Leakage Resilient Secret Sharing and Applications

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Abstract

A secret sharing scheme allows a dealer to share a secret among a set of n parties such that any authorized subset of the parties can recover the secret, while any unauthorized subset of the parties learns no information about the secret. A local leakage-resilient secret sharing scheme (introduced in independent works by (Goyal and Kumar, STOC 18) and (Benhamouda, Degwekar, Ishai and Rabin, Crypto 18)) additionally requires the secrecy to hold against every unauthorized set of parties even if they obtain some bounded local leakage from every other share. The leakage is said to be local if it is computed independently for each share. So far, the only known constructions of local leakage resilient secret sharing schemes are for threshold access structures for very low ($O(1)$) or very high ($n - o(\log n)$) thresholds.

In this work, we give a compiler that takes a secret sharing scheme for any monotone access structure and produces a local leakage resilient secret sharing scheme for the same access structure, with only a constant-factor blow-up in the sizes of the shares. Furthermore, the resultant secret sharing scheme has optimal leakage-resilience rate i.e., the ratio between the leakage tolerated and the size of each share can be made arbitrarily close to 1. Using this secret sharing scheme as the main building block, we obtain the following results:

- **Rate Preserving Non-Malleable Secret Sharing.** We give a compiler that takes any secret sharing scheme for a 4-monotone access structure[1] with rate R and converts it into a non-malleable secret sharing scheme for the same access structure with rate $\Omega(R)$. The prior such non-zero rate construction (Badrinarayanan and Srinivasan, 18) only achieves a rate of $\Theta(R/t_{\max} \log^2 n)$, where $t_{\max}$ is the maximum size of any minimal set in the access structure. As a special case, for any threshold $t \geq 4$ and an arbitrary $n \geq t$, we get the first constant rate construction of t-out-of-n non-malleable secret sharing.
- **Leakage-Tolerant Multiparty Computation for General Interaction Pattern.** For any function, we give a reduction from constructing leakage-tolerant secure multi-party computation protocols obeying any interaction pattern to constructing a secure (and not necessarily leakage-tolerant) protocol for a related function obeying the star interaction pattern. This improves upon the result of (Halevi et al., ITCS 2016), who constructed a protocol that is secure in a leak-free environment.

[1] A 4-monotone access structure has the property that any authorized set has size at least 4.

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1 Introduction

Secret sharing [Sha79, Bla79] is a fundamental cryptographic primitive that allows a secret to be shared among a set of parties in such a way that only certain authorized subsets of parties can recover the secret by pooling their shares together; and any subset of parties that is not authorized learn nothing about the secret from their shares. Secret sharing has had widespread applications across cryptography, ranging from secure multiparty computation [GMW87, BGW88, CCD88], threshold cryptographic systems [DF90, Fra90, DDFY94] and leakage resilient circuit compilers [ISW03, FRR+10, Rot12]

In this paper, we study secret sharing with a stronger privacy requirement – that the secret remain hidden from unauthorized subsets of parties even if they have access to some small amount of information about the shares of the remaining parties (referred to as leakage). Such *leakage resilience* has been widely studied in the past as a desirable property in various settings and cryptographic primitives [MR04, DP08, AGV09, NS09].

More specifically, we are interested in *local* leakage resilience, which means that secrets are hidden from an adversary that works as follows. First, it specifies an unauthorized subset of parties, and for each of the remaining parties, it specifies a leakage function that takes its share as input and outputs a small pre-determined number of bits. Once the shares are generated, the adversary is given all the shares of the unauthorized subset, and the output of the corresponding leakage function applied to the each of the remaining shares.

This form of leakage resilience for secret sharing was formalized in recent work by Goyal and Kumar [GK18a], and Benhamouda, Degwekar, Ishai and Rabin [BDIR18]. These papers showed examples of leakage-resilient threshold secret sharing schemes (where subsets above a certain size are authorized) for certain thresholds. They then showed applications of such schemes to constructions of leakage-resilient multi-party computation protocols and non-malleable secret sharing schemes. Given the prevalence of secret-sharing in cryptographic constructions and the importance of resilience to leakage, one may reasonably expect many more applications to be discovered in the future.

In this work, we are interested in constructing local leakage resilient secret sharing schemes for a larger class of access structures[2] (in particular, for all thresholds). Beyond showing feasibility, our focus is on optimizing the following parameters of our schemes:

- the *rate*, which is the ratio of the size of the secret to the size of a share, and,
- the *leakage-resilience rate*, which is the ratio of the number of bits of leakage tolerated per share to the size of a share.

1.1 Results and Techniques

Our primary result is a transformation that converts a secret sharing scheme for any access structure $\mathcal{A}$ into a local leakage resilient secret sharing scheme for $\mathcal{A}$ whose rate is a small constant factor less than that of the original scheme, and which has an optimal leakage-resilience rate of 1.

Informal Theorem 1.1 *There is a compiler that, given a secret sharing scheme for a monotone access structure $\mathcal{A}$ with rate R, produces a secret sharing scheme for $\mathcal{A}$ that has rate $R/3.01$ and is local leakage resilient with leakage-resilience rate tending to* 1.

[2] The access structure of a secret sharing scheme is what we call the set of authorized subsets of parties.

In particular, for any $k \leq n$, starting from k-out-of-n Shamir secret sharing [Sha79] gives us a k-out-of-n threshold secret sharing scheme with rate $1/3.01$ and leakage rate 1. The only results known even for threshold access structures before our work were for either very small or very large thresholds. Goyal and Kumar [GK18a] presented a construction for $k = 2$, which had both rate and leakage-resilience rate $\Theta(1/n)$. This was extended to any constant k by Badrinarayanan and Srinivasan [BS18], with rate $\Theta(1/\log(n))$ and leakage-resilience rate $\Theta(1/n\log(n))$. Benhamouda et al. [BDIR18] showed that k-out-of-n Shamir secret sharing over certain fields is local leakage-resilient if $k = n - o(n)$, and this has rate 1 and leakage-resilience rate roughly $1/4$.

Outline of Construction. We will now briefly describe the functioning of our compiler for the case of a k-out-of-n threshold secret sharing scheme, for simplicity. It makes use of a strong seeded randomness extractor Ext, which is an algorithm that takes two inputs – a seed s and a source w – and whose output $\mathrm{Ext}(s, w)$ is close to being uniformly random if s is chosen at random and w has sufficient entropy. The extractor being "strong" means that the output remains close to uniform even if the seed is given.

We take any threshold secret sharing scheme (such as Shamir's [Sha79]), and share our secret m with it to obtain the set of shares $(\mathsf{Sh}_1, \ldots, \mathsf{Sh}_n)$. We first choose a uniform seed s, and for each $i \in [n]$, we choose a uniformly random source w_i (all of appropriate lengths), and mask Sh_i using $\mathrm{Ext}(s, w_i)$. That is, we compute $\mathsf{Sh}'_i = \mathsf{Sh}_i \oplus \mathrm{Ext}(s, w_i)$. We then secret share s using a 2-out-of-n secret sharing scheme to get the set of shares $S_1, \ldots, S_n$. The share corresponding to party i in our scheme is now set to $(w_i, \mathsf{Sh}'_i, S_i)$.

Given k such shares, to recover the secret, we first reconstruct the seed from any two S_i's and then unmask Sh'_i by XORing with $\mathrm{Ext}(s, w_i)$ to obtain Sh_i. We then use the reconstruction procedure of the underlying secret sharing to recover the message.

The correctness and privacy of the constructed scheme are straightforward to check. To argue the local leakage resilience of this construction, we go over a set of $n - k + 1$ hybrids where in each hybrid, we will replace one Sh_i with the all 0's string. Once we have replaced $n - k + 1$ such shares with the 0's string, we can then rely on the secrecy of the underlying secret sharing scheme to show that the message is perfectly hidden. Thus, it is now sufficient to show that any two adjacent hybrids in the above argument are statistically close. To argue that the adjacent hybrids, say Hyb_i and Hyb_{i+1}, are statistically close, we rely on the randomness property of the extractor. Note that even given bounded leakage from the source w_i, we can show that w_i has sufficient entropy so that the output of the extractor on the weak source w_i is statistically close to random. This allows us to argue that $\mathrm{Ext}(s, w_i)$ acts as a one-time pad and thus, we can replace Sh_i with the all 0's string.

However, in order to make the argument work, we must ensure that the leakage from the source is independent of the seed. This is where we will be using the fact that the seed is secret shared using a 2-out-of-n secret sharing scheme. In our reduction, we fix the share S_i to be independent of the seed and then leak from the source w_i. Once the seed is known[3], we can sample the other shares $(S_1, \ldots, S_{i-1}, S_{i+1}, \ldots, S_n)$ as a valid 2-out-of-n secret sharing of s that is consistent with the fixed share S_i. This allows us to argue that the leakage on w_i is independent of the source. There is a small caveat here that the masked value Sh'_i is dependent on the seed and hence we cannot argue independence of the leakage on the source and the seed. However, we use a simple trick of masking Sh'_i by another one-time pad and then secret share the one-time pad key along with the seed s and use this argue that this masked value is independent of the seed.

[3] As the extractor is a strong seeded extractor, $\mathrm{Ext}(s, w_i)$ is statistically close to uniform even given the seed.

