{l-1}^*) would have been added to B before it. Moreover, both $Cres$ and $Rres$ would not have changed and since b_1 was chosen first means it is smaller than b_2 and thus in this case the B would be increasing. Otherwise, if $t_{l-1}^* = 1$, the "else" branch would be taken. Denote the number of paths r_{l-1}^* and r_l^* share at the $(l-1)^{th}$ iteration be s. Then, $Rat_l[2, r_l^*] = \frac{Cres_{l-1}[r_l^*] - Rat_{l-1}[1, r_{l-1}^*]}{|Rres_{l-1}[r_l^*]|+1-s}$. We will show that it is larger than $Rat_{l-1}[1, r_{l-1}^*] = \frac{Cres_{l-1}[r_{l-1}^*]}{|Rres_{l-1}[r_{l-1}^*]|}$.

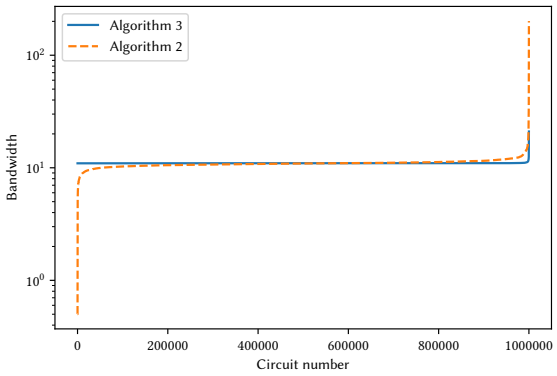


Figure 4: Graph comparing the bandwidth allocation of one million paths generated using the locally optimal algorithm 2 and the random algorithm 3. Note the logarithmic scale of the y-axis.

We know that $Rat_{l-1}[2, r_l^*] := \frac{Cres_{l-1}[r_l^*]}{|Rres_{l-1}[r_l^*]|+1} > \frac{Cres_{l-1}[r_{l-1}^*]}{|Rres_{l-1}[r_{l-1}^*]|}$. Hence,

$$
\begin{aligned}
& Cres_{l-1}[r_l^*]|Rres_{l-1}[r_{l-1}^*]| > Cres_{l-1}[r_{l-1}^*](|Rres_{l-1}[r_l^*]| + 1) \\
\implies & Cres_{l-1}[r_l^*]|Rres_{l-1}[r_{l-1}^*]| - s(Cres_{l-1}[r_{l-1}^*]) \\
& \qquad > Cres_{l-1}[r_{l-1}^*](|Rres_{l-1}[r_l^*]| + 1 - s) \\
\implies & |Rres_{l-1}[r_{l-1}^*]|(Cres_{l-1}[r_l^*] - s\frac{Cres_{l-1}[r_{l-1}^*]}{|Rres_{l-1}[r_{l-1}^*]|}) \\
& \qquad > Cres_{l-1}[r_{l-1}^*](|Rres_{l-1}[r_l^*]| + 1 - s) \\
\implies & \frac{Cres_{l-1}[r_l^*] - s\frac{Cres_{l-1}[r_{l-1}^*]}{|Rres_{l-1}[r_{l-1}^*]|}}{|Rres_{l-1}[r_l^*]| + 1 - s} > \frac{Cres_{l-1}[r_{l-1}^*]}{|Rres_{l-1}[r_{l-1}^*]|}.
\end{aligned}
$$

Therefore, the ratios are non decreasing in the iterations at which $t^* = 1$ and constant (other than the one added to B) when $t^* = 2$ which means that every time a ratio is added to B, it will be equal or larger than all previously added ones. □

Corollary 2. *Choosing the last three relays of the resulting B for the new path would result in maximum bandwidth allocated for it by algorithm 1. That bandwidth is the one associated with the third relay from the end of B.*

4.3 Results in Tor Network

Starting with an empty set, one million paths were created by repetitively running our locally optimal algorithm 3 while adding the output path to the network. The bandwidth allocation of the one million paths generated is found using the max-min fair algorithm 1. Those results were then compared to the bandwidth allocations of one million paths generated using the random algorithm 2. The results are shown in fig. 4.

