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Who started this rumor? Quantifying the natural differential privacy guarantees of gossip protocols

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Abstract: Gossip protocols, also called rumor spreading or epidemic protocols, are widely used to disseminate information in massive peer-to-peer networks. These protocols are often claimed to guarantee privacy because of the uncertainty they introduce on the node that started the dissemination. But is that claim really true? Can one indeed start a gossip and safely hide in the crowd? This paper is the first to study gossip protocols using a rigorous mathematical framework based on differential privacy to determine the extent to which the source of a gossip can be traceable. Considering the case of a complete graph in which a subset of the nodes are curious, we derive matching lower and upper bounds on differential privacy showing that some gossip protocols achieve strong privacy guarantees. Our results further reveal an interesting tension between privacy and dissemination speed: the standard “push” gossip protocol has very weak privacy guarantees, while the optimal guarantees are attained at the cost of a drastic increase in the spreading time. Yet, we show that it is possible to leverage the inherent randomness and partial observability of gossip protocols to achieve both fast dissemination speed and near-optimal privacy.

1 INTRODUCTION

Peer-to-peer networks enable people to share information without the need for any central authority. Some of this information may be sensitive, and people sharing it may not want to be identified, for instance because of copyright infringement when sharing music, or in the case of whistle-blowers. Anonymous sharing platforms can also help people exercise their right to freedom of expression in totalitarian regimes. Conversely, it may be important to locate the source of a (computer or biological) virus, or fake news, spreading in a network. Therefore, it is crucial to understand the fundamental limits on privacy and anonymity in information dissemination.

Gossip protocols (also called *rumor spreading* or *epidemic protocols*), in which nodes *randomly* choose a neighbor to exchange information, are both simple and efficient [1, 8, 30, 42, 50]. They can be used to spread and aggregate information in distributed databases [2, 9, 12, 19, 43, 44] and social networks [20, 35], as well as to optimize cost functions involving distributed datasets in machine learning [15, 21, 37, 38, 51, 59]. A folklore belief is that gossip protocols guarantee some form of *source anonymity* because users cannot know who issued the information in the first place [33]. Although a lot of work has been devoted to assessing how efficiently one could locate the source of a gossip in specific settings [41, 49, 54], the general anonymity claim has never been studied from a pure *privacy* perspective, independently of any particular attack model. Intuitively indeed, random and local exchanges improve privacy, but to what extent? Given the importance of privacy and peer-to-peer information dissemination, it is crucial to study the limitations of this claim through a principled approach. This is the challenge we take up in this paper for the classic case of a complete network graph.

Our first contribution is an information-theoretic model of anonymity in gossip protocols based on an adaptation of (ϵ, δ) -*differential privacy* (DP) [22]. Originally introduced in the database community, DP is a precise mathematical framework recognized as the gold standard for studying the privacy guarantees

of information release protocols. In our proposed model, the information to protect is the source of the gossip, while the attackers are a subset of curious nodes monitoring the communications they receive and their relative order. Our notion of DP then requires that the probability of any possible observation of the attackers is almost the same regardless of which node is the source. A key aspect of our model is that the mechanism that seeks to ensure DP comes only from the *natural* randomness and partial observability of gossip protocols, not from additional perturbation or noise which affects the desired output as generally needed to guarantee DP [24]. We believe our adaptation of DP to be of independent interest. We also complement it with a notion of *prediction uncertainty* which guarantees that even unlikely events do not fully reveal the identity of the source under a uniform prior. This property gives an upper bound on the probability of success of any source prediction attack, including the maximum likelihood estimate.

Based on our proposed model, we then establish *matching upper and lower bounds* on the privacy guarantees of gossip protocols. Essentially, our upper bounds on differential privacy are obtained by tightly lower bounding the probability that the node starting the rumor discloses it to the attacker before another node does and upper bounding the probability that this happens for a random node fixed in advance, in a way that holds for all gossip algorithms. Remarkably, despite the fact that the source node always has a non-negligible probability to send its first message to a curious node, our results highlight the fact that some gossip protocols achieve strong privacy guarantees in several regimes, including the strict $(\epsilon, 0)$ -DP as well as prediction uncertainty. We further show that the upper bounds on privacy are matched by a gossip protocol which has very slow spreading time (*log-linear* in the number of nodes), highlighting an interesting tension between privacy and *dissemination speed*.

To capture this trade-off, we introduce a *parameterized* gossip protocol in which nodes have a fixed probability to stop emitting after each communication (until they receive the message again). This gives the protocol the ability to forget initial conditions, thereby ensuring the privacy of the source. The standard “push” gossip protocol [50], as well as the optimally private but slow protocol we previously introduced, can both be derived from our parameterized scheme with specific choices of the parameter. We show that the standard gossip protocol is inherently not differentially private for arbitrarily large graphs, but that it is possible to devise gossip protocols that are *near-optimally private* with spreading time *logarithmic* in the size of the graph. We prove the protocol speed by analyzing the mean dynamics of gossip and leveraging concentration inequalities. The privacy results are obtained by showing that only a small fraction of the possible outcomes have different probabilities when two different nodes initially have the gossip. This requires to precisely evaluate the probability of well-chosen worst-case sequences, which is generally hard as randomness is involved both when nodes decide to stop sending messages as well as when they choose who to send messages to.

The rest of the paper is organized as follows. We discuss related work in Section 2. In Section 3, we formally introduce our model of differential privacy for gossip protocols. In Section 4, we give matching upper and lower bounds on the privacy guarantees of gossip protocols, and present a privacy-optimal but slow protocol. Section 5 studies how to control the trade-off between speed and privacy. Finally, we conclude in Section 6 by discussing open questions.

2 BACKGROUND AND RELATED WORK

2.1 Gossiping

The idea of disseminating information in a distributed system by having each node *push* messages to a randomly chosen neighbor, initially coined the *random phone-call model*, dates back to even before the democratization of the internet [30, 50]. Such protocols, later called *gossip*, *epidemic* or *rumor spreading*, were for instance applied to ensure the consistency of a replicated database system [2, 19]. They have gained even more importance when argued to model information spreading in social networks [20]. Gossip

protocols can also be used to compute aggregate queries on a database distributed across the nodes of a network [9, 12, 43, 44], and have recently become popular in machine learning for optimizing cost functions involving distributed datasets [15, 21, 37, 38, 51, 59]. Gossip protocols differ according to their interaction schemes, i.e., *pull* or *push*, sometimes combining both [42]. In this work, we focus on the classical *push* form in the standard case of a *complete* graph. The analysis of gossip protocols has also been generalized to more general graphs [10, 14, 25, 34, 53].

2.2 Locating the gossip source

Determining the source of a gossip has been an active research topic, especially given the potential applications to social networks (see [41] for a recent survey). Existing approaches have focused so far on building attacks that to compute or approximate the maximum likelihood estimate of the source given some observed information. Each approach typically assumes a specific kind of graphs (e.g., trees, small world, etc.), dissemination model and observed information. In *rumor centrality* [54–56], the gossip communication graph is assumed to be fully observed and the goal is to determine the *center* of this graph in order to deduce the node that started the gossip. Another line of work studies the setting in which some nodes are *curious sensors* that inform a central entity whenever they receive a message [49]. Gossiping is assumed to happen at random times and the source node is estimated by comparing the different timings at which the information reaches the sensors. The proposed attack is very natural in trees but does not generalize to highly connected graphs in which all nodes are approximately at the same distance. The work of [29] focuses on the problem of hiding the source instead of locating it. The observed information is a snapshot of who has the rumor at a given time. A specific dissemination protocol is proposed to hide the source but the privacy guarantees only hold for tree graphs. Other lines of work focus on hiding the content of the message instead of the source [32], or the initial values of the nodes when trying to compute an aggregate function of nodes individual values (see [17, 18] and references therein), which are orthogonal problems.

We emphasize that the privacy guarantees which can be derived from the above work (i.e., the probability not to be detected) only hold under the specific attacks considered therein. Furthermore, all approaches rely on maximum likelihood and hence assume a uniform prior on the probability of each node to be the source. The guarantees would thus break in case the protocol was run twice from the same source, or if the attacker knew that some of the nodes could not have started the rumor.

2.3 Differential privacy

While we borrow ideas from the approaches mentioned above (e.g., we assume that a subset of nodes are curious sensors as in [49]), our work differs fundamentally for we aim at studying the fundamental privacy limits of any gossip source location protocol, independently of any specific attack, by evaluating the amount of information that is released during a gossip scheme. For this purpose, a general and robust notion of privacy is required. *Differential privacy* [22, 24] has emerged as a gold standard for it holds independently of any assumption on the model, the computational power, or the background knowledge that the attacker may have. Differentially private protocols have been proposed for numerous problems in the fields of databases, data mining and machine learning: examples include computing aggregate and linear counting queries [24, 46], releasing and estimating graph properties [16, 47], clustering [40] and recently deep learning [57].

In this work, we consider the classic relaxed version of differential privacy which involves two parameters $\epsilon, \delta \geq 0$ that quantify the privacy guarantee [23]. More precisely, given any two databases \mathcal{D}_1 and \mathcal{D}_2 that

differ in at most one row (all the rows are the same except for one),¹

$$p(\mathcal{P}(\mathcal{D}_1) \in S) \leq e^\epsilon p(\mathcal{P}(\mathcal{D}_2) \in S) + \delta. \quad (1)$$

Parameter ϵ places a bound on the ratio of the output probabilities when changing one entry of the database, while parameter δ is assumed to be small and allows the bound to be violated with small probability. When $\epsilon = 0$, δ gives a bound on the total variation distance between the output distributions while $\delta = 0$ recovers the strict ϵ -differential privacy. DP guarantees hold regardless of the attack and are robust against strong background knowledge that the attacker may have about the records of the database (in particular, the attacker may know all records in \mathcal{D}_1 and \mathcal{D}_2 except the differing ones). In our context, the background information could be the knowledge that the source is among a subset of k nodes. Robustness against such background knowledge is crucial in some applications, for instance when sharing secret information that few people could have known and leaked in the first place. Another important feature of differential privacy is *composability*: if (ϵ, δ) -differential privacy holds for a release protocol, then querying this protocol two times about the same dataset satisfies $(2\epsilon, 2\delta)$ -differential privacy. This is important in our context for it enables to quantify privacy when the source propagates multiple messages and the adversary is able to link them to the same source (e.g., due to the content of the message).