This construction described above has several advantages. The main advantage is that the transformation is rather simple and only incurs a very small overhead when compared to the original secret sharing scheme. In particular, the rate of the resultant leakage resilient secret sharing has only a small constant factor loss when compared to the initial secret sharing scheme. Also, we can sample the seed s of the extractor once and use it for sharing multiple secrets.[4] The second advantage is that it easily generalizes to all monotone access structures, basically, the only difference is that we use a secret sharing scheme for this access structure to obtain the set of shares $(\mathsf{Sh}_1, \ldots, \mathsf{Sh}_n)$, and the rest of the steps are exactly the same as before. The third advantage is that the resultant secret sharing scheme has optimal leakage-resilience rate, i.e., the ratio between the number of bits of leakage tolerated and size of the share tends to 1 as the amount of leakage that the scheme is designed to handle increases. Finally, if we use the inner product two-source extractor of Chor and Goldreich [CG88] as the underlying extractor then the sharing procedure is a linear function of the secret and a quadratic function of the randomness, and this can be implemented efficiently.

Rate-Preserving Non-Malleable Secret Sharing. Non-malleable secret sharing schemes, introduced by Goyal and Kumar [GK18a], are secret sharing schemes where it is not possible to tamper with the shares of a secret s (in certain limited ways) so as to convert them to shares corresponding to a different secret s' that is related to s (such as $s+1$ or s with the first bit flipped). We are interested in security against an adversary that tampers each share independent of the others (called individual tampering). In this setting, Goyal and Kumar presented a construction of a non-malleable k-out-of-n threshold secret sharing scheme, and in a later paper [GK18b] extended this to general access structures. Their constructions, however, had an asymptotic rate of zero.

Badrinarayanan and Srinivasan [BS18] gave a compiler that takes any secret sharing scheme for a 4-monotone[5] access structure and outputs a non-malleable secret sharing scheme for the same access structure. The main tool used in their compiler was a local leakage resilient threshold secret sharing scheme. The loss in the rate of the resulting non-malleable secret sharing scheme depended on the parameters of the underlying local leakage resilient secret sharing. In particular, to have only a constant loss in the rate, it was important to have a local leakage resilient threshold secret sharing scheme that had a constant rate and a constant leakage-resilience rate. We plug in our basic leakage resilient secret sharing scheme that has both these features with the compiler of Badrinarayanan and Srinivasan to obtain a rate-preserving compiler for non-malleable secret sharing.

Informal Theorem 1.2 *There is a compiler that, given a secret sharing scheme for a 4-monotone access structure $\mathcal{A}$ with rate R, produces a secret sharing scheme for $\mathcal{A}$ that has rate $\Omega(R)$ and is non-malleable against individual tampering.*

Leakage-Tolerant MPC for General Interaction Pattern. Next, we provide an application of our constructions to multi-party computation, an area where secret sharing is rather pervasive. An interaction pattern is a directed graph which specifies the sequence of messages that have to be sent during the execution of a protocol. To give an example, consider a ring interaction pattern. Here, the first message is sent by the party P_1 to the party P_2 and depending on this message, P_2

[4] For the security of this modification to go through, we need the adversary to specify all the secrets and leakage functions upfront – it cannot adaptively choose the secrets and leakage functions depending on the previous leakage.

[5] k-monotone means that all authorized sets in the access structure are of size at least k.

sends a message to P_3 and so on. Finally, the party P_n sends a message to P_1 who computes the output based on this message. The directed graph corresponding to this has $(n+1)$ nodes, one corresponding to each message and one for the output, and the graph is a single directed path that goes from the first message to the last and then to the output node. To give another example, a standard 2-round MPC protocol with n parties can be represented by an interaction pattern graph with two sets of n^2 nodes, representing the messages sent by each party to each other party in the two rounds. The edges then go from the nodes corresponding to first-round messages to second-round messages, according to the protocol.

It is well-known from a sequence of works [HLP11,GGG+14,BGI+14] that the standard notion of MPC which guarantees that only the output is leaked is impossible to achieve for certain interaction patterns. To see this, consider the star interaction pattern where there is a special party called as the evaluator and every party sends a single message to the evaluator who computes the output. In this interaction pattern, if the evaluator colludes with some subset of the parties then it is easy to see that the colluding parties learn the residual function with the honest parties inputs to be fixed and the corrupted parties inputs to be varied. To capture this inherent leakage with certain interaction patterns, Halevi et al. [HIJ+16] introduced the notion of fixed and free inputs. Intuitively, every honest party's input is fixed. In addition to this, every corrupted party's input is fixed if there exists a path from the corrupted party to the evaluator that passes through at least one honest party. The rest of the corrupted parties' inputs are free. A secure MPC protocol that is compliant with an interaction pattern guarantees that only the residual function with the above defined notion of fixed inputs is revealed to the adversary.

We define the notion of leakage tolerance for an MPC protocol that is compliant with an interaction pattern along the same lines as that of leakage tolerant MPC [BCH12]. In the setting of leakage tolerance, as in the standard setting, we consider an adversary who corrupts an arbitrary subset of parties and can see their entire views. But in addition to this, the adversary also obtains bounded leakage on the complete internal state – that includes both the input, the secret randomness, and the entire view of the protocol – of every honest party. Here, the adversary can potentially learn bounded information about the honest party's input since it has access to all of the honest parties' secret state. We would like to guarantee that nothing beyond such bounded information about the inputs and the residual function is actually leaked to the adversary. Technically, we account for this leakage by allowing the simulator to learn the same amount of information about the honest parties' inputs.

What makes the task of providing such security non-trivial is that unlike a standard MPC simulator who is allowed to cheat in generating the protocol messages, a simulator in the leakage tolerance setting cannot deviate from the protocol specification. This is because any deviation can be caught by the adversary by leveraging the leakage on the secret state of the honest party. At first sight, the task of designing such a simulator seems impossible as we require the simulator to generate the correct protocol messages based only the output (or more generally, based on the residual function). However, notice that the leakage functions are local to the honest party's view. Hence, the simulator must follow the protocol correctly at the local level but must somehow cheat in the global level, i.e., in generating the joint distribution of the protocol messages. To make this task even more demanding, we do not wish to use any computational assumptions and only make use of information theoretic tools to achieve leakage tolerance.

In this setting, we upgrade one of the results of Halevi et al [HIJ+16] to have the additional guarantee of leakage tolerance. They showed that the star interaction pattern described earlier is

complete for MPC on general interaction patterns – given a secure protocol for a function f that is compliant with the star interaction pattern, they showed how to construct a secure protocol for f compliant with any other interaction pattern. Similarly, we obtain the following.

Informal Theorem 1.3 *There is a compiler that, given a function $f : \{0,1\}^n \to \{0,1\}$, an interaction pattern $\mathcal{I}$, and a secure protocol for f compliant with the star interaction pattern, produces a secure protocol (with a setup phase producing correlated randomness) for f compliant with $\mathcal{I}$ that is also leakage tolerant. Further, if in the original protocol each party uses M bits of communication and R bits of correlated randomness, then each party in the new protocol uses $O(n^2M)$ bits of communication and $O(R + n^2M)$ bits of correlated randomness.*

Using the known protocols for the star interaction pattern [BGI+14, BKR17, GGG+14], we obtain the following corollaries.

Corollary 1.1 (from [BGI+14]) *For any function $f : \{0,1\}^n \to \{0,1\}^*$ and any interaction pattern $\mathcal{I}$, there exists a $\mathcal{I}$-compliant protocol that computes f with leakage tolerance upto $n-1$ passive corruptions. The communication complexity of the protocol is exponential in n.*

Corollary 1.2 (from [BKR17]) *For any function $f : \{0,1\}^n \to \{0,1\}^*$ that is computable in NC^1 and any interaction pattern $\mathcal{I}$, there exists an efficient, $\mathcal{I}$-compliant protocol that computes f with leakage tolerance upto a constant number of passive corruptions. Furthermore, assuming the existence of one-way functions, for any function $f : \{0,1\}^n \to \{0,1\}^*$ that is computable by a circuit, there exists an efficient $\mathcal{I}$-compliant protocol that computes f with leakage tolerance upto a constant number of passive corruptions.*

Corollary 1.3 (from [GGG+14]) *Assuming the existence of indistinguishability obfuscation and one-way functions, for any interaction pattern $\mathcal{I}$ and any function $f : \{0,1\}^n \to \{0,1\}^*$ that is computable by circuits, there exists an efficient, $\mathcal{I}$-compliant protocol that computes f with leakage tolerance upto $n-1$ passive corruptions.*

Our actual construction also covers functions where each party has multiple bits as input (see Theorem 5.11). The compiler we use is the same as that of Halevi et al., except using a leakage-resilient secret sharing scheme where theirs uses additive secret sharing. However, the proof of leakage tolerance is quite involved, and, in fact, it turns out that standard local leakage resilience is insufficient for this purpose. We now provide some intuition on why this is the case. In the Halevi et al.'s construction, some set of secrets are shared among all the parties in the correlated randomness generation phase. The messages sent during the execution of the protocol comprise of a subset of a party's shares. So, a party's secret state not only includes its own shares, but also the shares received from the other parties. Thus, any leakage function on a honest party's internal state is not local as it gets to see a subset of other parties' shares. Thus, we need a secret sharing scheme satisfying a stronger notion of leakage resilience that we describe next.