The locally optimal algorithm 3 produces a much more well-balanced set of paths: the minimum allocation if 10.96 is nearly the same as the average of 10.97, and the maximum is 21. In contrast, the paths created using the random algorithm 2, while having a

similar average bandwidth allocation of 10.94, span a much broader range, with a minimum allocation of 1 and a maxium of 210.14.

5 DIFFERENTIALLY PRIVATE ALGORITHM:

The locally optimal algorithm 3 requires the knowledge of the state of the network as input. Thus, if it is going to be used, every user will need to know the state of the network, i.e., the paths of every other user, in order to be able to construct a path for herself. This will defeat the purpose of onion routing as it will then be possible to deanonymize users. In this section, we will be discussing how to implement a differentially private version of algorithm 3.

In the private version of the algorithm, we first decompose each circuit (r_1, r_2, r_3) into two circuit segments, (r_2, r_1) and (r_2, r_3). We then adapt the locally optimal algorithm 3 to use these segments, rather than complete circuits, in creating an optimal new path. To add privacy, we summarize the list of segments as a histogram, indexed by pairs of relays, of the number of circuit segments (r_i, r_j) that are present. We then create a differentially private version of this histogram by using a threshold-based differentially private count, to account for the sparse nature of the histogram.

One feature of the private count algorithm is that each histogram entry can be processed individually. Therefore, a relay r_i can apply it to the count of each histogram entry (r_i, r_j), since it knows the (actual, non-private) number of such flow segments. The private counts can then be aggregated and distributed using a modification of the existing Tor directory mechanism or another peer-to-peer broadcast scheme.

Since the private histogram can only be updated periodically, we use this histogram to generate the next N near-optimal circuits. We cannot assign these circuits directly to each new user; instead, we count the number of times each relay appears in these N circuits and use it to induce a distribution over relays that the users sample from for their circuit. This approach is similar to the random algorithm 2, except that the weights reflect relays that are underloaded in the current state of the network, rather than the static relay capacities.

5.1 Pair-baserd Algorithm Description

As in the locally optimal algorithm 3, we denote the capacity of the r^{th} relay by $C[r]$. As discussed before, the paths are now ordered pairs of relays. Being ordered is essential for the correctness of the pair-based algorithm 4 as will be discussed later. We denote by P the map from these paths to the ordered pairs of relays to which they belong. For example, those corresponding to the p^{th} path are $P[p] = (r_{p,1}, r_{p,2})$. Also, as before, we define $R[r]$ to be the set $\{p \mid r \in P[p]\}$ of paths to which relay r belongs. However, we decompose R into two maps: R_c mapping relays to the paths in which they appear in the first component of the ordered pair, i.e. the central relay, and R_e mapping relays to the paths in which they appear in the second component of the ordered pair, i.e. the end relay.

We finally define $Nres[r]$ as the number of actual paths (paths with three relays) containing relay r. We can compute it as: $Nres[r] = |Rres_c[r]|/2 + |Rres_e[r]|$, since for an actual path in the Tor network where r is the central relay, it will appear as two pairs in $Rres_c[r]$ while it should be counted once. It will only appear once in $Rres_e[r]$ if r is one of its end relays.

In algorithm 4, as in the previous two algorithms, in the iterations at which $t^* = 1$, $Rat[1, r^*] = \frac{Cres[r]}{Nres[r]}$ of bandwidth would be allocated to each of the paths passing through it. After that, the paths in $Rres_e[r^*]$, $Rat[1, r^*]/2$ would be deducted from the residual capacity of the central relay in the path. That is because the bandwidth of the other path that corresponds to the same actual path will be subtracted from the residual capacity of the central relay in that path too. For the paths in $Rres_c[r^*]$, $Rat[1, r^*]$ would be deducted entirely from the residual capacity of the end relay of the path. That is because the end relay of an actual path of the Tor network only appears once in the two-relay paths corresponding to that path. Using the same analogy, for the paths in $Rres_c[r^*]$, $Nres[r]$, r being the end relay in the path, is deducted by one and for paths in $Rres_e[r^*]$, $Nres[r]$, r being the central relay in the path, is deducted by half. The rest of the algorithm follows the same steps of algorithm 3.