Existing differentially private protocols typically introduce additional *perturbation* (also called *noise*) to hide critical information [24]. In contrast, an original aspect of our work is to solely rely on the *natural* randomness and limited observability brought by gossip protocols to guarantee differential privacy.

3 A MODEL OF DIFFERENTIAL PRIVACY FOR GOSSIP PROTOCOLS

Our first contribution is a precise mathematical framework for studying the fundamental privacy guarantees of gossip protocols. We define the family of protocols we consider, their inputs and the outputs observed by the attackers during the execution of a protocol, as well as the privacy notions we consider. In the following, we consider a complete graph with n nodes labeled from 0 to $n - 1$.

3.1 Gossip protocols

To specify the class of protocols we consider in this paper, we first define a key communication primitive. Denoting by I the set of informed nodes, $\text{tell_gossip}(i, I)$ allows an informed node $i \in I$ to tell the information to another node $j \in \{0, \dots, n - 1\}$ chosen uniformly at random. $\text{tell_gossip}(i, I)$ returns j (the node that received the message) and the updated I (the new set of informed nodes that includes j). Equipped with this primitive, we can now define gossip protocols as follows.²

Definition 3.1 (Gossip protocols). A gossip protocol on a complete graph is one that (a) terminates, (b) ensures that at the end of its execution, the set of informed nodes $I = \{0, \dots, n - 1\}$, and (c) can modify I only through calls to the tell_gossip primitive.

3.2 Inputs and outputs

As recalled in Section 2.3, differential privacy is a probabilistic notion that evaluates a protocol based on the variations of the *output* distribution for a change in the *input*. In this paper, we adapt it to our gossip context. We first formalize the *inputs* and *outputs*, in the case of a *single piece of information to disseminate* (multiple pieces can be addressed through composition, see Section 2.3). A single node has the information (the gossip, or rumor) at the beginning of the protocol. This node defines the input of the gossip protocol,

¹This is a slight abuse of terminology common in the literature. Strictly speaking, it should be called (ϵ, δ) -*indistinguishability* [23].

²Definition 3.1 is stronger than the one of *address independent protocol* introduced in [42]: it enforces the communication protocol, thus enforcing address independence.

and it is the actual “database” that we want to protect. In this sense, the source node is a database with n rows, each with a binary attribute which is 1 for the source node and 0 elsewhere.

We define the output of a gossip protocol as the information disclosed to an attacker during the execution of the protocol. In this work, we focus on attackers that can monitor a set of *curious nodes* C of size f , i.e. they observe all communications involving a curious node. More formally, a gossip protocol generates an ordered sequence S_{omni} of triplets (t, i, j) of executions of `tell_gossip` where t counts the number of times the `tell_gossip` primitive has been called (starting from $t = 0$), i is the node on which `tell_gossip` was used and j the node that was told the information. This sequence corresponds to the output that would be observed by someone who could eavesdrop on all communications. Since the attacker only monitors a subset of the nodes, through the (random) execution of the protocol it gathers a (random) subsequence $S = ((i, j)|(t, i, j) \in S_{\text{omni}}, j \in C)$. In other words, the attacker only has access to the *relative order* in which the information has been disclosed to curious nodes and is oblivious to `tell_gossip` calls among non-curious nodes. Note that because we focus on complete graphs, knowing which curious node received the rumor gives no information on the starting node. For a given output sequence S , we write $S_t = i$ to denote that the t -th `tell_gossip` in S has been used by node i .

The ratio f/n of curious nodes in the graph determines the probability of the attacker to gather information. Unless otherwise noted, we assume this ratio to be constant. In particular, we see it as a quantity independent of n , otherwise the attacker would only become weaker as the graph grows bigger.

3.3 Privacy definitions

Now that we have precisely defined the inputs and outputs of the “release protocols” that we consider, we formally introduce privacy definitions for the gossip problem. To ease notations, we denote by I_0 the source of the gossip (the set of informed nodes at time 0), and for any given $i \in \{0, \dots, n - 1\}$, we denote by $p_i(E) = p(E|I_0 = i)$ the probability of event E if node i is the source of the gossip. The release protocol is therefore abstracted in this notation. Recalling that \mathcal{S} is the set of all possible outputs of the information release procedure, we say that a gossip protocol is (ϵ, δ) -differentially private if:

$$p_i(S) \leq e^\epsilon p_j(S) + \delta, \quad \forall S \subset \mathcal{S}, \forall i, j \in \{0, \dots, n - 1\}, \quad (2)$$

where $p(S)$ is the probability that the output belongs to the set S . This formalizes a notion of *source indistinguishability* in the sense that, with high probability, any output is almost as likely to be observed by the attackers regardless of who started the gossip (“nothing bad will happen with high probability”). Note however that when $\delta > 0$, this definition allows a protocol to release the identity of the source with small probability. To capture the fact that “nothing too bad will ever happen”, we favor DP protocols that also guarantee the complementary notion of *c-prediction uncertainty*: there exists a constant $c > 0$ such that for a uniform prior $p(I_0)$ on source nodes and any $i \in \{0, \dots, n - 1\}$:

$$p(I_0 \neq \{i\}|S)/p(I_0 = \{i\}|S) \geq c, \quad \forall S \subset \mathcal{S}, p_i(S) > 0. \quad (3)$$

Prediction uncertainty guarantees that no observable output S can identify a node as the source with large enough probability, ensuring that the probability of success of any source prediction attack is upper bounded by $1/(1 + c)$. This holds in particular for the maximum likelihood estimate. Prediction uncertainty does not have the robustness of differential privacy against background knowledge, as it assumes a uniform prior on the source. While it can be shown that $(\epsilon, 0)$ -DP with $\epsilon > 0$ implies prediction uncertainty, the converse is not true. Indeed, prediction uncertainty is satisfied as soon as no output identifies any node with enough probability, without necessarily making all pairs of nodes indistinguishable as in DP. As we show later, prediction uncertainty allows to rule out some naive protocols that have nonzero probability of generating sequences that reveal the source with certainty. Thanks to the symmetry of our problem, we

consider without loss of generality that node 0 starts the rumor ($I_0 = \{0\}$) and therefore verify Equation 2 and Equation 3 only for $i = 0$ and $j = 1$.

REMARK 3.1 (MODEL EXTENSIONS). *We have kept our model relatively simple to avoid unnecessary technicalities in the derivation and presentation of our results. For completeness, we discuss the impact of some possible extensions (e.g., information observed by attackers, malicious behavior, termination criterion) in Appendix A.*

4 OPTIMAL PRIVACY

In this section, we study the fundamental limits of gossip protocols in terms of privacy. Our main result is a set of tight bounds on the privacy guarantees that can be achieved by gossip protocols. We state and discuss these bounds in Section 4.1. We then present the optimally private gossip protocol that matches our bounds and discuss its properties in Section 4.2.

4.1 Main result: matching upper and lower bounds on privacy

We now prove limits on the differential privacy and prediction uncertainty guarantees of any gossip protocol in the sense of Definition 3.1, and show they these bounds are matched by a concrete protocol.

THEOREM 4.1. *If a gossip protocol satisfies (ϵ, δ) -differential privacy for any $\epsilon \geq 0$ and c -prediction uncertainty, then we have $\delta \geq \frac{f}{n} \left(1 - \frac{e^\epsilon - 1}{f}\right)$ and $c \leq \frac{n}{f+1} - 1$. Furthermore, these bounds are tight and matched by Algorithm 1 when its parameter is set to $s = 0$.*

SKETCH OF PROOF. (The complete proof can be found in Appendix B). To prove the lower bounds on the parameter δ of differential privacy, we upper bound the probability that the first node that communicates with a curious node is the source of the rumor. Then, we lower bound the probability that another node fixed in advance communicates with a curious node. We consider the output sequences S such that $S_0 = 0$ (node 0 is the first to communicate with a curious node) and we heavily rely on the fact that all nodes are equally likely to be the first to disclose information to curious nodes after the first message has been sent. The upper bound on prediction uncertainty is obtained by considering the same sequences but summing over all non-curious nodes. To show that Algorithm 1 (with $s = 0$) matches these bounds, we rely on the fact that for this protocol the only outputs that have a different probability if node 0 starts (compared to the case when 1 starts) are those in which 0 (or 1) communicates with a curious node at time 0. We get the desired result by analyzing the probabilities $p_0(S_0 = 0)$ and $p_1(S_0 = 0)$. \square

Theorem 4.1 shows that some gossip protocols are able to provide strong privacy guarantees. In the regime where $\epsilon = 0$ (where DP corresponds to the total variation distance), δ cannot be smaller than the proportion of curious nodes. This is rather intuitive since the source node has probability at least f/n to send its first message to a curious node. Remarkably, however, one can also achieve differential privacy with δ much smaller than f/n by trading-off with $\epsilon > 0$. In particular, the *strict* version of differential privacy ($\delta = 0$) is attained for $\epsilon \approx \log f$, which provides good privacy guarantees when the number of curious nodes is not too large. Furthermore, even though the probability of disclosing *some* information is of order f/n , attackers always have a high probability of making a mistake in their attempt to locate the source, as shown by the prediction uncertainty guarantee.

Crucially, these privacy guarantees are made possible by the *natural* randomness and partial observability of gossip protocols. We show in Appendix C that for a stronger attacker with access to the global timing of messages instead of their relative order, the guarantee in total variation distance still holds but it becomes impossible to achieve prediction uncertainty and strict differential privacy.