Stronger Leakage-Resilience. Motivated by the aforementioned application to leakage tolerance in multi-party computation, we study secret sharing resilient to a stronger form of leakage. In the earlier definition of local leakage, the leakage functions that are applied on the shares of honest parties are required to be specified independently of the shares that are completely revealed

to the adversary. In our stronger definition, these leakage functions are allowed to depend on some number of the adversary's shares. We construct k-out-of-n threshold secret sharing schemes that are resilient to such stronger leakage where the adversary is given $(k-1)$ shares, and the leakage functions applied on the honest party's shares are allowed to depend on $(k-2)$ of these shares. This construction, which is in fact a simple modification of our earlier one, has worse rate, but still has optimal leakage-resilience rate. Referring temporarily to the above as $(k-2, k-1)$-strong local leakage, we have the following.

Informal Theorem 1.4 *For any $k \leq n$, there is a k-out-of-n threshold secret sharing scheme that is resilient against $(k-2, k-1)$-strong local leakage, has rate $\Omega(1/n)$, and leakage-resilience rate tending to 1.*

It is easy to check that this definition is impossible to achieve for a k-out-of-n threshold secret sharing scheme if we allow the leakage functions to depend on all $(k-1)$ of the adversary's shares, as the leakage function on any honest party's share can use the $(k-1)$ shares along with this share to reconstruct the secret and leak a few bits of the secret. Even when $k = n$, we note that construction of a n-out-of-n secret sharing in the work of Benhamouda et al. [BDIR18] does not imply the above theorem because their construction only has a constant privacy error when the leakage function on the honest party's share is allowed to depend on $k-2$ corrupted shares.

1.2 Related Work

In concurrent and independent work, Aggarwal et al. [ADN$^+$18] also construct leakage-resilient secret sharing schemes for any access structure (from any secret sharing scheme for that access structure). They use this to construct non-malleable secret sharing for general monotone access structures and threshold signatures that are resilient to leakage and mauling attacks. Whereas our construction of non-malleable secret sharing for general monotone access structures is rate-preserving, their construction has an asymptotic rate of 0. On the other hand, their work additionally considers the stronger model of concurrent tampering and gives positive results in this model as well.

In another concurrent and independent work, Kumar et al. [KMS18] also consider the problem of obtaining leakage-resilient secret sharing scheme in a stronger leakage model. In particular, they consider a leakage model where every bit of the leakage can depend on the shares of an adaptive chosen set of $O(\log n)$ parties. They give constructions of such secret sharing schemes for general access structures via a connection to problems that have large communication complexity. The rate and the leakage-resilience rate of the construction of Kumar et al. are both $\Theta(1/\text{poly}(n))$. As an application, they give a construction leakage-resilient non-malleable secret sharing scheme where the tampering function can obtain bounded, adaptive leakage from each share. In comparison, we give construction of leakage-resilient secret sharing scheme against non-adaptive leakage functions where the leakage on the honest party's share could depend on at most $n-2$ shares with rate $\Omega(1/n)$ and leakage-resilience rate tending to 1.

Most closely related to our work are the papers by Goyal and Kumar on non-malleable secret sharing [GK18a, GK18b], Benhamouda et al on leakage-resilient secret sharing and MPC [BDIR18], and Badrinarayanan and Srinivasan on non-malleable secret sharing with non-zero rate [BS18].

Local leakage resilient secret sharing (in the sense in which we use this term) was first studied by Goyal and Kumar [GK18a] and Benhamouda et al [BDIR18] (independently of each other). [GK18a]

constructed a local leakage resilient 2-out-of-n threshold secret sharing scheme with rate and leakage-resilience rate both $\Theta(1/n)$. They used this as a building block to construct non-malleable threshold secret sharing schemes secure against individual and joint tampering (where the adversary is allowed to jointly tamper sets of shares). A later paper also by Goyal and Kumar [GK18b] extended this to a compiler that adds non-malleability to a secret sharing scheme for any access structure. The non-malleable schemes resulting from both of these works, however, had rate tending to 0. Badrinarayanan and Srinivasan [BS18] later presented a compiler that converts any rate R secret sharing scheme to a non-malleable one for the same access struture with rate $\Theta(R/t \log^2 n)$, where t is the maximum size of any minimal set in the access structure. In the process, they constructed local leakage resilient k-out-of-n secret sharing schemes for a constant k that had rate $\Theta(1/\log(n))$ and leakage-resilience rate $\Theta(1/n \log(n))$.

Benhamouda et al. [BDIR18] were interested in studying the leakage-resilience of existing secret sharing schemes and MPC protocols. Inspired by the results of Guruswami and Wootters [GW16] that implied the possibility of recovering the secret from single-bit local leakage of Shamir shares over small characteristic fields, they investigated the leakage resilience of Shamir secret sharing over larger characteristic fields. They showed that, for large enough characteristic and large enough number of parties n, this scheme is leakage-resilient (with leakage-resilience rate close to 1/4) as long as the threshold is large (at least $n - o(\log(n))$). They used this fact to show leakage-resilience of the GMW protocol [GMW87] (using Beaver's triples), and to show an impossibility result for multi-party share conversion.

Boyle et al. [BGK14] define and construct leakage-resilient verifiable secret sharing schemes where the sharing and reconstruction are performed by interactive protocols (as oppoosed to just algorithms). They also show that a modification of the Shamir secret sharing scheme satisfies a weaker notion of leakage-resilience than the one we consider here, where it is only required that a random secret retain sufficient entropy given the leakage on the shares.

Dziembowski and Pietrzak [DP07] construct secret sharing schemes (that they call intrusion-resilient) that are resilient to adaptive leakage where the adversary is allowed to iteratively ask for leakage from different shares. Their reconstruction procedure is also interactive, however, requiring as many rounds of interaction as the adaptivity of the leakage tolerated.

Leakage-resilience of secure multiparty computation has been studied in the past in various settings [BGJK12, GIM+16, DHP11]. More broadly, leakage-resilience of various cryptographic primitives have been quite widely studied – we refer the reader to the survey by Alwen et al [ADW09] and the references therein. The notion of leakage tolerance was introduced by Garg et al [GJS11] and Bitansky et al [BCH12], and has been the subject of many papers since [BCG+11,BGJ+13,BDL14].

Secure multiparty computation with general interaction patterns was first studied by Halevi et al [HIJ+16], who showed a reduction from general interaction patterns to the star pattern (which is what we base our reduction on). They then showed an inefficient information-theoretically secure protocol for general functions, and an efficient one for symmetric functions. They also showed a computationally secure protocol for general functions assuming the existence of indistinguishability obfuscation and one-way functions, and for symmetric functions under an assumption about multilinear maps.

2 Preliminaries

Notation. We use capital letters to denote distributions and their support, and corresponding lowercase letters to denote a sample from the same. Let $[n]$ denote the set $\{1, 2, \ldots, n\}$ and U_r denote the uniform distribution over $\{0,1\}^r$. For a finite set S, we denote $x \xleftarrow{\$} S$ as sampling x uniformly at random from the set S. For any $i \in [n]$, let x_i denote the symbol at the i-th co-ordinate of x, and for any $T \subseteq [n]$, let $x_T \in \{0,1\}^{|T|}$ denote the projection of x to the co-ordinates indexed by T. We write $\circ$ to denote concatenation.

Standard definitions of min-entropy and statistical distance are given below.

Definition 2.1 (Min-entropy) *The min-entropy of a source X is defined to be*

$$H_\infty(X) = \min_{s \in \text{support}(X)} \{\log(1/\Pr[X = s])\}$$

A (n,k)-source is a distribution on $\{0,1\}^n$ with min-entropy k.

Definition 2.2 (Statistical distance) *Let D_1 and D_2 be two distributions on a set S. The statistical distance between D_1 and D_2 is defined to be:*

$$|D_1 - D_2| = \max_{T \subseteq S} |D_1(T) - D_2(T)| = \frac{1}{2} \sum_{s \in S} |\Pr[D_1 = s] - \Pr[D_2 = s]|$$

D_1 is ε-close to D_2 if $|D_1 - D_2| \leq \varepsilon$.

We will use the notation $D_1 \approx_\varepsilon D_2$ to denote that the statistical distance between D_1 and D_2 is at most ε.

Lemma 2.3 (Triangle Inequality) *If $D_1 \approx_{\varepsilon_1} D_2$ and $D_2 \approx_{\varepsilon_2} D_3$ then $D_1 \approx_{\varepsilon_1+\varepsilon_2} D_3$.*

We now recall the definition of (average) conditional min-entropy [DORS08].

Definition 2.4 ([DORS08]) *The average conditional min-entropy is defined as*

$$\widetilde{H}_\infty(X|W) = \log\left(E_{w \leftarrow W}\left[\max_x \Pr[X = x | W = w]\right]\right) = -\log E\left[2^{-H_\infty(X|W=w)}\right]$$

We recall some results on conditional min-entropy from [DORS08].

Lemma 2.5 ([DORS08]) *If a random variable B can take at most ℓ values, then $\widetilde{H}_\infty(A|B) \geq H_\infty(A) - \log \ell$.*

Seeded Extractors. We now recall the definition of a strong seeded extractor.

Definition 2.6 (Strong seeded extractor) *A function $\text{Ext} : \{0,1\}^n \times \{0,1\}^d \to \{0,1\}^m$ is called a strong seeded extractor for min-entropy k and error ε if for any (n,k)-source X and an independent uniformly random string U_d, we have*

$$|\text{Ext}(X, U_d) \circ U_d - U_m \circ U_d| < \varepsilon,$$

where U_m is independent of U_d. Further if the function $\text{Ext}(\cdot, u)$ is a linear function over $\mathbb{F}_2$ for every $u \in \{0,1\}^d$, then Ext is called a linear seeded extractor.