Algorithm 4 Pair-based Locally Optimal Path Allocation Algorithm

input: C, R, R_c, R_e, P
$B \leftarrow \emptyset$
$Cres[r] \leftarrow C[r], \forall\, r \in [n]$
$Rres[r] \leftarrow R[r], \forall\, r \in [n]$
$band[p] \leftarrow 0, \forall\, p \in [m]$
$Nres[r] \leftarrow \frac{|Rres_c[r]|}{2} + |Rres_e[r]|, \forall\, r \in [n]$
while $\exists\ i \notin B$ **do**
    $Rat[1, r] \leftarrow \begin{cases} \frac{Cres[r]}{Nres[r]} & \forall\, r \mid Nres[r] \neq 0 \\ \infty & otherwise \end{cases}$
    $Rat[2, r] \leftarrow \begin{cases} \frac{Cres[r]}{Nres[r]+1} & \text{if } r \notin B \\ \infty & otherwise \end{cases}$
    $(t^*, r^*) \leftarrow \underset{t \in \{1,2\}, r \in [n]}{\text{argmin}} Rat[t, r]$
    if $t^* == 2$ **then**
        $B.push([Rat[t^*, r^*], r^*])$
        $Rat[t^*, r^*] \leftarrow \infty$
    else
        for p $\in Rres[r*]$ **do**
            $band[p] \leftarrow Rat[t^*, r^*]$
            for $r \in P[p]$ **do**
                if $p \in Rres_c[r]$ **then**
                    $Cres[r] \leftarrow Cres[r] - Rat[t*, r*]/2$
                    $Nres[r] \leftarrow Nres[r] - 1/2$
                    $Rres[r] \leftarrow Rres[r] \setminus \{p\}$
                    $Rres_c[r] \leftarrow Rres_c[r] \setminus \{p\}$
                else
                    $Cres[r] \leftarrow Cres[r] - Rat[t*, r*]$
                    $Nres[r] \leftarrow Nres[r] - 1$
                    $Rres[r] \leftarrow Rres[r] \setminus \{p\}$
                    $Rres_e[r] \leftarrow Rres_e[r] \setminus \{p\}$
return: $\{r_1, r_2, r_3\}$

Hence, algorithm 4 takes as an input a path matrix of ordered pairs of relays along with the capacity matrix. The output is the path of three relays that when added to the system gives the highest

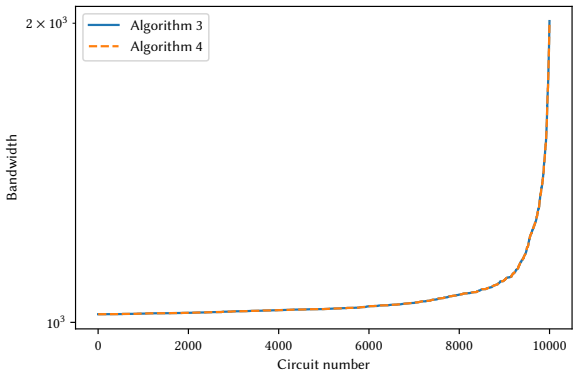


Figure 5: Graph comparing the bandwidth allocation of 10 000 paths generated using the pair-based algorithm 4 and the locally optimal algorithm 3.

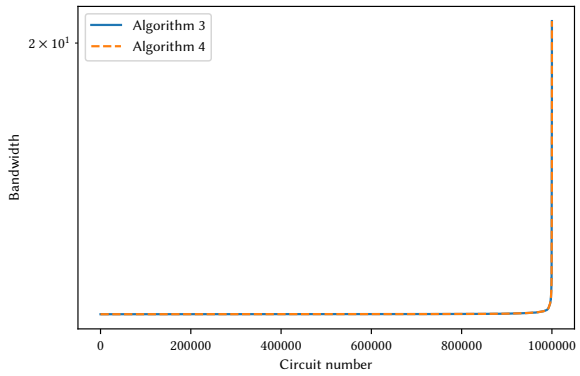


Figure 6: Graph comparing the bandwidth allocation of one million paths generated using pair-based algorithm 4 and locally optimal algorithm 3 (in blue). Note the logarithmic scale of the y-axis.

bandwidth allocation compared to any other path that can be added to the system.