4.2 Optimally private protocol

We now discuss the gossip protocol that matches the bounds of Theorem 4.1, which turns out to be very simple: nodes forward the message to exactly one random neighbor when they receive it and then stop emitting until they receive the message again. This protocol, which we refer to as *private gossip*, corresponds to a special case of the more general protocol described in Algorithm 1 when its parameter is $s = 0$ (see discussion in Section 5). Private gossip is similar to the protocol introduced by [29] in the sense that at each time step, the source changes and it is quickly impossible to recover which node started the gossip (as initial conditions are quickly forgotten). Private gossip ensures that (i) the gossip does not die before all nodes are informed, (ii) the state of the system (the set informed nodes I) after the first message is sent is completely independent from the source node, and (iii) all nodes follow the same behavior. The first property ensures that the protocol falls within Definition 3.1, the second one is key to match the optimal privacy parameters stated in Theorem 4.1 (see Appendix B.2 for the proof), and the third one prevents the source to be identified due to a special behavior, in contrast to the naive alternative described below.

REMARK 4.1 (DELAYED START PROTOCOL). *A naive alternative to the private gossip is as follows: the source node transmits the rumor to a random node and forgets it, then a fast gossip protocol (such as Algorithm 1 with $s = 1$) may start normally from the node that received the information. While this delayed start protocol achieves optimal differential privacy in some regimes, it is fundamentally flawed. In particular, it does not guarantee prediction uncertainty in the sense that $c \rightarrow 0$ as the size of the graph increases. This is because attackers can identify the source with high probability by detecting that it communicated only once and then stopped emitting for many rounds. We refer to Appendix E for details.*

An obvious drawback of the private gossip protocol is that it is very slow in practice, since only one node can transmit the rumor at any given time. To precisely quantify its dissemination speed, one can observe that it performs a random walk on the complete graph. Therefore, the number of gossip operations needed to inform all nodes can be reduced to the time needed for the classical coupon collection problem: it takes $O(n \log n)$ communications to inform all nodes with probability at least $1 - 1/n$ [26]. As the private gossip protocol performs exactly one communication at any given time, it needs time $O(n \log n)$ to inform all nodes with high probability. This is much slower than the standard “push” gossip protocol in which all informed nodes can transmit the rumor in parallel, requiring only time $O(\log n)$ [30]. This motivates the exploration of the privacy-speed trade-off. In the next section, we introduce gossip protocols with $O(\log n)$ speed and nearly optimal privacy.

5 FASTER PRIVATE GOSSIP PROTOCOLS

In this section, we study faster variants of the private gossip protocol in which nodes do not necessarily stop emitting after they first transmit the information. Algorithm 1 describes a class of gossip protocols parameterized by $s \in [0, 1]$, which fits Definition 3.1. They maintain a set A of *active nodes* (initialized to the source node) which spread the rumor asynchronously and in parallel: this is modeled by the fact that at each step of the algorithm, a randomly selected node $i \in A$ invokes the `tell_gossip` primitive to send the information to another node (which in turn becomes active), while i also stays active with probability s . This protocol, illustrated in Figure 1, can be understood as a gossip protocol with a randomized version of *fanout* [28].³ It is also related to the SIS epidemic model [39] but here the rumor never dies regardless of the value of $s \in [0, 1]$ (i.e., there always remain some active nodes). Parameter s controls the trade-off between privacy and speed: in particular, $s = 0$ recovers the private gossip protocol (optimal privacy) and $s = 1$ recovers the standard “push” gossip protocol (optimal speed).

³Unlike in classic fanout, nodes start to gossip again each time they receive a message instead of deactivating permanently.

Algorithm 1 Fast Private Gossip

Require: n {Number of nodes}, k {Source node}, s {Probability for a node to remain active}

Ensure: $I = \{0, \dots, n-1\}$ {All nodes are informed}

- 1: $I \leftarrow \{k\}, A \leftarrow \{k\}$
 - 2: **while** $|I| < n$ **do**
 - 3: Sample i uniformly at random from A
 - 4: $j, I \leftarrow \text{tell_gossip}(i, I), A \leftarrow A \cup \{j\}$
 - 5: $A \leftarrow A \setminus \{i\}$ with probability $1 - s$
 - 6: **end while**
-

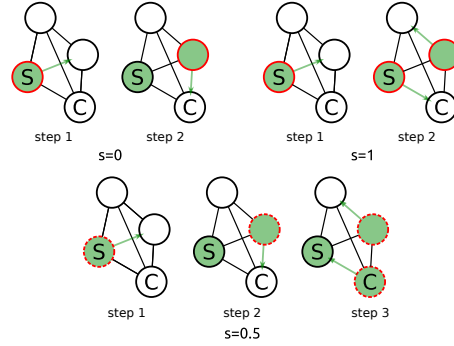


Fig. 1. *Left:* Fast Private Gossip protocol. *Right:* Illustration of the role of parameter s . S indicates the source and C a curious node. Green nodes know the rumor, and red circled nodes are active. When $s = 0$, there is only one active node at a time, which always stops emitting after telling the gossip. The resulting protocol (private gossip) is private but slow. In the case $s = 1$, nodes always remain active once they know the rumor, leading to a fast but non private protocol (this is the standard push gossip protocol [50]). When $0 < s < 1$, each node remains active with probability s after each communication, providing a trade-off between privacy and speed.

In the rest of this section, we study how parameter s of Algorithm 1 impacts its privacy guarantees and its dissemination speed. Section 5.1 establishes that the privacy guarantees of the standard gossip protocol ($s = 1$) must be arbitrarily bad for large graphs. Then, we show in Section 5.2 that nearly optimal privacy can be achieved for smaller s . Finally, Section 5.3 studies the dissemination speed and shows that the known logarithmic diffusion time of the standard gossip protocol also holds for $s > 0$, leading to a sweet spot in the privacy-speed trade-off.

5.1 The standard “push” protocol is not differentially private

Section 4 hints at the fact that gossip protocols need to forget initial conditions quickly in order to be private. In this section, we strengthen this intuition by showing that the differential privacy guarantees of the standard “push” gossip protocol (corresponding to Algorithm 1 with $s = 1$) become arbitrarily bad as the size of the graph increases (keeping the fraction of curious nodes constant).

THEOREM 5.1. *If Algorithm 1 with $s = 1$ guarantees (ϵ, δ) -differential privacy for all values of n and constant $\epsilon < \infty$, then $\delta = 1$.*

The proof of this result can be found in Appendix D. Essentially, it comes from the fact that the event “node 0 communicates with a curious node before node 1 gets the message” becomes more and more likely as n grows, hence preventing any meaningful differential privacy guarantee when n is large enough. This motivates our interest for gossip protocols with parameter $s < 1$.

5.2 Privacy guarantees of the fast private gossip protocol

The previous section clearly highlights the fact that the standard gossip protocol ($s = 1$) is not differentially private. We now show that giving nodes the possibility to stop emitting by setting $s < 1$ is enough for the protocol to have non-trivial privacy guarantees.

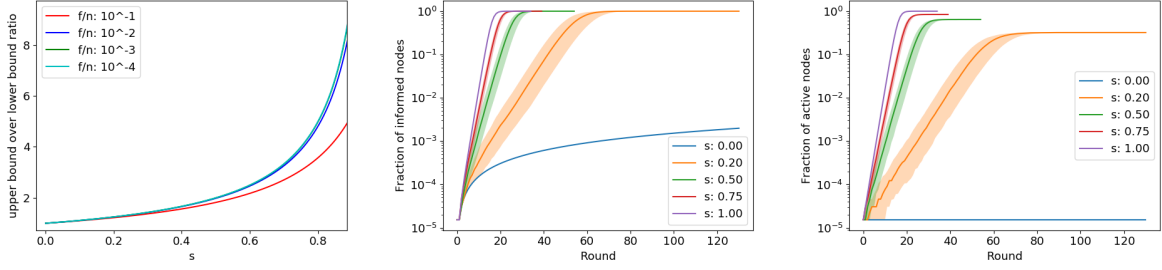


Fig. 2. Effect of parameter s of Algorithm 1 on the privacy guarantees and spreading speed. *Left*: Ratio between the upper bound and lower bound on the differential privacy parameter δ . *Middle, Right*: Fractions of informed and active nodes along rounds on a network of $n = 2^{16}$ nodes. The lines stop when all nodes are informed (and so the protocol terminates), except for $s = 0$ since the protocol is very slow in this case. The curves represent median values and the shaded area represents the 10 and 90 percent confidence intervals computed over 100 runs.

THEOREM 5.2. *For $0 < s < 1$ and any fixed $r \in \mathbb{N}^*$, Algorithm 1 with parameter s guarantees $(0, \delta)$ -differential privacy with:*

$$\delta = 1 - (1 - s) \sum_{k=0}^{\infty} s^k \left(1 - \frac{f}{n}\right)^{k+1} \leq 1 - (1 - s^r) \left(1 - \frac{f}{n}\right)^r. \quad (4)$$

PROOF. See Appendix F.1. □

Theorem 5.2 proves a $(0, \delta)$ -differential privacy result, which means that apart from some unlikely outputs that may disclose the identity of the source node, most of them actually have the same probability regardless of which node started the diffusion. The guarantee we obtain here holds for any graph with fixed proportion f/n of curious nodes. Figure 2 (left) shows the gap between the differential privacy guarantees given by Theorem 5.2 and the optimal guarantees of Theorem 4.1 (i.e., the ratio between the upper bound and lower bound on δ). We see that both bounds are of the same order of magnitude when s is not too large: in particular, the ratio is less than 2 for all $s \leq 0.5$. This indicates that the privacy guarantees are very tight in this regime. Note that setting $r = 1$ in Theorem 5.2 leads to an additive gap of $s(1 - f/n)$ between the privacy of Algorithm 1 and the optimal guarantee, showing that one can be as close as desired to the optimal privacy as long as s is chosen close enough to 0. We also recover exactly the optimal guarantee of Theorem 4.1 in the case $s = 0$ (without the ability to control the trade-off between ϵ and δ).