An average case seeded extractor requires that if a source X has average case conditional min-entropy $\tilde{H}_\infty(X|Z) \geq k$ then the output of the extractor is uniform even when Z is given. We recall the following lemma from [DORS08] which states that every strong seeded extractor is also an average-case strong extractor.

Lemma 2.7 ([DORS08]) *For any $\delta > 0$, if* Ext *is a (k, ε)-strong seeded extractor then it is also a $(k + \log\left(\frac{1}{\delta}\right), \varepsilon + \delta)$ average case strong extractor.*

2.1 Secret Sharing Scheme

We first give the definition of a k-monotone access structure, then define a sharing function and finally define a secret sharing scheme.

Definition 2.8 (k-Monotone Access Structure) *An access structure $\mathcal{A}$ is said to be monotone if for any set $S \in \mathcal{A}$, any superset of S is also in $\mathcal{A}$. We will call a monotone access structure $\mathcal{A}$ as k-monotone if for any $S \in \mathcal{A}$, $|S| \geq k$.*

Definition 2.9 (Sharing Function [Bei11]) *Let $[n] = \{1, 2, \ldots, n\}$ be a set of identities of n parties. Let $\mathcal{M}$ be the domain of secrets. A sharing function* Share *is a randomized mapping from $\mathcal{M}$ to $\mathcal{S}_1 \times \mathcal{S}_2 \times \ldots \times \mathcal{S}_n$, where $\mathcal{S}_i$ is called the domain of shares of party with identity i. A dealer distributes a secret $m \in \mathcal{M}$ by computing the vector $\mathsf{Share}(m) = (\mathsf{S}_1, \ldots, \mathsf{S}_n)$, and privately communicating each share S_i to the party i. For a set $T \subseteq [n]$, we denote $\mathsf{Share}(m)_T$ to be a restriction of $\mathsf{Share}(m)$ to its T entries.*

Definition 2.10 ($(\mathcal{A}, n, \varepsilon_c, \varepsilon_s)$-Secret Sharing Scheme [Bei11]) *Let $\mathcal{M}$ be a finite set of secrets, where $|\mathcal{M}| \geq 2$. Let $[n] = \{1, 2, \ldots, n\}$ be a set of identities (indices) of n parties. A sharing function* Share *with domain of secrets $\mathcal{M}$ is a $(\mathcal{A}, n, \varepsilon_c, \varepsilon_s)$-secret sharing scheme with respect to monotone access structure $\mathcal{A}$ if the following two properties hold :*

- ***Correctness:*** *The secret can be reconstructed by any set of parties that are part of the access structure $\mathcal{A}$. That is, for any set $T \in \mathcal{A}$, there exists a deterministic reconstruction function $\mathsf{Rec} : \otimes_{i \in T} \mathcal{S}_i \to \mathcal{M}$ such that for every $m \in \mathcal{M}$,*

  $$\Pr[\mathsf{Rec}(\mathsf{Share}(m)_T) = m] = 1 - \varepsilon_c$$

  where the probability is over the randomness of the Share *function. We will slightly abuse the notation and denote* Rec *as the reconstruction procedure that takes in $T \in \mathcal{A}$ and $\mathsf{Share}(m)_T$ as input and outputs the secret.*

- ***Statistical Privacy:*** *Any collusion of parties not part of the access structure should have "almost" no information about the underlying secret. More formally, for any unauthorized set $U \subseteq [n]$ such that $U \notin \mathcal{A}$, and for every pair of secrets $m_0, m_1 \in M$, for any distinguisher D with output in $\{0, 1\}$, the following holds :*

  $$|\Pr[D(\mathsf{Share}(m_0)_U) = 1] - \Pr[D(\mathsf{Share}(m_1)_U) = 1]| \leq \varepsilon_s$$

We define the rate of the secret sharing scheme as $\lim_{|m| \to \infty} \frac{|m|}{\max_{i \in [n]} |\mathsf{Share}(m)_i|}$

Remark 2.11 (Threshold Secret Sharing Scheme) *For ease of notation, we will denote a t-out-of-n threshold secret sharing scheme as $(t, n, \varepsilon_c, \varepsilon_s)$-secret sharing scheme.*

3 Leakage Resilient Secret Sharing Scheme

In this section, we will define and construct a leakage resilient secret sharing scheme against a class of local leakage functions.

3.1 Definition

We first recall the definition of a leakage resilient secret sharing scheme from [GK18a].

Definition 3.1 (Leakage Resilient Secret Sharing [GK18a]) *A $(\mathcal{A}, n, \varepsilon_c, \varepsilon_s)$ secret sharing scheme $(\mathsf{Share}, \mathsf{Rec})$ for message space $\mathcal{M}$ is said to be ε-leakage resilient against a leakage family $\mathcal{F}$ if for all functions $f \in \mathcal{F}$ and for any two messages $m_0, m_1 \in \mathcal{M}$:*

$$|f(\mathsf{Share}(m_0)) - f(\mathsf{Share}(m_1))| \leq \varepsilon$$

3.2 Local Leakage Resilience

In this subsection, we will transform any secret sharing scheme to a leakage resilient secret sharing scheme against the local leakage function family. We first recall the definition of this function family.

Local Leakage Function Family. Let $\mathsf{Share} : \mathcal{M} \to \mathcal{S}_1 \times \mathcal{S}_2 \ldots \times \mathcal{S}_n$. We are interested in constructing leakage resilient secret sharing schemes against the specific function family $\mathcal{F}_{\mathcal{A},\mu} = \{f_{K,\vec{\tau}} : K \subseteq [n], K \notin \mathcal{A}, \tau_i : \mathcal{S}_i \to \{0,1\}^\mu\}$ where $f_{K,\vec{\tau}}$ on input $(\mathsf{share}_1, \ldots, \mathsf{share}_n)$ outputs share_i for each $i \in K$ in the clear and outputs $\tau_i(\mathsf{share}_i)$ for every $i \in [n] \setminus K$. Following [BDIR18], we will call secret sharing schemes resilient to $\mathcal{F}_{\mathcal{A},\vec{\tau}}$ as local leakage resilient secret sharing. We will define the leakage-rate of such secret sharing schemes to be $\lim_{\mu\to\infty} \frac{\mu}{\max_{i\in[n]} |\mathsf{Share}(m)_i|}$.

Remark 3.2 *We remark that definition 3.1 is satisfiable against the leakage function class $\mathcal{F}_{\mathcal{A},\mu}$ (for any $\mu > 0$) only if the access structure is 2-monotone (see Definition 2.8). Hence, in the rest of the paper, we will concentrate on 2-monotone access structures.*

3.2.1 Description of the Compiler

We will give a compile that takes any $(\mathcal{A}, n, \varepsilon_c, \varepsilon_s)$ secret sharing scheme for any 2-monotone $\mathcal{A}$ and outputs a local leakage resilient secret sharing scheme for $\mathcal{A}$. We give the description of the compiler in Figure 1 and give an informal overview below.

Overview. The sharing function does the following. On input a secret m, it first shares m using the underlying secret sharing scheme for the access structure $\mathcal{A}$ to get the set of shares $(\mathsf{Sh}_1, \ldots, \mathsf{Sh}_n)$. It chooses an uniform seed s and for each $i \in [n]$, chooses an uniform source w_i for the underlying strong seeded extractor Ext. For each $i \in [n]$, it masks Sh_i using the output of the extractor on the source w_i using the seed s to obtain the masked shares Sh'_i. It then chooses an uniform string r and computes $S'_i = \mathsf{Sh}'_i \oplus r$. It shares (s, r) using a 2-out-of-n Shamir secret sharing scheme to obtain the set of shares $(S_1, \ldots, S_n)$. The share corresponding to party i is the triple $\mathsf{share}_i = (w_i, S'_i, S_i)$. The reconstruction function first obtains (s, r) from the S_i's. Notice that since $\mathcal{A}$ is 2-monotone each authorized set has atleast two shares and we can reconstruct (s, r) from any

authorized set of shares. It then computes $\mathsf{Sh}_i = S'_i \oplus r \oplus \mathrm{Ext}(w_i, s)$. It finally reconstructs m from the Sh_i's.

Let $(\mathsf{Share}, \mathsf{Rec})$ be a $(\mathcal{A}, n, \varepsilon_c, \varepsilon_s)$ secret sharing scheme for sharing secrets from $\mathcal{M}$ with share size equal to ρ bits. Let $(\mathsf{Share}_{(2,n)}, \mathsf{Rec}_{(2,n)})$ be a 2-out-of-n Shamir Secret sharing. Let $\mathrm{Ext} : \{0,1\}^\eta \times \{0,1\}^d \to \{0,1\}^\rho$ be a $(\eta - \mu, \varepsilon)$-average-case, strong seeded extractor.

$\mathsf{LRShare}$: To share a secret $m \in \mathcal{M}$:

1. Run $\mathsf{Share}(m)$ to obtain the shares $(\mathsf{Sh}_1, \ldots, \mathsf{Sh}_n)$.
2. Choose an uniform seed $s \xleftarrow{\$} \{0,1\}^d$ and a masking string $r \xleftarrow{\$} \{0,1\}^\rho$.
3. For each $i \in [n]$ do:
   (a) Choose $w_i \xleftarrow{\$} \{0,1\}^\eta$.
   (b) Set $\mathsf{Sh}'_i = \mathsf{Sh}_i \oplus \mathrm{Ext}(w_i, s)$.
4. Run $\mathsf{Share}_{(2,n)}(s, r)$ to obtain $S_1, \ldots, S_n$.
5. Output share_i as $(w_i, \mathsf{Sh}'_i \oplus r, S_i)$.