5.2 Experimental Results

Starting from an empty set, the pair-based algorithm 4 was used repetitively to generate a set of 10 000 paths of three relays. At each run, a path matrix representing the decomposition of the paths in the network into segments of two relays is constructed, and then input to algorithm 4 along with the capacity matrix. The output, the path of three relays, is then added to the network. The bandwidth allocations for these paths are then computed using the max-min fairness algorithm 1.

The results were then compared to the bandwidth allocations of 10 000 paths generated by repetitive application of locally optimal algorithm 3 starting from an empty set to know how much accuracy we lost by the decomposition of paths into pairs. The results are shown in fig. 5. The two curves coincide which shows there is zero loss of accuracy.

The same experiment is repeated but instead 1 million paths were created using both algorithms. The results are shown in fig. 6. The maximum difference between the two allocations was 0.918.

5.3 Batch Path Allocation Algorithm Results

As we discussed earlier, the server would periodically collect data from the relays and create a histogram mapping each ordered pair of relays to the number of paths passing through them. It would then create a differentially private version of this histogram. After that, given the current state of the network, it runs the pair-based algorithm 4 repetitively to generate K additional paths. These paths are not added to the actual network but to a virtual copy of the network. The number of times each relay appeared in these K paths is counted to generate a probability distribution over the relays. The distribution is then released to the public. Incoming users would sample that distribution (without replacement) to get three different relays which would constitute their paths.

The batch algorithm 5 creates a batch of L random paths using the procedure we just described. L represents the expected number of paths that would be created in a single period before the distribution gets updated.

Algorithm 5 Batch Path Allocation Algorithm

1: **input:** C, R, R_c, R_e, P, K, L
2: Repeat K times:
3: Input C, $Rres$, R_c, R_e, P to algorithm 4.
4: Add the returned path to the network.
5: Update $Rres$, R_c, R_e, P accordingly.
6: Compute the distribution of the relays in the added K paths.
7: Sample L paths with three relays from that distribution
8: **return:** Generated L paths

To evaluate the behaviour of the batch algorithm 5, K is set to 10 000 and L to 200. Then, starting from a network with no paths, we ran it repetitively to create a network with 600 000 paths, 200 at a time. At each run, a path matrix is constructed by decomposing the paths in the network to paths of two ordered relays as previously discussed and taken as input for the next run.

After that, the bandwidth allocations of the created paths were computed using the max-min fair algorithm 1 and compared to the bandwidth allocations of 600 000 paths generated using the the random algorithm 2. The results are shown in fig. 7.

The minimum bandwidth of a path generated using the batch algorithm 5 was 6.33, the maximum was 37.23 and the average was 17.29. On the other hand, for the set of paths generated randomly, the minimum bandwidth of a path was 1, the maximum was 287.99 and the average was 17.23.

The simulation demonstrates the fairness property of algorithm 5 The range of the bandwidth allocations of the paths is [6.33, 37.23] compared to a range of [1, 287.99] for the random paths, while the average is being conserved. The algorithm hence avoids the

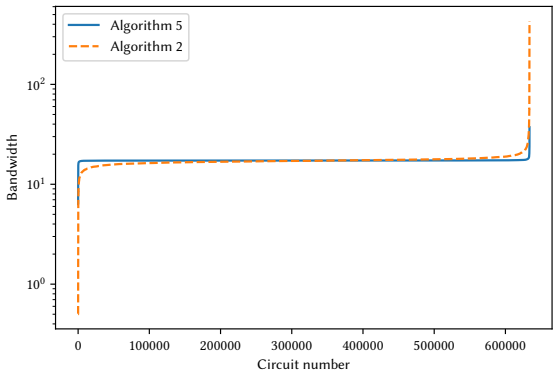


Figure 7: Graph comparing the bandwidth allocation of 600 000 paths generated using the random algorithm 2 and the batch algorithm 5. Note the logarithmic scale of the y-axis.

use of paths with very low bandwidth and guarantees a more fair distribution of the bandwidth between users.