Importantly, we also prove (see Appendix F.2) that Algorithm 1 with $s < 1$ satisfies prediction uncertainty, unlike the case where $s = 1$.

THEOREM 5.3. *Algorithm 1 guarantees prediction uncertainty with $c = (1 - \frac{f+1}{n})(1 - s)$.*

5.3 Dissemination speed

We have shown in the previous section that parameter s has a significant impact on privacy, from optimal ($s = 0$) to very weak ($s = 1$) guarantees. Intuitively, s also impacts the dissemination speed: the larger s , the more active nodes at each round. This is highlighted by the two extreme cases, for which the speed is already known and exhibits a large gap: $O(\log n)$ for $s = 1$ [30] while it is only $O(n \log n)$ when $s = 0$. To establish whether we can obtain a protocol that is both private and fast, we need to characterize the dissemination speed for the cases where $0 < s < 1$.

The key result of this section is to prove that the logarithmic speed of the standard “push” gossip protocol holds more generally for all $s > 0$. This result is derived from the fact that the ability to forget

does not prevent an *exponential growth* phase. What changes is that the population of active nodes takes approximately $1/s$ rounds to double instead of 1 for standard gossip. For ease of presentation, we state below the result for the synchronous version of Algorithm 1, in which the notion of *round* corresponds to iterating over the full set A . A similar result (with an appropriate notion of rounds) can be obtained for the asynchronous version given in Algorithm 1 (see Appendix G.2 for details).

THEOREM 5.4. *For a given $s > 0$, there exists $\alpha > 0$ such that for all $C > 0$, there exists n large enough such that the synchronous version of Algorithm 1 with parameter s sends at least $Cn \log n$ messages in $C\alpha^{-1} \log n$ rounds with probability at least $1 - 1/n$.*

SKETCH OF PROOF. The key argument of the proof is that the gossip process very closely follows its mean dynamics. After a transition phase of a logarithmic number of rounds, a constant fraction of the nodes (depending on s) remains active despite the probability to stop emitting after each communication. This “determinism of gossip process” has been introduced in [52], but their analysis only deals with the case $s = 1$. Our proof takes into account the nontrivial impact of nodes deactivation in the exponential and linear growth phase. Besides, we introduce and analyze a last phase, showing that with high probability the population never drops below a critical threshold of active nodes. The details can be found in Appendix G.1. \square

This result shows that our parameterized gossip protocol with $s > 0$ still has a logarithmic spreading time even if nodes can stop transmitting the message. Note that the constant α depends on s and will go to infinity as $s \rightarrow 0$ because $1/s$ rounds are needed in expectation to double the population of active nodes (even without taking collisions into account). Simulations shown in Figure 2 (right) confirm that the fraction of active nodes grows exponentially fast for all values of s and then reaches a plateau when the probability of creating a new active node is compensated by the probability of message collisions (informing an already active node). Empirically, this happens when the fraction of active nodes is of order s , meaning that the last phase (during which the remaining uninformed nodes need to be reached by a stable number of $s \times n$ active nodes) remains short. This highlights the fact that Algorithm 1 remains significantly faster than the slow private gossip: for instance, dissemination speed for $s = 0.5$ is very close to the fastest case $s = 1$ (see Figure 2, middle). We note incidentally that gossip protocols are often praised for their robustness [3, 31]. While the slow private gossip protocol does not tolerate a single message failure, fast private gossip with $s > 0$ is much more resilient thanks to the linear proportion of active nodes, making it very unlikely that the protocol stops because of lost messages as long as s is larger than the probability of losing messages. To summarize, we have shown that one can achieve both fast spreading and near-optimal privacy.

6 CONCLUDING REMARKS

This paper initiates the formal study of privacy in gossip protocols to determine the extent to which the source of a gossip can be traceable. Our contributions are the following. (1) We proposed a formal model of anonymity in gossip protocols based on an adaptation of differential privacy. (2) We established tight upper bounds on the privacy of gossip protocols, highlighting their natural privacy guarantees. (3) We precisely captured the trade-off between privacy and speed with a parameterized gossip protocol allowing nodes to stop gossiping after some time, showing that it is possible to design gossip protocols that are both fast and near-optimally private.

Our work opens several interesting perspectives. In particular, it paves the way to the study of differential privacy guarantees for gossip protocols in *general graphs*, which is a very challenging problem that requires to introduce some relaxations of differential privacy in order to obtain nontrivial guarantees. We refer to Appendix H for a brief discussion of these open questions.

Another exciting avenue for future research is motivated by some recent work showing that hiding the source of a message can significantly amplify differential privacy guarantees for the *content* of the message

[6, 13, 27]. However, primitives to hide the source of messages such as onion routing [58] can be difficult and costly to deploy. Showing that gossiping can *naturally* amplify differential privacy for the message contents would make gossip protocols very desirable for privacy-friendly distributed applications and privacy-preserving decentralized machine learning [7].

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A MODEL EXTENSIONS

We have kept our model of Section 3 relatively simple to avoid unnecessary complexity in the notations and additional technicalities in the derivation and presentation of our results. In this section, we briefly discuss some possible extensions. Our main point is to illustrate the fact that they will generally lead to some technical complications without impacting the privacy guarantees significantly.

A.1 Pull and push-pull protocols

Our study focuses on the classic *push* form of gossip protocols. This can be justified by the fact that, for regular graphs, synchronous push has asymptotic spreading time guarantees that are comparable with the push-pull variant [36]. Besides, the differential privacy guarantees of any gossip protocol are limited by the probability that the first node informed by the source is a curious node, and we show this bound can be matched with push protocols. Nevertheless, extensions of our results to pull and push-pull variants of gossip protocols [42] are possible. Forgetting mechanisms similar to the ones in Algorithm 1 can be introduced for these protocols, i.e. nodes would have a probability $1 - s$ to stop disclosing information after each time they are pulled (if they do not pull someone with the information in between). Although slightly different, the optimal privacy guarantees would remain of the same order of magnitude. Yet, we expect pull guarantees to be much worse in the case $s = 1$ because curious nodes could stop suspecting all nodes that they have pulled and that did not have the rumor. Besides, the pull protocol for $s = 0$ would be even slower than its push counterpart.

A.2 Eavesdropping adversary

Since we consider a complete graph, our formalization of the attackers as a fraction f/n of curious nodes is closely related to an eavesdropping attacker who would observe each communication with probability f/n . Indeed, both models consider that each communication has a probability f/n of being disclosed to the attacker. Most of our results are thus easily transferable to this alternative setting. The only difference would be that all nodes can be suspected in the eavesdropping model, thus introducing a $(1 - f/n)^{-1}$ factor each time we consider the population of non-curious nodes.

A.3 Information observed by attackers

We discuss two possible generalizations of the output observed by the attackers.

Messages sent by curious nodes. For simplicity of exposition, we consider that curious nodes only observe messages that are sent to them and not the messages that they send. However, including the messages sent by curious nodes in their observed output would not impact the bounds on privacy (i.e., the guarantees for the algorithms). For the optimal algorithm of Section 4.2, we only consider what happens during the first round, so including the messages sent by curious nodes does not change the result. This in particular implies that the fundamental limits of Theorem 4.1 remain the same (since the attackers observe strictly more information). Similarly, for the parameterized algorithm of Section 5, Theorem 5.2 is obtained by bounding the probability of a set \hat{S} . Then, we have $p(\hat{S}, S_{\text{out}}) \leq p(\hat{S})$ where S_{out} is the sequence of messages sent by the curious nodes. In general, adding the messages sent by curious nodes to the output sequences has little or no impact on the results.

Message ordering. We assume that the relative order of messages is preserved in the output sequence observed by curious nodes. This could be relaxed, as in practical scenarios a message sent before another may well be received after it. One could for instance introduce a random swapping model to take this into account and investigate whether this weaker output leads to an improvement in the privacy guarantees.

However, we argue that this improvement would be quite limited. First of all, it would not affect the privacy guarantees of Section 4: since there is a single active node able to send a message at any given time, swapping is not possible. Furthermore, even when several nodes are active at the same time (e.g., in Algorithm 1 with $s > 0$), the proofs can be adapted to work with counting the messages *received* instead of the messages *sent*. In this case, swapping is as likely to expose the source (making its messages arrive earlier) than to hide it (delaying the messages it sends). Therefore, privacy would not improve substantially.

A.4 Malicious behavior

In this work, we have assumed for simplicity that nodes are *curious* but not *malicious*, i.e., they follow the protocol. This is motivated by a practical scenario where a subset of nodes are simply being monitored by a curious entity. If curious nodes can also act maliciously, they have three possible ways to affect the protocol: emitting more, emitting less, or not choosing neighbours uniformly at random. If they emit more, they will inform more nodes, which makes it more difficult for them to locate the source. If they emit less (potentially not at all), then in the case $s < 1$ the protocol could stop before all nodes are informed. Yet, the privacy bounds are derived from the fact that the source forgets the information before communicating to a curious node. If they choose the neighbors they send the messages to, it reduces to the case in which they emit less (because they do not send messages to uninformed nodes) but without affecting protocol speed or termination (because it does not reduce the number of active nodes). Thus, the impact on the observed output and therefore on the privacy would be minimal. In the case $s = 1$, malicious nodes have slightly more impact but remain quite small as it only makes the set of informed nodes grow slightly slower.

A.5 Termination criterion

For simplicity, in all our gossip protocols we have used a global termination criterion (the protocol terminates when all nodes are informed). Termination without using global coordination is a problem in its own right that has been extensively studied (see for instance [42]). Although some termination criteria could have a great impact on privacy, we argue that termination can be handled late in the execution so as to reveal very little about the beginning, hence avoiding any significant impact on privacy. For instance, it is possible to design a variant of Algorithm 1 in which nodes only flip a coin with probability s for a fixed number of times, and then stop emitting completely. This fixed number would have to depend on s , but then if it is large enough, it would guarantee both termination and privacy. Indeed, nodes would not communicate with curious nodes each time they are activated with high probability so this counter would actually provide very little information to the curious nodes. Determining how large this number of iterations should be, and the exact impact on privacy (which we argue is very small), is beyond the scope of this paper.