LRRec : Given the shares $\mathsf{share}_{j_1}, \mathsf{share}_{j_2}, \ldots, \mathsf{share}_{j_\ell}$ where $K = \{j_1, \ldots, j_k\} \in \mathcal{A}$ do:

1. For each $i \in K$, parse share_i as (w_i, S'_i, S_i).
2. Run $\mathsf{Rec}_{(2,n)}(S_{j_1}, S_{j_2})$ to recover (s, r)
3. For each $i \in K$ do:
   (a) Compute $\mathsf{Sh}'_i = S'_i \oplus r$.
   (b) Recover Sh_i by computing $\mathsf{Sh}'_i \oplus \mathrm{Ext}(w_i, s)$.
4. Run $\mathsf{Rec}(\mathsf{Sh}_{j_1}, \ldots, \mathsf{Sh}_{j_k})$ to recover the secret m.

Figure 1: Local Leakage-Resilient Secret Sharing

Theorem 3.3 *Consider any 2-monotone access structure $\mathcal{A}$ and $\mu \in \mathbb{N}$ and a secret domain $\mathcal{M}$ with secrets of length m. Suppose for some $\eta, d, \rho \in \mathbb{N}$ and $\varepsilon_c, \varepsilon_s, \varepsilon \in [0,1)$, the following exist:*

- *A $(\mathcal{A}, n, \varepsilon_c, \varepsilon_s)$ secret sharing scheme for the secret domain $\mathcal{M}$ with share length ρ.*
- *A $(\eta - \mu, \varepsilon)$-average-case strong seeded extractor* $\mathrm{Ext} : \{0,1\}^\eta \times \{0,1\}^d \to \{0,1\}^\rho$.

Then, the construction in Figure 1, when instantiated with these, is a $(\mathcal{A}, n, \varepsilon_c, \varepsilon_s)$ secret sharing scheme for $\mathcal{M}$ that is $2(\varepsilon_s + n \cdot \varepsilon)$-leakage resilient against $\mathcal{F}_{\mathcal{A},\mu}$. It has share size $(\eta + 2\rho + d)$.

Proof We note that correctness follows directly from the correctness of $(\mathsf{Share}, \mathsf{Rec})$ and that of Shamir's secret sharing. We will now argue leakage resilience and privacy will follow directly from this argument.

Leakage Resilience. Let us fix a function $f_{K,\vec{\tau}} \in \mathcal{F}_{\mathcal{A},\mu}$ where $K \notin \mathcal{A}$. Recall that $f_{K,\vec{\tau}}$ on input $(\mathsf{share}_1, \ldots, \mathsf{share}_n)$ will output share_i in the clear for every $i \in K$ and for all other i, it will output $\tau_i(\mathsf{share}_i)$. We need to show that for any two secrets $m, m' \in \mathcal{M}$:

$$|f_{K,\vec{\tau}}(\mathsf{LRShare}(m)) - f_{K,\vec{\tau}}(\mathsf{LRShare}(m'))| \leq 2(\varepsilon_s + n \cdot \varepsilon)$$

The proof strategy is as follows. We start with the distribution wherein we will output $f_{K,\vec{\tau}}(\mathsf{Share}(m))$. Then, for each $i \notin K$, we will choose R_i randomly from $\{0,1\}^\rho$ and run the function τ_i with input $(w_i, R_i \oplus r, S_i)$ instead of $(w_i, \mathsf{Sh}'_i \oplus r, S_i)$. We will show that the output of $f_{K,\vec{\tau}}$ on the modified input is statistically close to the real world leakage by relying on the extractor property. Once we have replaced all Sh_i with a random string for $i \notin K$, we can rely on the privacy of secret sharing scheme $(\mathsf{Share}, \mathsf{Rec})$ to argue that the message is statistically hidden given the leakage. We will now formalize this argument.

Let us consider some total ordering $\prec$ of the elements in the set $[n] \setminus K$. We define a sequence of hybrids Hyb_i for every $i \in [n] \setminus K$ where we use the modified sharing procedure $\mathsf{LRShare}'_i$ described below.

Description of $\mathsf{LRShare}'_i$:

1. Run $\mathsf{Share}(m)$ to obtain the shares $(\mathsf{Sh}_1, \ldots, \mathsf{Sh}_n)$.
2. Choose an uniform seed $s \xleftarrow{\$} \{0,1\}^d$ and a masking element $r \xleftarrow{\$} \{0,1\}^\rho$.
3. For each $j \in [n]$ do:
   (a) Choose $w_j \xleftarrow{\$} \{0,1\}^\eta$.
   (b) Choose $\mathsf{Sh}'_j \leftarrow \{0,1\}^\rho$ if $j \in [n] \setminus K$ and $j \prec i$. Else, set $\mathsf{Sh}'_j = \mathsf{Sh}_j \oplus \mathrm{Ext}(w_j, s)$.
4. Run $\mathsf{Share}_{(2,n)}(s, r)$ to obtain $S_1, \ldots, S_n$.
5. Output share_i as $(w_i, \mathsf{Sh}'_i \oplus r, S_i)$.

The output of Hyb_i is $f_{K,\vec{\tau}}(\mathsf{share}_1, \ldots, \mathsf{share}_n)$. Let $\ell = |[n] \setminus K|$ and let i_1 be the first element and i_ℓ to be the last element in $[n] \setminus K$ as per the ordering $\prec$. Notice that in Hyb_{i_1}, the distribution of the shares given as input to $f_{K,\vec{\tau}}$ is identical to a valid secret sharing of m. We first prove the following claim.

Claim 3.4 *For every $i, i' \in [n] \setminus K$ such that i' is the successor of i as per the ordering $\prec$, we have $\mathsf{Hyb}_{i'} \approx_\varepsilon \mathsf{Hyb}_i$.*

Proof Assume for the sake of contradiction that the statistical distance between $\mathsf{Hyb}_{i'}$ and Hyb_i is greater than ε. We will use this to break the property of the strong, average-case seeded extractor Ext. The reduction is described below.

The reduction runs $\mathsf{Share}(m)$ to obtain $(\mathsf{Sh}_1, \ldots, \mathsf{Sh}_n)$. It then chooses a random string $r_i \xleftarrow{\$} \{0,1\}^\rho$ and a random Shamir share S_i. It then defines a function $f_{r_i,S_i} : \{0,1\}^\eta \to \{0,1\}^\mu$ as follows: on input w_i, run $\tau_i(w_i, r_i, S_i)$ and output whatever it outputs. Note that since the output length of f_{r_i,S_i} is μ bits, it follows from Lemma 2.5 that for a randomly chosen $w_i \xleftarrow{\$} \{0,1\}^\eta$, $\widetilde{H}_\infty(w_i | f_{r_i,S_i}(w_i)) \geq \eta - \mu$. The reduction receives from the extractor challenger $(s, f_{r_i,S_i}(w_i), R_i)$ where either $R_i = \mathrm{Ext}(w_i, s)$ or R_i is chosen uniformly at random. It sets $\mathsf{Sh}'_i = \mathsf{Sh}_i \oplus R_i$ and sets

$r = r_i \oplus \mathsf{Sh}'_i$. It then generates Shamir secret sharing of (s, r) such that it is consistent with the share S_i to obtain $S_1, \ldots, S_n$. For every other $j \neq i$, it chooses $w_j \xleftarrow{\$} \{0,1\}^\eta$, and chooses $\mathsf{Sh}'_j \xleftarrow{\$} \{0,1\}^\rho$ if $j \prec i$ and $j \in [n] \setminus K$; else, it sets $\mathsf{Sh}'_j = \mathsf{Sh}_j \oplus \mathrm{Ext}(w_j, s)$. For every $k \in K$, the reduction outputs $(w_k, \mathsf{Sh}'_k \oplus r, S_k)$ and for every other $j \in [n] \setminus K \cup \{i\}$, it outputs $\tau_j(w_j, \mathsf{Sh}'_j \oplus r, S_j)$. For $j = i$, it outputs $f_{r_i, S_i}(w_i)$.

Notice that if R_i is $\mathrm{Ext}(w_i, s)$ then the output of the reduction is identical to Hyb_i. Else, it is identical to $\mathsf{Hyb}_{i'}$. Thus, the statistical distance between $\mathsf{Hyb}_{i'}$ and Hyb_i is more than ε implies that this reduction can break the extractor property which is a contradiction. ■

By repeated application of Claim 3.4, we infer that $\mathsf{Hyb}_{i_1} \approx_{\ell\varepsilon} \mathsf{Hyb}_{i_\ell}$. We note that Hyb_{i_ℓ} is ε_s-close to another distribution $\mathcal{D}$ where $(\mathsf{Sh}_1, \ldots, \mathsf{Sh}_n)$ are generated as $\mathsf{Share}(0)$ instead of $\mathsf{Share}(m)$ and this follows directly from the privacy property of the $(\mathsf{Share}, \mathsf{Rec})$. The proof of theorem now follows from observing that $\mathcal{D} \approx_{\varepsilon_s + \ell\varepsilon} f_{K, \overrightarrow{\tau}}(\mathsf{LRShare}(m'))$.