5.4 Adding Differential Privacy

In this section, we show how the server ensures differential privacy of the statistics it releases about the network. As discussed earlier, at the end of each period, it generates a differentially private version of a histogram (matrix) of size n^2. In this paper, we use the notion (ϵ, δ)-differential privacy defined in section 2.3. We did not use the usual ϵ-differential privacy since the histogram that we are making private is very large ($n^2 \approx 36\,000\,000$) and is very sparse. Using (ϵ, δ)-differential privacy allows us to operate on the sparse histogram, as described below, rather than producing a noisy version of each 0 value in the histogram, which would overwhelm the algorithm. We replace n with n^2 in the original definition since in our case the dataset is of size n^2.

The following theorem shows that there is a mechanism that ensures the differential privacy of the histogram against queries consisting of point functions while guaranteeing an acceptable level of accuracy of query responses. Before stating the theorem, let us define formally the set of queries for which the mechanism ensures privacy.

Definition 2 (page 6 in [25]). Point Functions (Histograms): Let $\mathcal{X}$ be an arbitrary set and for each $y \in \mathcal{X}$, we consider the predicate $q_y : \mathcal{X} \to \{0, 1\}$ that evaluates to 1 only on input y. The family $Q^{pt} = Q^{pt}(\mathcal{X})$ consists of the counting queries corresponding to all point functions on data universe $\mathcal{X}$. (Approximately) answering all of the counting queries in Q^{pt} amounts to (approximately) computing the histogram of the dataset.

As before we substitute n by n^2 from the original theorem.

Theorem 3. *(stability-based histograms [5]). For every finite data universe χ, $n \in \mathbb{N}$, $\epsilon \in (0, 2\ln n)$, and $\delta \in (0, 1/n^2)$ there is an (ϵ, δ)-differentially private mechanism $\mathcal{M} : \chi^{n^2} \to \mathbb{R}^{\chi}$ that on every dataset $x \in \chi^{n^2}$, with high probability $\mathcal{M}(x)$ answers all of the counting queries in $Q^{pt}(\chi)$ to within error*

$$O\left(\frac{\log(1/\delta)}{\epsilon n^2}\right)$$

In the proof of the theorem, the authors provided such mechanism which takes a dataset (histogram) $x \in \chi^{n^2}$ as input and returns a privatized version of it. The mechanism is shown in algorithm 6. It iterates over the elements of the histogram, those that are zero are kept zero. A positive entry would be added to an independent Laplace variable with $\lambda = 2/(\epsilon n^2)$. If the result is below the threshold $((2\ln(2/\delta))/(\epsilon)) + (1)$, the entry would be set to zero, otherwise it is set to the result. Note that we multiplied the threshold that is in the mechanism by n^2 since we do not need the result of the query to be normalized, i.e. between 0 and 1, so we do not need the n^2 factor.

Algorithm 6 (ϵ, δ)-differentially Private Histogram Mechanism

input: x, χ
For every $y \in \chi$:
  If $q_y(x) = 0$ then:
    Set $a_y = 0$
  If $q_y(x) > 0$ then:
    Set $a_y \leftarrow q_y(x) + Lap(2/(\epsilon n^2))$.
    If $a_y < 2\ln(2/\delta)/\epsilon n^2 + 1/n^2$ then:
      Set $a_y = 0$
return: $(a_y)_{y \in \chi}$

As discussed earlier, the server releases a distribution over the relays at the beginning of every period based on the generated private histogram. Thus, information about the paths that remain over multiple periods in the network will be released several times. This will deteriorate the level of privacy they are guaranteed. To bound this deterioration we consult the following composition theorem of differential privacy.