B PROOFS OF THE OPTIMAL PRIVACY RESULTS

In this section, we prove Theorem 4.1. To do so, we start by deriving the lower bounds on the privacy parameters and then show that they are matched by the private gossip protocol.

B.1 Lower bounds

To this end, we introduce a useful technical fact which directly follows from the definition of differential privacy and is at the heart of our lower bound proofs. Fact 1 means that proving a lower bound on the differential privacy parameters can be achieved by finding a set of possible outputs S (here, a set of ordered sequences) that is more likely if node 0 starts the gossip than if node 1 does. It is a direct application of the definition of differential privacy.

FACT 1. Given any gossip protocol, let $S \subset \mathcal{S}$ and $w_0, w_1 \in \mathbb{R}$ be such that $w_0 \leq p_0(S)$ and $p_1(S) \leq w_1$. If the protocol satisfies (ϵ, δ) differential privacy then $\delta \geq w_0 - e^\epsilon w_1$.

We then prove the lower bound part of Theorem 4.1:

PROOF OF THEOREM 4.1 (LOWER BOUND PART). Since `tell_gossip` requires that the input node i is in I and that at the beginning, $I = \{0\}$, the first time the procedure is called must be on node 0. The procedure is called at least once otherwise the protocol terminates with $I = \{0\}$, violating the conditions of Definition 3.1. We denote by $S^{(0)}$ the set of output sequences such that $S_0 = 0$ (i.e., 0 is the first to communicate with a curious node). We also denote by T_0^c the event such that the starting node does not communicate with a curious node for its first communication. Conditionally upon T_0^c , the node that started the gossip is at least as likely as any other node to emit the second message, because with probability $\frac{1}{n}$ it is the only node with the rumor after the first message is sent. Since the probability to hit a curious node is the same regardless of who sends the message, we have for all $i, j \notin C$: $p_j(S_0 = i|T_0^c) \leq p_j(S_0 = j|T_0^c)$. From this inequality we get

$$\sum_{i \notin C} p_0(S_0 = 0|T_0^c) \geq \sum_{i \notin C} p_0(S_0 = i|T_0^c) = 1 \geq \sum_{i \notin C} p_0(S_0 = 1|T_0^c),$$

where the equality comes from the fact that S_0 is the first node that communicates with a curious node (and the curious nodes do not start with the information). The second inequality comes from the fact that $p_j(S_0 = i|T_0^c) = p_j(S_0 = k|T_0^c)$ for all i, k different from j . Therefore, we have $p_0(S_0 = 0|T_0^c) \geq \frac{1}{n-f}$ and $p_0(S_0 = 1|T_0^c) \leq \frac{1}{n-f}$.

Combining the above expressions, we derive the probability of $S^{(0)}$ when 0 started the diffusion:

$$\begin{aligned} p_0(S^{(0)}) &= p_0(S^{(0)}, t_c = 0) + p_0(S^{(0)}, T_0^c) = p_0(t_c = 0) p_0(S^{(0)}|t_c = 0) + p_0(S^{(0)}|T_0^c) p_0(T_0^c) \\ &\geq \frac{f}{n} \times 1 + \frac{1}{n-f} \left(1 - \frac{f}{n}\right) = \frac{f}{n} + \frac{1}{n}. \end{aligned}$$

We then do the same split if node 1 initially has the message, but in this case $p_1(S^{(0)}|t_c = 0) = 0$ and we get $p_1(S^{(0)}) = p_1(T_0^c) p_1(S^{(0)}|T_0^c) \leq \frac{1}{n}$. We conclude again by Fact 1.

The upper bound on prediction uncertainty is derived using the same quantities. More precisely:

$$\frac{p(I_0 \neq 0|S^{(0)})}{p(I_0 = 0|S^{(0)})} = \sum_{i \notin C \cup \{0\}} \frac{p_i(S^{(0)})}{p_0(S^{(0)})} \leq (n-f-1) \frac{p_1(S^{(0)})}{p_0(S^{(0)})} \leq (n-f-1) \frac{1}{f+1} = \frac{n}{f+1} - 1 \quad (5)$$

□

B.2 Optimal protocol

We now proceed to proving the second part of Theorem 4.1, i.e. that these bounds are matched by the private gossip protocol.

PROOF OF THEOREM 4.1 (MATCHING PRIVACY GUARANTEES). For this protocol, the only outputs that have a different probability if node 0 starts (compared to the case when 1 starts) are those in which 0 (or 1) communicates with a curious node at time 0. Following our previous notations, we write these two events $S_0 = 0$ and $S_0 = 1$ and further denote by $S^{(0)}$ (resp. $S^{(1)}$) the set of output sequences such that $S_0 = 0$ (resp. $S_0 = 1$).

We denote by T_0 the event such that node 0 communicates with a curious node (and T_0^c the negation of this event). We have:

$$p_0(S_0 = 0) = p(T_0) p_0(S_0 = 0|T_0) + p(T_0^c) p_0(S_0 = 0|T_0^c). \quad (6)$$

For any $i \notin C$ where C is the set of curious nodes, we have that $p_0(S_0 = 0|T_0^c) = p_0(S_0 = i|T_0^c) = \frac{1}{n-f}$. Indeed, given T_0^c , the node that received the first message was selected uniformly at random among non-curious nodes, and has the same probability to disclose the gossip at future rounds. Plugging into (6), we obtain:

$$p_0(S_0 = 0) = \frac{f}{n} + \left(1 - \frac{f}{n}\right) \frac{1}{n-f} = \frac{f+1}{n}.$$

For any other node $i \neq 0$, $p_0(S_0 = i) = p_0(T_0^c)p_0(S_0 = i|T_0^c) = \frac{1}{n}$ because $p_0(S_0 = i|T_0) = 0$. Combining these results we get $p_0(S^{(0)}) \leq e^\epsilon p_1(S^{(0)}) + \delta$ for any $\epsilon > 0$ and $\delta = \frac{f}{n}(1 - \frac{e^\epsilon - 1}{f})$. By symmetry, we make a similar derivation for $S^{(1)}$, which concludes the proof.

To prove the prediction uncertainty result, we use the differential privacy result with $e^\epsilon = f + 1$ (and thus $\delta = 0$ and write that for any $S \in \mathcal{S}$:

$$\frac{p(I_0 \neq 0|S)}{p(I_0 = 0|S)} = \sum_{i \notin CU\{0\}} \frac{p_i(S)}{p_0(S)} \geq (n-f-1)e^{-\epsilon} = \frac{n}{f+1} - 1 \quad (7)$$

□

C PRIVACY GUARANTEES IN THE CASE OF A STRONG ATTACKER

In our model, we assume that curious nodes only have access to the relative ordering in which they received the messages but they have no information on the global time at which it was sent. We briefly discuss here how the privacy guarantees are affected if one considers a stronger attacker that has access to the number of times the `tell_gossip` procedure has been called. Formally, this attacker observes the sequence $S = ((t, i, j) \in S_{\text{omni}} | j \in C)$. Note that this is not a realistic attacker as gossip protocols naturally enforce partial observability of the events, but the point is to study the influence of the timing information on the privacy guarantees.

The following result quantifies the limits of privacy for this stronger adversary, which can be compared to the results of Theorem 4.1 in the main text. We can see that in the regime $\epsilon = 0$ (total variation distance), the limits remain the same. However, achieving $\delta < f/n$ and prediction uncertainty is not possible against this stronger adversary.

THEOREM C.1. *If a gossip protocol satisfies (ϵ, δ) -differential privacy and c -prediction uncertainty then we have $\delta \geq \frac{f}{n}$ and $c = 0$ in the strong attacker setting. Furthermore, these bounds are tight and matched by Algorithm 1 when its parameter is set to $s = 0$.*

PROOF. By the same reasoning as before, `tell_gossip` is called at least once and is first called on node 0. Sequence $S^{(0)}$ now denotes the fact that node 0 communicates with a curious node at time 0. Since the protocol is run on the complete graph, the node selected by `tell_gossip` is chosen uniformly within $\{0, \dots, n-1\}$, so a curious node is selected with probability $\frac{f}{n}$. We thus have $p_0(S^{(0)}) = \frac{f}{n}$. Besides, node 0 cannot communicate with a curious node at time 0 if node 1 starts the rumor so $p_1(S^{(0)}) = 0$. We conclude by Fact 1. For prediction uncertainty, using the same sequence $S^{(0)}$ yields $\frac{p_i(S^{(0)})}{p_0(S^{(0)})} = 0$ for all $i \neq 0$ and therefore $c = 0$.

It remains to show that these bounds are matched by the same private gossip protocol. The fact that the only outputs that have a different probability if node 0 starts (compared to the case when 1 starts) are those in which 0 (or 1) communicates with a curious node for its first communication is still true with the stronger attacker. Then, we write $p_0(S_0 = 0) = p_1(S_0 = 1) = \frac{f}{n}$ and $p_0(S_0 = 1) = p_1(S_0 = 0) = 0$. This ensures that $p_0(S^{(0)}) \leq p_1(S^{(0)}) + \frac{f}{n}$ (similarly for $S^{(1)}$), and the result follows. □

D LIMITS ON THE DIFFERENTIAL PRIVACY OF STANDARD GOSSIP

In this section, we prove that Algorithm 1 with $s = 1$ cannot achieve meaningful privacy guarantees.