■

3.2.2 Instantiation

We first recall the explicit construction of strong seeded extractors from the work of Guruswami, Umans and Vadhan [GUV09].

Theorem 3.5 ([GUV09]) *For any constant $\alpha > 0$, and all integers $n, k > 0$ there exists a polynomial time computable (k, ε)-strong seeded extractor* $\mathrm{Ext} : \{0,1\}^n \times \{0,1\}^d \to \{0,1\}^m$ *with $d = O(\log n + \log(\frac{1}{\varepsilon}))$ and $m = (1 - \alpha)k$.*

We now instantiate our scheme with the following building blocks:

- Let $(\mathsf{Share}, \mathsf{Rec})$ be a secret sharing scheme for a 2-monotone access structure $\mathcal{A}$ for sharing m-bit messages with rate R.
- We use the Guruswami, Umans and Vadhan [GUV09] strong seeded extractor (refer Theorem 3.5). We set $n = 1.01m/R + \log(1/\varepsilon) + \mu$ and $d = O(\log n + \log(1/\varepsilon))$ and from Theorem 3.5 and Lemma 2.7, it follows that Ext is a $(1.01m/R + \log(1/\varepsilon), 2\varepsilon)$ average-case, strong seeded extractor.

Thus, we get $|\mathsf{share}_i| = |w_i| + |\mathsf{Sh}_i| + |S_i| = n + m/R + (m/R + d) = 3.01m/R + \mu + O(\log m + \log \mu + \log 1/\varepsilon)$.

Corollary 3.6 *If there exists a secret sharing scheme for a 2-monotone access structure with rate R, then there exists an ε-local leakage resilient secret sharing for $\mathcal{A}$ for some negligible ε with rate $R/3.01$.*

For the special case of threshold secret sharing scheme for which we know constructions with rate 1 [Sha79], we obtain the following corollary.

Corollary 3.7 *For any $n, t, \mu \in \mathbb{N}$ such that $t \leq n$, and $\varepsilon \in (0, 1)$, there is a t-out-of-n threshold secret sharing scheme that is $(2n\varepsilon)$-leakage resilient against $\mathcal{F}_{t-1,\mu}$, and has rate $O(1)$, and leakage-resilience rate 1.*

3.3 Strong Local Leakage Resilience

In this subsection, we consider a stronger notion of leakage resilience for secret sharing, in which the leakage on the "honest" shares is allowed to depend arbitrarily on the "corrupted" shares – this is meant to capture a scenario where an adversary first learns the shares of t of the n parties, and then specifies leakage functions that are applied to the remaining $(n-t)$ shares, the outputs of which are then given to the adversary. This corresponds to leakage resilience against the function family described below. (For simplicity, we will describe our results in this subsection only for threshold access structures, but they can be generalized to all access structures in a straightforward manner.)

Strong Local Leakage Function Family. Let $\mathsf{Share} : \mathcal{M} \to \mathcal{S}_1 \times \cdots \times \mathcal{S}_n$ be the sharing algorithm of the scheme under consideration, and $t \in [n]$ and μ be natural numbers. A semi-local leakage function family is parametrized by three numbers t (the adaptivity threshold), t' (the corruption threshold), and μ (the amount of leakage), such that $t \leq t'$. The family $\mathcal{H}_{t,t',\mu}$ consists of functions $\{h_{T,T',\vec{\tau}}\}$, where the subsets $T \subseteq T' \subseteq [n]$ are such that $|T| = t$ and $|T'| = t'$; and for $i \in [n] \setminus T'$, the function τ_i takes inputs from $(\mathcal{S}_{i_1} \times \cdots \times \mathcal{S}_{i_t}) \times \mathcal{S}_i$ (where $T = \{i_1, \ldots, i_t\}$), and outputs μ bits. The function $h_{T,\vec{\tau}}$, when given input $(\mathsf{share}_1, \ldots, \mathsf{share}_n)$, outputs share_i for each $i \in T'$, and $\tau_i((\mathsf{share}_{i_1}, \ldots, \mathsf{share}_{i_t}), \mathsf{share}_i)$ for $i \notin T'$. A secret sharing scheme resilient to leakage by such function families is said to be *strongly local leakage resilient* against $\mathcal{H}_{t,t',\mu}$.

Game-based Definition. Strong local leakage resilience may alternatively, and perhaps more naturally, be defined as the inability of the adversary to guess the bit b correctly in the following game. The adversary picks a set T of size at most t and gives two messages m_0 and m_1 to the challenger. The challenger tosses a random coin b, secret shares m_b and gives the shares corresponding to the set T to the adversary. Depending on this set, the adversary chooses a local leakage function that operates on the rest of the $(n-t)$ shares. This function outputs some set of shares of size $t'-t$ in the clear and outputs a bounded leakage on the remaining shares. The challenger provides the output of this local leakage function to the adversary and it guesses the bit b. Formally, this is defined as follows:

1. The adversary selects the sets $T \subseteq T' \subseteq [n]$ such that $|T| = t$ and $|T'| = t'$. It then picks messages $m_0, m_1 \in \mathcal{M}$, and sends all of these to the challenger.
2. The challenger picks a random bit b and computes $(\mathsf{share}_1, \ldots, \mathsf{share}_n) \leftarrow \mathsf{LRShare}(m_b)$. It sends share_T to the adversary.
3. The adversary now chooses a local leakage function $f_{(T' \setminus T),\mu}$ that operates on the $(n-t)$ shares $(\mathsf{share}_i)_{i \notin T}$. It sends this to the challenger.
4. The challenger sends the leakage $f_{(T' \setminus T),\mu}((\mathsf{share}_i)_{i \notin T})$.
5. The adversary outputs a guess b' for b.

We require that $\Pr[b = b'] = 1/2 + \text{negl}$.

We will show that a modification of the construction from Section 3.2.1 can achieve strong local leakage resilience.

Theorem 3.8 *Consider any $n, t, \mu \in \mathbb{N}$ such that $t \leq n$ and a secret domain $\mathcal{M}$. Suppose for some $\eta, d, R \in \mathbb{N}$ and $\varepsilon \in [0, 1)$, the following exist:*

Let $(\mathsf{Share}_{(t,n)}, \mathsf{Rec}_{(t,n)})$ represent a t-out-of-n threshold secret sharing scheme for secrets in an unspecified domain; let ρ be the bit-length of each share under this scheme when the secret is from the secret domain $\mathcal{M}$. Let η and d be such that there is a (k, ε)-average-case strong seeded extractor $\mathrm{Ext} : \{0,1\}^\eta \times \{0,1\}^d \to \{0,1\}^\rho$ that outputs ρ bits, where $k = (\eta - \mu)$.

$\mathsf{LRShare}$: To share a secret $m \in \mathcal{M}$:

1. Run $\mathsf{Share}_{(t,n)}(m)$ to obtain the shares $(\mathsf{Sh}_1, \ldots, \mathsf{Sh}_n)$.
2. Choose a uniform seed $s \xleftarrow{\$} \{0,1\}^d$.
3. For each $i \in [n]$ do:
   (a) Choose $w_i \xleftarrow{\$} \{0,1\}^\eta$.
   (b) Choose a masking string $r_i \xleftarrow{\$} \{0,1\}^\rho$.
   (c) Set $\mathsf{Sh}'_i = \mathsf{Sh}_i \oplus \mathrm{Ext}(w_i, s) \oplus r_i$.
   (d) Run $\mathsf{Share}_{(t,n)}(r_i)$ to obtain $r_{(i,1)}, \ldots, r_{(i,n)}$.
4. Run $\mathsf{Share}_{(t,n)}(s)$ to obtain $S_1, \ldots, S_n$.
5. Output share_i as $(w_i, \mathsf{Sh}'_i, S_i, (r_{(1,i)}, \ldots, r_{(n,i)}))$.

LRRec : Given any set of t shares $\mathsf{share}_{i_1}, \mathsf{share}_{i_2}, \ldots, \mathsf{share}_{i_t}$, do:

1. For each i_j, parse share_{i_j} as $(w_{i_j}, S'_{i_j}, S_{i_j}, (r_{(1,i_j)}, \ldots, r_{(n,i_j)}))$.
2. Run $\mathsf{Rec}_{(t,n)}(S_{i_1}, \ldots, S_{i_t})$ to recover s.
3. For each i_j, do:
   (a) Run $\mathsf{Rec}_{(t,n)}(r_{(i_j,i_1)}, \ldots, r_{(i)j,i_t)})$ to recover r_{i_j}.
   (b) Recover Sh_{i_j} by computing $S'_{i_j} \oplus \mathrm{Ext}(w_{i_j}, s) \oplus r_{i_j}$.
4. Run $\mathsf{Rec}(\mathsf{Sh}_{i_1}, \ldots, \mathsf{Sh}_{i_t})$ to recover the secret m.

Figure 2: Strongly Local Leakage-Resilient Secret Sharing

- *A perfect t-out-of-n Shamir secret sharing scheme with share size ρ for secrets in $\mathcal{M}$.*
- *A $(\eta - \mu, \varepsilon)$-average-case strong seeded extractor* $\mathrm{Ext} : \{0,1\}^\eta \times \{0,1\}^d \to \{0,1\}^\rho$.

Then, the construction in Figure 2, when instantiated with these, is a t-out-of-n threshold secret sharing scheme for $\mathcal{M}$ that is $(2n\varepsilon)$-leakage resilient against $\mathcal{H}_{(t-2),(t-1),\mu}$. It has share size $(\eta + \rho + d + n\rho)$.