Theorem 4. *(Composition of (ϵ, δ)-differentially-private algorithms, Theorem 16 in [11]). Let $T_1 : D \to T_1(D)$ be (ϵ, δ)-differentially-private, and for all $J \geq 2$, $T_j : (D, s_1, ..., s_{J-1}) \to T_J(D, s_1, ..., s_{J-1}) \in \zeta_J$ be (ϵ, δ)-differentially-private, for all given $(s_1, ..., s_{J-1}) \in \otimes_{j=1}^{J-1} \zeta_j$, where "$\otimes$" denotes direct product of spaces. Then for all neighboring D, D' and all $S \subseteq \otimes_{j=1}^{J-1} \zeta_j$:*

$$P((T_1, ..., T_J) \in S) \leq e^{J\epsilon} P'((T_1, ..., T_J) \in S) + J\delta$$

We can conclude that the parameters of privacy ϵ and δ that we can guarantee for a certain path increase linearly in the number of periods it stays in the network.

We chose the parameters ϵ and δ to be 0.3 and 0.001, respectively. To evaluate the behaviour of algorithm 5 after we privatize the histogram before using it to generate the distribution over relays, we conducted the following experiment: starting from an empty network, we generated $N = 1\,000\,000$ paths using repetitive application of batch algorithm 5 with $K = 10\,000$ and $L = 200$ while using a private version of the histogram before each run, using algorithm 7, and generating the distribution from the result.

The bandwidth allocations of those paths are then computed using max-min fair allocation algorithm 1 and compared to the

Algorithm 7 (ϵ, δ)-differentially Private Optimal Path Allocation Algorithm

1: **input:** $C, R, R_c, R_e, P, N, K, L$
2: **for** $i \in [\lceil \frac{N}{L} \rceil]$ **do**
3: A histogram mapping the pairs of relays in the network to the number of paths between each pair is constructed.
4: A private version of the histogram is generated using the mechanism described in algorithm 6.
5: The private histogram is input to the batch algorithm 5 with parameters K and L.
6: The generated L paths are added to the network.
7: Update R, R_c, R_e, P.

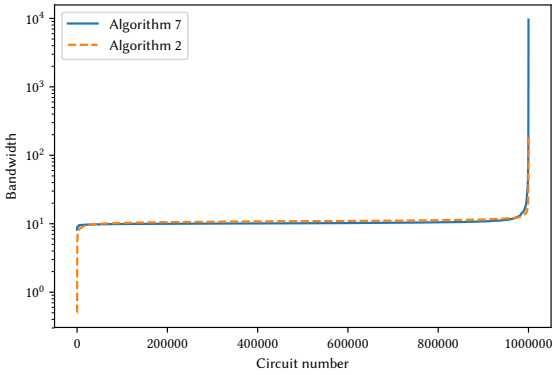


Figure 8: Graph comparing the bandwidth allocation of one million paths generated using the differentially private algorithm 7 and using the random algorithm 2. Note the logarithmic scale of the y-axis.

bandwidth allocations of one million paths generated using the random algorithm 2. The sorted (in bandwidth) results are shown in fig. 8.

The minimum bandwidth allocation of a path of differentially private algorithm 7 was 8.2, the maximum was 9603.5 and the average was 10.54. While for those generated randomly using algorithm 2, the minimum was 1, the maximum was 210.14 and the average was 10.94. Although the average was lower for our algorithm, it has no paths with low bandwidth as the random one. The tail on the left is much shorter while the tail on the right is much longer than the random one. The difference in most paths is negligible.

Finally we evaluate the scenario where only a subset of the generated paths are active, as discussed in section 3. We take the million circuits generated using the privacy-preserving algorithm 7, as above, and choose a random sample of 10 000 circuits as being active. We then compare the bandwidth allocated to those circuits with 10 000 circuits chosen by the random algorithm 2 in fig. 9. Note that the randomized algorithm has a significantly longer tail on the left side of the graph, representing circuits with a pathological bandwidth allocation. The minimum bandwidth in algorithm 2 is 7, whereas the minimum for the sampled algorithm 7 is 171. The randomized algorithm produces 953 flows with a bandwidth less