PROOF OF THEOREM 5.1. The proof consists in studying $S_r^{(0)}$, the set of output sequences such that the rank of node 0 in the sequence is less than r . For a specific sequence to not be in $S_r^{(0)}$, there must have been at least r communications (because r nodes must have communicated with curious nodes), and none of them involved 0 and a curious node. Therefore, if we note $n_c(r)$ the number of communications that actually happened before the output sequence reached size r , we have $n_c(r) \geq r$. Then, denoting by $C(t)$ the node that communicated with a curious node at time t (with $C(t) = -1$ when the communication did not involve a curious node):

$$p_0(S_r^{(0)}) = 1 - p\left(\bigcap_{t=0}^{n_c(r)} C(t) \neq 0\right) = 1 - \prod_{t=0}^{n_c(r)} p(C(t) \neq 0) \geq 1 - \prod_{t=0}^r (1 - p(C(t) = 0)) \geq 1 - \prod_{t=0}^r \left(1 - \frac{f}{n} \frac{1}{t}\right),$$

where the last step comes from the fact that the probability of node 0 to be selected at time t is $\frac{1}{|I_t|} \geq \frac{1}{t}$ because at most one node is informed at each step and the active node is selected uniformly among nodes that have the information. We use the fact that $\log(1+x) \leq x$ for any $x > -1$ on $x = -\frac{f}{n} \frac{1}{t}$ to show that:

$$\prod_{t=0}^r \left(1 - \frac{f}{n} \frac{1}{t}\right) = e^{\sum_{t=0}^r \log\left(1 - \frac{f}{n} \frac{1}{t}\right)} \leq e^{-\frac{f}{n} \sum_{t=0}^r \frac{1}{t}}. \quad (8)$$

Therefore, $p_0(S_r^{(0)})$ goes to 1 as r goes to infinity.

Then, for a given r and for any $k > 0$, $p(n_c(r) \leq k)$ is equal to $p(\text{Binom}(k, \frac{f}{n}) \geq r)$ where $\text{Binom}(k, \frac{f}{n})$ is the binomial law of parameters k and $\frac{f}{n}$. This is because it is the probability of having exactly r successes with the sum of less than k Bernoullis of parameter $\frac{f}{n}$, which is equal to the probability of having more than r successes with the sum of k Bernoullis of the same parameters. Therefore, $p(n_c(r) \leq k)$ is independent of n and we can choose k^* independently of n such that $p(n_c(r) > k^*) \leq \frac{1}{n}$. Then, we write that

$$p_1(S_r^{(0)}) = p_0(S_r^{(0)}, n_c(r) \leq k^*) + p_0(S_r^{(0)}, n_c(r) > k^*) \leq p_0(S_r^{(0)} | n_c(r) \leq k^*) + \frac{1}{n}.$$

This implies $p_0(S_r^{(0)} | n_c(r) \leq k^*) \leq p_0(0 \in I_r | n_c(r) \leq k^*) \leq 1 - p(0 \notin I_r | n_c(r) \leq k^*)$. We know that only r communications have reached curious nodes but the others have reached a random node in the graph, and there is at most k^* of them, so finally:

$$p_1(S_r^{(0)}) \leq 1 - \left(1 - \frac{1}{n}\right)^{k^*} + \frac{1}{n}.$$

We immediately see that $p_1(S_r^{(0)})$ goes to 0 as n grows because k^* is independent of n , and we have shown above that $p_0(S_r^{(0)})$ goes to 1 as n grows. Since we must have that $p_0(S_r^{(0)}) \leq e^\epsilon p_1(S_r^{(0)}) + \delta$, we must have $\delta = 1$ if we want δ and ϵ to be independent of n . \square

E DELAYED START GOSSIP

Consider the protocol described in Remark 4.1, that we call *delayed start gossip*:

1. The source calls `tell_gossip` once to transmit the rumor to an arbitrary node, say node j .
2. Node j then starts a standard gossip (Algorithm 1 with $s = 1$).

This simple protocol is clearly optimal from the point of view of differential privacy in the regime $\epsilon = 0$ (total variation distance) because if the first communication does not hit a curious node then the probability

of a given output when two different nodes start the gossip is the same. It is also fast since it runs the standard gossip after the first round.

Yet, this naive protocol has a major flaw. Indeed, when the first communication hits a curious node, the attackers can monitor whether the sender communicates with them again in the next rounds. If it does not, they can guess that the node is the source, and they will in fact make a correct guess with probability arbitrarily close to 1 for large enough graphs. On the other hand, when the sender communicates again with a curious node shortly after, they can be very confident that this node is not the source. Hence, it is possible to design a very simple attack with a very high precision (almost always right) and almost optimal recall (finds the source every time the information is actually released, i.e. with probability $\frac{f}{n}$).

Making sure that the attacker is uncertain about its prediction is therefore a desirable property. This is captured by the notion of *prediction uncertainty* that we introduced in Section 3.3. The following proposition formalizes the above claims and motivates the need for more involved protocols such as the faster private gossip protocol presented in Section 5.

PROPOSITION E.1. *We call c_{ds} the prediction uncertainty constant of the delayed start protocol and we assume the ratio of curious nodes f/n to be constant. Then $c_{ds} \rightarrow 0$ when $n \rightarrow \infty$.*

More generally, it is in principle possible to prove similar results for any protocol in which the source node does not behave like other nodes. Indeed, if the special behaviour can be detected, then attackers can know for sure the source of the rumor.

PROOF OF PROPOSITION E.1. The proof reuses some elements of the proof of Theorem 5.1 (see Appendix D). We consider the sequence $S_r^{(0)}$ such that node 0 is the first node to communicate with a curious node ($S_0 = 0$) and then r other nodes communicate with curious nodes before 0 does ($S_i \neq 0$ for $i \in \{1, \dots, r\}$). We denote by t_0 the time at which node 0 gets the message and becomes active again (we refer here to the global order, although of course the curious nodes do not have access to it). Then, with the usual notations we have:

$$\begin{aligned} p_0 \left(S_r^{(0)} \right) &= p_0(S_0 = 0) p_0 \left(S_r^{(0)} | S_0 = 0 \right) \geq \frac{f}{n} p_0 \left(\bigcap_{i=1}^r S_i \neq 0 | S_0 = 0 \right) \geq \frac{f}{n} p_0(t_0 \geq r) \\ &\geq \frac{f}{n} p_0(n_c(r) \leq k^*) p_0(t_0 \geq r | n_c(r) \leq k^*). \end{aligned}$$

Then, we recall from the proof of Theorem 5.1 (Appendix D) that

$$p_0(n_c(r) \leq k) = p \left(\text{Binom}(k, \frac{f}{n}) \geq r \right) = p \left(\text{Binom}(k, 1 - \frac{f}{n}) < k - r \right) = 1 - p \left(\text{Binom}(k, 1 - \frac{f}{n}) \geq k - r \right),$$

so if we set $k = \frac{2n}{f}r$ and use tail bounds on the binomial law (Theorem 1 of [5]) then there exists a constant H (only depending on $\frac{f}{n}$) such that:

$$p_0(n_c(r) \leq r \frac{2n}{f}) \geq 1 - e^{-rH}.$$

Therefore, we have:

$$p_0 \left(S_r^{(0)} \right) \geq \frac{f}{n} \left(1 - e^{-rH} \right) \left(1 - \frac{1}{n} \right)^{r \frac{2f}{n}} \geq C_1(r, n). \quad (9)$$

The last line comes from calculations done in the proof of Theorem 5.1 (Appendix D).

We now study $p_1(S_r^{(0)})$. Since node 1 started the protocol then it means that no other node (and in particular 0) will stop emitting the message. Therefore, if node 0 is the first to communicate with a curious

node then it will remain active for the whole duration of the protocol. Consider that the first disclosure happens after T_f communications. We can write:

$$p_1 \left(S_r^{(0)} \right) \leq p_1(S_0 = 0) p_1 \left(\bigcap_{i=1}^r S_i \neq 0 \mid S_0 = 0, T_f \leq t_f \right) + p_1(T_f > t_f).$$

Since the fraction of curious nodes is constant, we can choose t_f independently of n or r such that $p(T_f > t_f) \leq e^{-\frac{f}{n} t_f} \leq \frac{\epsilon}{4(n-f)}$ if $t_f = \frac{n}{f} \log \left(\frac{4(n-f)}{\epsilon} \right)$ in order to control the second term. Then,

$$p_1 \left(\bigcap_{i=1}^r S_i \neq 0 \mid S_0 = 0, T_f \leq t_f \right) \leq \prod_{t=t_f}^{t_f+r} \left(1 - \frac{f}{n} \frac{1}{t} \right) \leq e^{-\frac{f}{n} \sum_{t=t_f}^{t_f+r} \frac{1}{t}}.$$

A series-integral comparison yields that if $r = \log^2(n)$ then $e^{-\frac{f}{n} \sum_{t=t_f}^{t_f+r} \frac{1}{t}} \leq \frac{\epsilon}{4}$ for n large enough. Finally, we use the fact that $p_1(S_0 = 0) \leq \frac{1}{n-f}$ to write that:

$$p_1 \left(S_r^{(0)} \right) \leq \frac{\epsilon}{2(n-f)}. \quad (10)$$

Finally, we observe that $C_1(\log^2 n, n) \rightarrow \frac{f}{n}$ when $n \rightarrow \infty$ where C_1 is defined in Equation 9. In particular, $C_1(\log^2 n, n) \geq \frac{f}{2n}$ for n large enough, so we have

$$\frac{p(I_0 \neq 0 \mid S_r^{(0)})}{p(I_0 = 0 \mid S_r^{(0)})} = \sum_{i \notin C \cup \{0\}} \frac{p_i(S_r^{(0)})}{p_0(S_r^{(0)})} \leq \frac{n}{f} \epsilon. \quad (11)$$

Since ϵ can be picked arbitrarily small and $\frac{n}{f}$ is assumed to be constant then the previous ratio can be made arbitrary small. \square

F PRIVACY GUARANTEES OF FAST PRIVATE GOSSIP

F.1 Differential privacy

We prove here the differential privacy guarantees for Algorithm 1 when $0 < s < 1$ (Theorem 5.2).

PROOF OF THEOREM 5.2. The proof of this theorem requires looking at very specific sequences and events that need to be picked carefully. In particular, it is necessary to distinguish between two types of sequences. If S is such that node 1 communicates with a curious node before node 0, then $p_0(S) \leq p_1(S)$ because node 0 needs to receive the rumor before being able to communicate it to curious nodes.