Using the same instantiations as in Section 3.2.2, we get the following.

Corollary 3.9 *For any $n, t, \mu \in \mathbb{N}$ such that $t \leq n$, and $\varepsilon \in [0,1]$, there is a t-out-of-n threshold secret sharing scheme that is $(2n\varepsilon)$-leakage resilient against $\mathcal{H}_{t-2,\mu}$, and has rate $\Omega(1/n)$, and leakage-resilience rate 1.*

We now prove Theorem 3.8 along the same lines as Theorem 3.3.

Proof of Theorem 3.8: Correctness of the scheme in Figure 2 follows immediately from the correctness of the underlying threshold secret sharing scheme. Privacy also follows from the privacy of the underlying scheme by the observation that information about Sh_i is present only in the ith share. The size of each share may also be verified by inspection. In the remainder of the proof, we argue leakage resilience against $\mathcal{H}_{(t-2),(t-1),\mu}$.

Fix a function $h_{T,T',\vec{\tau}} \in \mathcal{H}_{(t-2),(t-1),\mu}$. Recall that $h_{T,T',\vec{\tau}}$ on input $(\mathsf{share}_1, \ldots, \mathsf{share}_n)$ will output share_i for every $i \in T'$ and for all other i, it will output $\tau_i(\mathsf{share}_T, \mathsf{share}_i)$. We need to show that for any two secrets $m, m' \in \mathcal{M}$:

$$|h_{T,T',\vec{\tau}}(\mathsf{LRShare}(m)) - h_{T,T',\vec{\tau}}(\mathsf{LRShare}(m'))| \leq 2n\varepsilon$$

The proof strategy is as follows. We construct a sequence of hybrid sharing procedures $\mathsf{LRShare}_0$, $\ldots, \mathsf{LRShare}_n$, where $\mathsf{LRShare}_1(m)$ is the same as $\mathsf{LRShare}(m)$, and the output of $\mathsf{LRShare}_n(m)$ contains no information about m. We will then show that the distributions of the output of $h_{T,T',\vec{\tau}}$ when run on the shares produced by two consecutive hybrids have statistical distance at most ε.

For each $i \notin T'$, in the ith hybrid, we will replace the $(\mathsf{Sh}_i \oplus \mathrm{Ext}(w_i, s) \oplus r_i)$ in the ith share of the previous hybrid with a random string of appropriate length. We will show that the output of $h_{T,T',\vec{\tau}}$ in either case is statistically close by relying on the properties of the extractor. Once we have replaced all Sh_i's with random strings for all $i \notin T'$, we can rely on the privacy of the secret sharing scheme $(\mathsf{Share}, \mathsf{Rec})$ to argue that the message is statistically hidden given the leakage. We will now formalize this argument.

Let us consider some total ordering $\prec$ of the elements in the set $[n] \setminus T'$. We define a sequence of hybrid distributions for every $i \in [0, n]$ where we use the modified sharing procedure $\mathsf{LRShare}_i$ described below. (We use notation that was set up in Figure 2.)

$\mathsf{LRShare}_i$: On input a secret $m \in \mathcal{M}$:

1. Run $\mathsf{Share}_{(t,n)}(m)$ to obtain the shares $(\mathsf{Sh}_1, \ldots, \mathsf{Sh}_n)$.
2. Choose a uniform seed $s \xleftarrow{\$} \{0,1\}^d$.
3. For each $j \in [n]$ do:
   (a) Choose $w_j \xleftarrow{\$} \{0,1\}^\eta$.
   (b) Choose a masking string $r_j \xleftarrow{\$} \{0,1\}^\rho$.
   (c) If $j \in [n] \setminus T$ and $j \prec i$ or $j = i$, choose $\mathsf{Sh}'_j \leftarrow \{0,1\}^\rho$.
   (d) Else, set $\mathsf{Sh}'_j = \mathsf{Sh}_j \oplus \mathrm{Ext}(w_j, s) \oplus r_j$.
   (e) Run $\mathsf{Share}_{(t,n)}(r_j)$ to obtain $r_{(j,1)}, \ldots, r_{(j,n)}$.
4. Run $\mathsf{Share}_{(t,n)}(s)$ to obtain $S_1, \ldots, S_n$.
5. Output share_i as $(w_i, \mathsf{Sh}'_i, S_i, (r_{(1,i)}, \ldots, r_{(n,i)}))$.

For any secret $m \in \mathcal{M}$, the hybrid Hyb_i^m is defined as the distribution of $h_{T,T',\vec{\tau}}(\mathsf{share}_1, \ldots, \mathsf{share}_n)$, where the shares are obtained from $\mathsf{LRShare}_i(m)$. Notice that in Hyb_0^m, the distribution of the shares given as input to $h_{T,T',\vec{\tau}}$ is identical to a valid secret sharing of m. We now prove the following claim.

Claim 3.10 *For any $m \in \mathcal{M}$ and every $i, i' \in [0, n] \setminus T'$ such that i' is the successor of i under the ordering $\prec$, we have $\mathsf{Hyb}_{i'}^m \approx_\varepsilon \mathsf{Hyb}_i^m$.*

Proof Assume for the sake of contradiction that the statistical distance between $\mathsf{Hyb}_{i'}^m$ and Hyb_i^m is greater than ε. We will use this to break the property of the strong, average-case seeded extractor Ext. The reduction is described below.

The reduction's task it to specify a leakage function h such that the distributions of the form $(s, h(w), z)$ where z is either $\mathrm{Ext}(w, s)$ or a random string, are statistically far. It will pick ℓ in such a way that it can use the "challenge" $(s, \ell(w), z)$ to sample a distribution that is the same as Hyb_i^m if z is an extractor output, and the same as $\mathsf{Hyb}_{i'}^m$ if z is a random string.

Initially, the reduction sets the shares share_j for $j \in T$ to be completely random – it picks w_j, Sh'_j, S_j and $r_{(\cdot,j)}$'s uniformly at random from the appropriate domains, and sets $\mathsf{share}_j = (w_j, \mathsf{Sh}'_j, S_j, (r_{(1,j)}, \ldots, r_{(n,j)}))$. Then, for i', it picks $\mathsf{Sh}'_{i'}$, $S_{i'}$ and the $r_{(\cdot,i')}$'s at random. It then defines the leakage function $\ell : \{0,1\}^\eta \to \{0,1\}^\mu$ that, on input w, outputs $\tau_{i'}(\mathsf{share}_T, (w, \mathsf{Sh}'_{i'}, S_{i'}, (r_{(\cdot,i')})))$.

The reduction now receives an extractor challenge $(s, \ell(w), z)$ where either $z = \mathrm{Ext}(w, s)$ or z is chosen uniformly at random. It will then set things up so that w is being implicitly used as $w_{i'}$ in a hybrid sharing scheme, and z in place of $\mathrm{Ext}(w_i, s)$.

To do so, it first runs $\mathsf{Share}_{(t,n)}(m)$ to obtain $(\mathsf{Sh}_1, \ldots, \mathsf{Sh}_n)$. For each $j \in T$, it sets r_j to be $\mathsf{Sh}'_j \oplus \mathsf{Sh}_j \oplus \mathrm{Ext}(w_j, s)$. For i', it sets $r_{i'}$ to be $\mathsf{Sh}'_{i'} \oplus \mathsf{Sh}_{i'} \oplus z$. If $j \notin T' \cup \{i'\}$ and $j \prec i'$, it sets w_j, Sh'_j and r_j to be random strings. For all other j (which are either in T' or satisfy $i' \prec j$) , it picks w_j and r_j at random, and sets Sh'_j to be $\mathsf{Sh}_j \oplus \mathrm{Ext}(w_j, s) \oplus r_j$. It then sets all the $r_{(j,j')}$'s and S_j's that have not been set so far in such a way that they form valid random sharings of the r_j's and s. This last part can be done because all of these sharings are under a t-out-of-n threshold sharing scheme, and until this point we have only determined at most $(t-1)$ of the shares of any of these quantities, as $|T|$ is at most $(t-2)$. Each share_j can now be set to be $(w_j, \mathsf{Sh}'_j, S_j, (r_{(1,j)}, \ldots, r_{(n,j)}))$.

Finally, to produce the output of $h_{T,T',\vec{\tau}}$, for every $j \in T'$, the reduction outputs share_j, for i' it outputs $\ell(w)$ as the output of $\tau_{i'}$, and for every other $j \notin T' \cup \{i'\}$, it computes and outputs $\tau_j(\mathsf{share}_T, (w_j, \mathsf{Sh}'_j, S_j, (r_{(\cdot,j)})))$. Notice that if z is $\mathrm{Ext}(w, s)$ then the output of the reduction is identical to Hyb_i^m; else, it is identical to $\mathsf{Hyb}_{i'}^m$.