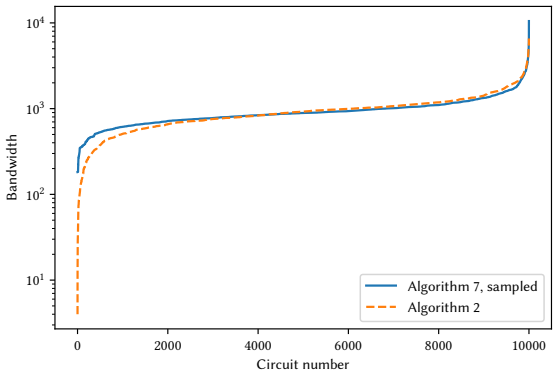


Figure 9: Comparing random allocation in algorithm 2 to a sample of 10 000 circuits out of 1 million that were generated by the differentially private algorithm 7.

than half of the average, 980. The sampled circuits from the private algorithm have a slightly lower average of 960, but only 326 flows have a bandwidth of less than 980/2. The lower average performance is due to some relays being underutilized, as evidenced by the longer tail on the right-hand side of the graph; in our ongoing work we are investigating adjustments to the algorithm to make better use of these relays.

6 RELATED WORK

We next review some closely related work; for more comprehensive of work covering all aspects of Tor performance we direct interested readers to the survey by AlSabah and Goldberg [2].

Snader and Borisov studied the problem of path selection in Tor and suggested biasing selection towards higher bandwidth relays, showing that it improved performance in both simulation and real-world Tor measurements [21, 23]. They also developed a flow simulator based on max-min fair flow allocation, using an algorithm similar to Algorithm 1. Herbert et al. modeled Tor traffic as an M/D/1 queuing network and proposed a relay selection algorithm for optimizing the queuing latency [13]. Both these papers argued that the bandwidth-weighted relay selection results in suboptimal path selection, but their solutions did not incorporate any feedback mechanisms.

Wang et al. proposed a congestion-aware path selection algorithm [26]. It suggests that user perform latency measurements on circuits in the network to detect congested relays and avoid them during path selection. Although each user will have a partial view of the network, their experiments show that significant improvements can be realized. Likewise, Conflux [1] mitigates congested paths by multiplexing traffic across two paths through the Tor network; this significantly improves performance although it cannot mitigate congestion at the last node in the circuits, where traffic must converge. Both these schemes use a partial view of the network to detect and avoid localized congestion, rather than balancing load across the entire Tor network, as in our scheme.

Load balancing in Tor requires an accurate measure of relay capacity, which is currently performed by TorFlow [17]. TorFlow uses bandwidth authorities to proactively measure relay performance, and incorporates a long-term feedback mechanism: relays that are assigned too high a bandwidth value and become overloaded will perform worse in subsequent measurements and thus have their consensus weight reduced, and vice versa. This feedback, however, occurs over a period of days and does not deal with more transient congestion and load imbalances. EigenSpeed [22] proposed an alternate measurement approach that relied on opportunistic measurements of relays by other relays; Johnson et al. identified several attacks on both TorFlow and EigenSpeed and proposed an improved peer measurement scheme called PeerFlow [16].

7 CONCLUSIONS AND FUTURE WORK

We have presented an algorithm that approximates locally optimal load-balancing of circuits in the Tor network while preserving user privacy. We demonstrated that the algorithm significantly improves on the randomized relay assignment in Tor using flow-level simulations of max-min fair bandwidth allocation.

Our promising results encourage the further exploration of using privacy-preserving feedback for load balancing in anonymity networks. Several important challenges remain. First, load imbalance occurs over short time scales, thus the private summary of the network state must be quickly distributed to all users. We note that this is a similar problem to that of distributing blocks in cryptocurrencies; in the Bitcoin network, which is similarly sized to Tor, measurements have shown that blocks reach the median node in 6.5s and the 90th percentile node in 26s [7]. Improving this latency while maintaining resilience to attack is an area of active research.

A second problem is that malicious nodes may misreport their contributions to the histogram to direct circuits away from honest nodes and towards malicious ones. We note that this problem is somewhat similar to the peer bandwidth measurement problem in EigenSpeed [22] and PeerFlow [16] and thus some of the defenses used in those systems may be adaptable to this setting.

Finally, flow-level simulation is a coarse-grained approximation of Tor traffic; web browsing is a dominant use of Tor and web traffic is known to be quite bursty. Further evaluation of the load-balancing mechanism using queuing-based traffic models or full network simulation [14] is needed.

ACKNOWLEDGMENTS

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