Suppose now that S is such that node 0 communicates with a curious node before node 1. In this case, we denote by F the event such that the source (in this case node 0) communicates with a curious node before stopping to emit for the first time. We recall that this happens with probability $1 - s$ after each call to the `tell_gossip` procedure. It is interesting to remark that in this case, $p_0(S \mid \bar{F}) \leq p_1(S \mid \bar{F})$. Indeed, if node 1 is the source then node 0 can communicate with a curious node as soon as it receives the rumor. If node 0 is the source then because of F it has to stop emitting before being able to communicate with a curious node. Therefore, conditionally on F , it is more likely that node 0 communicates with a curious node before node 1 does if node 1 is the source rather than if node 0 is the source. Then, we can write:

$$p_0(S) = p_0(S, \bar{F}) + p_0(S, F) \leq p_1(S, \bar{F}) + p_0(F) \leq p_1(S) + p_0(F).$$

Then, denoting by T_f the number of messages after which the source stops emitting, we write for $s > 0$:

$$p_0(F) = \sum_{k=1}^{\infty} p_0(T_f = k) p_0(F \mid T_f = k) = \sum_{k=0}^{\infty} (1-s) s^k \left(1 - \left(1 - \frac{f}{n} \right)^{k+1} \right).$$

Note that we can also write for $k \geq 1$:

$$p_0(F) = p_0(F, T_f \leq k) + p_0(F, T_f > k) \leq (1 - s^k) \left(1 - \left(1 - \frac{f}{n} \right)^k \right) + s^k = 1 - (1 - s^k) \left(1 - \frac{f}{n} \right)^k,$$

which concludes the proof. \square

F.2 Prediction uncertainty

We prove here that Algorithm 1 guarantees prediction uncertainty when $s < 1$ (Theorem 5.3).

PROOF OF THEOREM 5.3. For any set of sequences $S \subset \mathcal{S}$ such that $p_0(S) > 0$, we have:

$$\frac{p(I_0 \neq 0|S)}{p(I_0 = 0|S)} = \sum_{i \notin C \cup \{0\}} \frac{p_i(S)}{p_0(S)} \geq \sum_{i \notin C \cup \{0\}} \frac{p_i(A_1 = \{0\})p_i(S|A_1 = \{0\})}{p_0(S)},$$

where A_1 is the set of active nodes at round 1. Because the state of the system (active nodes) is the same in both cases we can write that $p_i(S|A_1 = \{0\}) = p_0(S)$. Besides, $p_i(A_1 = \{0\})$ corresponds to the probability that node i sends a message to node 0 and then stops emitting. Therefore:

$$\frac{p(I_0 \neq 0|S)}{p(I_0 = 0|S)} \geq \left(1 - \frac{f+1}{n} \right) (1-s) > 0.$$

This shows that the fast private gossip guarantees prediction uncertainty and concludes the proof. \square

REMARK F.1 (BACKGROUND KNOWLEDGE). *As stated in Section 3.3, prediction uncertainty guarantees are not robust to background knowledge. However, the definition can be adapted to include such knowledge, for example to model the fact that attackers may know that some nodes did not start the rumor. This corresponds to changing the $p(I_0 \neq \{0\}|S)$ to $p(I_0 \in J|S)$, where J is the set of nodes that can actually be suspected.*

G DISSEMINATION SPEED FOR ALGORITHM ??

In this section, we prove Theorem 5.4 and discuss how to extend this result to the asynchronous version of the algorithm.

G.1 Synchronous version

PROOF OF THEOREM 5.4. This idea of the proof is to rely on the ‘‘determinism’’ of gossip process, similarly to [52]. This means that the gossip process very closely follows its mean dynamics. In our case, there is an added difficulty in the fact that extra randomness is introduced by the deactivation of the nodes. Yet, we precisely quantify the impact of this phenomenon on the results. We start by showing that if more than $k(s)$ nodes are informed at a given time, then with very high probability the number of informed nodes will never drop below this fraction. Therefore, a number of messages proportional to the size of the graph will be sent at each round. The condition on s for this to happen is written in Equation 16. More formally, we fix $s \in (0, 1]$ and denote by A_t the number of nodes that are active at round t , which is such that $A_t = \alpha_t n$. Then, we note

$$f : \alpha \rightarrow 1 - p_u(\alpha)(1 - \alpha s), \tag{12}$$

where $p_u(\alpha) = (1 - \frac{1}{n})^{\alpha n}$. Note that $f(\alpha)$ can be rewritten $f(\alpha) = \frac{1}{n} \mathbb{E}[A_{t+1} - A_t | A_t = \alpha n]$. As a matter of fact, for each node, the probability of getting the message is exactly $1 - p_u(\alpha)$ so $n(1 - p_u(\alpha))$ nodes get the message in expectation. The rest of the active nodes at the following round is made of the nodes that were active, did not receive the message and did not deactivate, which represents a portion $n\alpha p_u(\alpha)s$ of the

nodes. Then, one can see that the function f is simply the sum of these 2 terms. We show by using that $(1-x)^y \leq e^{-xy} \leq 1-xy + \frac{x^2y^2}{2}$ that for $\alpha \leq \alpha_s = \frac{s}{1+2s}$, we have:

$$f(\alpha) \geq \left(1 + \frac{s}{2}\right)\alpha. \quad (13)$$

Then, we follow the same steps as in Lemma 15 in [52]. We call A_t the number of active nodes at round t , and $A_{t,m}$ the number of active nodes at round t after m messages have been sent (so during the round). Then, we can define $X_i = A_{t,i+1} - A_{t,i}$. $A_{t,i+1}$ only depends on $A_{t,i}$ and so does X_i :

$$X_i = \begin{cases} 1 & \text{with proba } s(1 - \frac{|A_{t,i}|}{n}) \\ -1 & \text{with proba } (1-s)\frac{|A_{t,i}|-1}{n} \\ 0 & \text{otherwise} \end{cases}$$

Then, we define the martingale $Z_i = \mathbb{E}[\sum_{i=1}^{A_t} X_i | X_1, \dots, X_i, A_t]$. This allows us to write $A_{t+1} - nf(\alpha) = Z_{A_t} - Z_0$. If we call $S_{k,t} = \sum_{i=k}^{A_t} X_i$ then for any $d \in \{-1, 0, 1\}$:

$$\mathbb{E}[S_{1,t} | X_1, \dots, X_i, X_{i+1} = 1, A_t] \geq \mathbb{E}[S_{1,t} | X_1, \dots, X_i, X_{i+1} = d, A_t] \geq \mathbb{E}[S_{1,t} | X_1, \dots, X_i, X_{i+1} = -1, A_t],$$

because the distribution of X_i only depends on $A_{t,i}$. Therefore, $|Z_{i+1} - Z_i| \leq (1 + \mathbb{E}[S_{i+1,t} | A_t + 1]) - (\mathbb{E}[S_{i+1,t} | A_t - 1] - 1) \leq 2$. Azuma's inequality [48] then gives:

$$p\left(A_{t+1} - nf\left(\frac{A_t}{n}\right) \leq -\lambda A_t | A_t = k\right) \leq e^{-\frac{\lambda^2 k}{8}}. \quad (14)$$

We also have that $p(A_{t+1} < k | A_t \geq k) \leq p(A_{t+1} \leq k | A_t = k)$. Then, for any $\alpha \leq \alpha_s$ we have that $p(A_{t+1} < k | A_t \geq k) \leq p(A_{t+1} - nf(\frac{A_t}{n}) \leq -\frac{s}{2} A_t | A_t = k)$ by using Equation 13. We can then bound this expression by using Equation 14 with $\lambda = \frac{s}{2}$. Therefore, we need that:

$$\alpha \leq \alpha_s. \quad (15)$$

Denoting by $N_{k,j}$ the number of messages sent between rounds k and j , we can decompose over $C\alpha^{-1} \log n$ rounds so that if m is such that there are at least α active nodes at round m then:

$$p(N_{m, m+C\alpha^{-1} \log n} \geq Cn \log n) \geq (1 - e^{-\frac{s^2 \alpha n}{32}})^{C\alpha^{-1} \log n},$$

because it is equal to the probability that the fraction of active nodes never goes below α for $C\alpha^{-1} \log n$ rounds. Therefore, if

$$s^2 \geq \frac{32}{\alpha n} \log \frac{3Cn \log n}{\alpha}, \quad (16)$$

then $p(N_{m, m+C\alpha^{-1} \log n} \geq Cn \log n) \geq 1 - \frac{1}{3n}$.

Equation 16 gives a lower bound on the value of α . Note that this lower bound goes to 0 as n grows so in particular, Equation 16 is satisfied for $\alpha = \alpha_s$ if n is large enough. It now remains to show that such a fraction α of active nodes can be reached in logarithmic time. Usual gossip analysis takes advantage of the exponential growth of the informed nodes during early rounds for which no collision occur. We have to adapt the analysis to the fact that nodes may stop communicating with some probability and split the analysis into two phases.

In the rest of the proof, we prove that a constant fraction of the nodes (independent of n) can be reached with a logarithmic number of rounds. We first analyze how long it takes to go from $O(\log n)$ to $O(n)$ active nodes and then from 1 to $O(\log n)$.

Equation 13 along with Equation 14 with $\lambda = \frac{s}{4}$ give that as long as $A_{t_0}(1 + \frac{s}{4})^t \leq \alpha_s n$ then

$$p\left(A_{t+t_0+1} \geq A_{t_0}\left(1 + \frac{s}{4}\right)^{t+1} | A_t = A_{t_0}\left(1 + \frac{s}{4}\right)^t\right) \geq 1 - e^{-\frac{\alpha n s^2}{128}} \quad (17)$$

for any $t \geq t_0$ such that $A_{t_0} \left(1 + \frac{\epsilon}{2}\right)^t \leq n\alpha_s$. Therefore, if we do this for all $t \leq t_\alpha = \frac{\log(\alpha n)}{\log(1 + \frac{\epsilon}{2})}$ rounds (so for a logarithmic number of rounds) then $p(A_{t_\alpha+t_0} \geq n\alpha | A_{t_0}) \geq \left(1 - e^{-\frac{A_{t_0}s^2}{128}}\right)^{t_\alpha}$ because in this case, $A_t \geq A_{t_0}$ for $t \geq t_0$. Therefore, if

$$A_{t_0} \geq -\frac{128}{s^2} \log \left(1 - \left(1 - \frac{1}{3n}\right)^{\frac{1}{t_\alpha}}\right), \quad (18)$$

then $p(A_{t_\alpha+t_0} \geq n\alpha) \geq 1 - \frac{1}{3n}$. If we use the fact that $(1-x)^y \leq e^{-xy} \leq 1 - xy + \frac{x^2y^2}{2}$ to simplify Equation 18, we can show that if A_{t_0} satisfies:

$$A_{t_0} \geq \frac{128}{s^2} \left(\log n + \log(3t_\alpha) - \log \left(1 - \frac{1}{6nt_\alpha}\right) \right), \quad (19)$$

then it also satisfies Equation 18. Since the terms on the right hand side are dominated by $\log n$, for n large enough, a sufficient condition for Equation 18 to hold is:

$$A_{t_0} \geq \frac{256}{s^2} \log n. \quad (20)$$

It only remains to prove that a logarithmic number of nodes can be reached in logarithmic time. For this, we use again Azuma inequality but we start from the very beginning of the protocol ($|A_0| = 1$) and for a fixed sequence of m messages. This time, we write $S_n = \sum_{i=1}^m X_i$ with the exact same notations, and by calling $A_{0,m}$ the number of actives nodes after m messages (without taking rounds into account) then $p(A_{0,m} - a \leq -\lambda) \leq e^{-\frac{\lambda^2}{8m}}$ for any $a \leq \mathbb{E}[S_n]$. Then, we denote E_i the event such that $X_i \geq 0$ for all i and write that $\mathbb{E}[S_n] \geq p(E_i)\mathbb{E}[S_n|E_i]$. Considering that $A_{0,m} \leq m$, we can write that $\mathbb{E}[S_n] \geq (1 - (1-s)\frac{m}{n})^m ms(1 - \frac{m}{n})$. Therefore, we can apply Azuma inequality with $\lambda = ms \left(\frac{1}{2} - \frac{m}{n} [(1-s)m - 1]\right)$, which yields:

$$p(A_{0,m} \leq \frac{ms}{2}) \leq e^{-\frac{ms^2}{8} \left(\frac{1}{2} - \frac{m}{n} [(1-s)m - 1]\right)^2}. \quad (21)$$

The number of messages sent during rounds 1 to t_0 is at least equal to $m \geq t_0$. We set $t_0 = \frac{512}{s^3} \log(3n)$, and since for n large enough we have $\frac{m}{n} [(1-s)m - 1] \leq \frac{1}{4}$, then

$$p\left(A_{t_0} \geq \frac{256}{s^2} \log(n)\right) \geq 1 - e^{-\frac{ms^2}{128}} \geq 1 - \frac{1}{3n}. \quad (22)$$

We conclude the proof by noting that

$$\begin{aligned} p\left(N_{0,t_0+t_\alpha+C\alpha^{-1}\log n} \geq Cn \log n\right) &\geq p\left(A_{t_0} \geq \frac{256}{s^2} \log n\right) p\left(A_{t_\alpha+t_0} \geq n\alpha | A_{t_0} \geq \frac{256}{s^2} \log n\right) \\ &\geq \left(1 - \frac{1}{3n}\right)^3 \geq 1 - \frac{1}{n}. \end{aligned}$$

The number of rounds is logarithmic since both t_0 and t_α depend logarithmically on n . \square

G.2 Extension to the asynchronous version

The first part of the proof directly extends to the asynchronous algorithm by simply considering slices of time during which a set of αn nodes send αn messages, which essentially means constant time. Then, we consider a logarithmic number of slices. The phase from 1 to $\mathcal{O}(\log n)$ active nodes requires sending a logarithmic number of messages and can thus be done in logarithmic time. Finally, phase 2 (going from $\mathcal{O}(\log n)$ to $\mathcal{O}(n)$ active nodes) consists in evaluating a logarithmic number of rounds during which a logarithmic number of nodes are active. Again, the only important thing is the number of messages sent

(and not which node sent them) so using constant time intervals ensures that enough messages are sent between each pseudo-rounds with high probability.

Therefore, it is possible to prove a statement very similar to that of Theorem 5.4 in the asynchronous setting, where the notion of rounds is replaced by constant time intervals. We omit the exact details of this alternative formulation.

H CHALLENGES OF PRIVATE GOSSIP FOR GENERAL GRAPHS

A natural extension of this work is to consider general graphs. We discuss in this section several aspects related to the natural privacy of gossip protocols in arbitrary graphs. In particular, we highlight the fact that problem-specific modeling choices are needed to go beyond the complete graph, and that even defining a notion of privacy that is suitable for all graphs is very challenging.

H.1 Average-case versus worst-case privacy

Unlike the case of complete graphs, the location of curious nodes critically impacts the privacy guarantees in arbitrary graphs. A naive way to deal with this issue is to randomize the location of curious nodes *a posteriori*. Let us denote by $\mathcal{L}_{i,j}^f$ the set containing all subsets of nodes of size f of the graph that do not contain i and j . For fixed nodes i and j , the set of curious nodes C is drawn from $U(\mathcal{L}_{i,j}^f)$, the uniform distribution over $\mathcal{L}_{i,j}^f$. For some parameters $\epsilon, \delta \geq 0$, privacy can be defined as follows:

$$\forall i, j \in \{0, \dots, n-1\}, \forall S \in \mathcal{S}, \mathbb{E}_{C \sim U(\mathcal{L}_{i,j}^f)} [p_i(S, C) - e^\epsilon p_j(S, C)] \leq \delta.$$

Note that $p_i(S, C) = 0$ if the output sequence S is not compatible with the set of curious nodes C , *i.e.* if $(k, l) \in S$ and $k, l \notin C$. To pick the curious nodes, it is possible to either pick a set of f curious nodes at once or to pick each node (except for i and j) with probability f/n . This randomized definition allows to prove a bound similar to that of Theorem 4.1 for arbitrary graphs. Indeed, the first node that receives the rumor has probability $\frac{f}{n}$ of being a curious node. However, such average-case notions of privacy are highly undesirable: in this case, no protection is provided against a (much more realistic) adversary that controls a fraction of nodes fixed in advance.

The worst-case approach consists in bounding the maximum difference instead of the expectation. This is the approach taken in our work for the complete graph (the max operator is implicit because the location of curious nodes does not matter in a complete graph). In the case of general graphs, the corresponding privacy definition is given by:

$$\forall i, j \in \{0, \dots, n-1\}, \forall S \in \mathcal{S}, \max_{C \in \mathcal{L}_{i,j}^f} [p_i(S, C) - e^\epsilon p_j(S, C)] \leq \delta.$$

We immediately observe that with this definition, it is impossible to have $\delta < 1$ as soon as there is a node in the graph with less than f neighbors. This modeling choice is quite unrealistic as well because having a node surrounded by curious nodes means that the attacker actually believes this specific node has a strong probability of being the source and therefore put more sensors around it. A possible alternative would be to place curious nodes so as to bound the maximum privacy for any pair of nodes, and then evaluate the minimum privacy in this setting. This definition would mean that the attacker wants to be able to distinguish any pair of nodes as best as possible.

We see that choosing the locations of the curious nodes in an arbitrary graph is a complex problem that is heavily dependent on the topology of the graph and on the prior of the attacker on the locations of the curious nodes. Indeed, the attacker may simply want to isolate a sufficiently small group of nodes that have a high probability of being the source.

H.2 Relaxing the differential privacy definition

Differential privacy is a very strong notion that enforces indistinguishability between all pairs of nodes, in order to be robust to any prior information about who might be the source. In particular, an attacker should not be able to precisely identify the source even if it knows that only two nodes in the graph can be the source. Although it was possible to obtain meaningful privacy guarantees of this kind for the complete graph, this appears to be too strong of a requirement for some graph topology and location of curious nodes. Consider for instance the extreme case of a line graph. It is clear that any non-trivial attacker can always distinguish between two segments of the line. This intuition directly extends to any graph which admits a cut with only curious nodes in it.

A natural idea is to restrict the pairs of nodes that are required to be indistinguishable. Several ways of doing this may be considered. For instance, one could require that each node is indistinguishable from k other nodes in the graph. Such relaxed definition could be obtained using the Pufferfish framework [45], which explicitly provides a notion of secret to protect. But how to choose such k nodes based on the topology and how to characterize the optimal locations of curious nodes is very challenging. Another direction could be to adapt the notions of metric-based differential privacy [4, 11] to design a notion of privacy where the required indistinguishability for a given node is a function of its distance to curious nodes in the graph, or to require that pairs of nodes become less indistinguishable with distance in the graph. Yet, it is not clear how to characterize the influence of the graph topology.

H.3 Optimality of private gossip

We have seen in this section that the privacy guarantees for arbitrary graphs heavily rely on the particular privacy notion and that some recent privacy frameworks may provide tools to relax the classic differential privacy definition which is generally too strong for arbitrary graphs. We conjecture that for some of these relaxed definitions, the optimal algorithm for general graphs will be the same as in our case of the complete graph. Indeed, the strength of the private gossip protocol (Algorithm 1 with $s = 0$) is to forget initial conditions quickly. In the complete graph, it does so in one step. In an arbitrary graph, the information about the part of the graph the source belongs to is still present after some steps, but the source should quickly be completely indistinguishable from its direct neighbors. In particular, attacks based on centrality [54–56] are rather meaningless against this algorithm because spreading only occurs along a random walk in the graph. Our fast private gossip protocol is likely to enjoy near-optimal privacy guarantees since it is an approximation of the private gossip protocol.