Note that since the output length of ℓ is μ bits, it follows from Lemma 2.5 that for a randomly chosen $w \xleftarrow{\$} \{0,1\}^\eta$, $\widetilde{H}_\infty(w|\ell(w)) \geq \eta - \mu = k$. Thus, if Ext is a (k, ε)-strong average-case seeded extractor, the distance between the distributions $(s, \ell(w), z)$ with z being $\mathrm{Ext}(w, s)$ or random is at most ε. Thus, the statistical distance between $\mathsf{Hyb}_{i'}$ and Hyb_i being more than ε would imply that this reduction can break this extractor property, which is a contradiction. This proves the claim. ■

By repeated application of Claim 3.4, we infer that for any m, $\mathsf{Hyb}_0^m \approx_{n\varepsilon} \mathsf{Hyb}_n^m$. In computing Hyb_n^m, the shares in T' of the underlying threshold scheme are computed as they are supposed to be, but all the other shares are replaced with random strings. By the privacy of this underlying threshold secret sharing scheme, as T' is of size less than t, the distribution Hyb_n^m is independent of m, and is thus the same as $\mathsf{Hyb}_n^{m'}$ for any secret m', which is at most $n\varepsilon$ far from $\mathsf{Hyb}_0^{m'}$. Thus, $\mathsf{Hyb}_0^m \approx_{2n\varepsilon} \mathsf{Hyb}_0^{m'}$, and the theorem now follows from observing that Hyb_0^m is the same as $h_{T,T',\vec{\tau}}(\mathsf{LRShare}(m))$.

■

4 Rate Preserving Non-Malleable Secret Sharing

In this section, we will use the leakage resilient secret sharing scheme in Section 3 to construct a non-malleable secret sharing scheme. Specifically, we give a compiler that takes any secret sharing scheme for a 4-monotone access structure (see Definition 2.8) with rate R and converts it into a non-malleable secret sharing scheme for the same access structure with rate $O(R)$. In appendix , we give some background on non-malleable codes.

4.1 Definition

We recall the definition of non-malleable secret sharing for a monotone access structure $\mathcal{A}$.

Definition 4.1 (Non-Malleable Secret Sharing for General Access Structures [GK18b]) *Let* $(\mathsf{Share}, \mathsf{Rec})$ *be a* $(\mathcal{A}, n, \varepsilon_c, \varepsilon_s)$*-secret sharing scheme for message space* $\mathcal{M}$ *and access structure* $\mathcal{A}$*. Let* $\mathcal{F}$ *be a family of tampering functions. For each* $f \in \mathcal{F}$*,* $m \in \mathcal{M}$ *and authorized set* $T \in \mathcal{A}$*, define the tampered distribution* $\mathsf{Tamper}_m^{f,T}$ *as* $\mathsf{Rec}(f(\mathsf{Share}(m))_T)$ *where the randomness is over the sharing function* Share*. We say that the* $(\mathcal{A}, n, \varepsilon_c, \varepsilon_s)$*-secret sharing scheme,* $(\mathsf{Share}, \mathsf{Rec})$ *is* ε'*-non-malleable w.r.t.* $\mathcal{F}$ *if for each* $f \in \mathcal{F}$ *and any authorized set* $T \in \mathcal{A}$*, there exists a distribution* $D^{f,T}$ *over* $\mathcal{M} \cup \{\mathsf{same}^\star\}$ *such that for any* m*,*

$$|\mathsf{Tamper}_m^{f,T} - \mathrm{copy}(D^{f,T}, m)| \leq \varepsilon'$$

where copy *is defined by* $\mathrm{copy}(x, y) = \begin{cases} x & \text{if } x \neq \mathsf{same}^\star \\ y & \text{if } x = \mathsf{same}^\star \end{cases}$.

4.2 Construction

We give a construction of a non-malleable secret sharing scheme for a 4-monotone access structures against the individual tampering function family $\mathcal{F}_{\mathsf{ind}}$ (see below).

Individual Tampering Family $\mathcal{F}_{\mathsf{ind}}$. Let Share be the sharing function of the secret sharing scheme that outputs n-shares in $\mathcal{S}_1 \times \mathcal{S}_2 \ldots \times \mathcal{S}_n$. The function family $\mathcal{F}_{\mathsf{ind}}$ is composed of functions $(f_1, \ldots, f_n)$ where each $f_i : \mathcal{S}_i \to \mathcal{S}_i$.

Construction. The construction is same as the one given in [BS18] but we instantiate the leakage-resilient secret sharing scheme with the one constructed in the previous section. We now give the description of the building blocks and then give the construction. In the following, we will denote a t-out-of-n monotone access structure as (t, n).

Building Blocks. The construction uses the following building blocks. We instantiate them with concrete schemes later:

- A 3-split-state non-malleable code $(\mathsf{Enc}, \mathsf{Dec})$ where $\mathsf{Enc} : \mathcal{M} \to \mathcal{L} \times \mathcal{C} \times \mathcal{R}$ and the simulation error of the scheme is ε_1. Furthermore, we assume that for any two messages $m, m' \in \mathcal{M}$, $(\mathsf{C}, \mathsf{R}) \approx_{\varepsilon_2} (\mathsf{C}', \mathsf{R}')$ where $(\mathsf{L}, \mathsf{C}, \mathsf{R}) \leftarrow \mathsf{Enc}(m)$ and $(\mathsf{L}', \mathsf{C}', \mathsf{R}') \leftarrow \mathsf{Enc}(m')$.

- A $(\mathcal{A}, n, \varepsilon_c, \varepsilon_s)$ (where $\mathcal{A}$ is 4-monotone) secret sharing scheme $(\mathsf{SecShare}_{(\mathcal{A},n)}, \mathsf{SecRec}_{(\mathcal{A},n)})$ with statistical privacy (with error ε_s) for message space $\mathcal{L}$. We will assume that the size of each share is m_1.
- A $(3, n, 0, 0)$ secret sharing scheme $(\mathsf{LRShare}_{(3,n)}, \mathsf{LRRec}_{(3,n)})$ that is ε_3-leakage resilient against leakage functions $\mathcal{F}_{(3,n),m_1}$ for message space $\mathcal{C}$. We assume that the size of each share is m_2.
- A $(2, n, 0, 0)$ secret sharing scheme $(\mathsf{LRShare}_{(2,n)}, \mathsf{LRRec}_{(2,n)})$ for message space $\mathcal{R}$ that is ε_4-leakage resilient against leakage functions $\mathcal{F}_{(2,n),\mu}$ where $\mu = m_1 + m_2$. We assume that the size of each share is m_3.

We give the formal description of the construction in Figure 3 (taken verbatim from [BS18]).

$\mathsf{Share}(m)$: To share a secret $s \in \mathcal{M}$ do:

1. Encode the secret s as $(\mathsf{L}, \mathsf{C}, \mathsf{R}) \leftarrow \mathsf{Enc}(s)$.
2. Compute the shares
$$(\mathsf{SL}_1, \ldots, \mathsf{SL}_n) \leftarrow \mathsf{SecShare}_{(\mathcal{A},n)}(\mathsf{L})$$
$$(\mathsf{SC}_1, \ldots, \mathsf{SC}_n) \leftarrow \mathsf{LRShare}_{(3,n)}(\mathsf{C})$$
$$(\mathsf{SR}_1, \ldots, \mathsf{SR}_n) \leftarrow \mathsf{LRShare}_{(2,n)}(\mathsf{R})$$
3. For each $i \in [n]$, set share_i as $(\mathsf{SL}_i, \mathsf{SC}_i, \mathsf{SR}_i)$ and output $(\mathsf{share}_1, \ldots, \mathsf{share}_n)$ as the set of shares.

$\mathsf{Rec}(\mathsf{Share}(m)_T)$: Given a set of shares in an authorized set $T' \in \mathcal{A}$, let $T \subseteq T'$ denote a minimal authorized set. To reconstruct the secret from the shares in set T (of size at most t), do:

1. Let the shares corresponding to the set T be $(\mathsf{share}_{i_1}, \ldots, \mathsf{share}_{i_t})$.
2. For each $j \in \{i_1, \ldots, i_t\}$, parse share_j as $(\mathsf{SL}_j, \mathsf{SC}_j, \mathsf{SR}_j)$.
3. Reconstruct
$$\mathsf{L} := \mathsf{SecRec}_{(\mathcal{A},n)}(\mathsf{SL}_{i_1}, \ldots, \mathsf{SL}_{i_t})$$
$$\mathsf{C} := \mathsf{LRRec}_{(3,n)}(\mathsf{SC}_{i_1}, \mathsf{SC}_{i_2}, \mathsf{SC}_{i_3})$$
$$\mathsf{R} := \mathsf{LRRec}_{(2,n)}(\mathsf{SR}_{i_1}, \mathsf{SR}_{i_2})$$
4. Output the secret s as $\mathsf{Dec}(\mathsf{L}, \mathsf{C}, \mathsf{R})$.

Figure 3: Construction of Non-Malleable Secret Sharing Scheme for 4-monotone access structure taken verbatim from [BS18]

Imported Theorem 4.2 ([BS18]) *For any arbitrary $n \in \mathbb{N}$ and and 4-monotone access structure $\mathcal{A}$, the construction given in Figure 3 is a $(\mathcal{A}, n, \varepsilon_c, \varepsilon_s)$ secret sharing scheme. Furthermore, it is $(\varepsilon_1 + \varepsilon_3 + \varepsilon_4)$-non-malleable against $\mathcal{F}_{\mathsf{ind}}$.*

4.3 Rate Analysis

We now instantiate the primitives and provide the rate analysis.

1. We instantiate the three split state non-malleable code from the works of [KOS18, GMW17] (see Theorem A.5). Using their construction, the $|\mathsf{L}| = |\mathsf{C}| = |\mathsf{R}| = O(m)$ bits and the error $\varepsilon_1 = 2^{-\Omega(m/\log^{1+\rho}(m))}$ for any $\rho > 0$.

2. We use a secret sharing scheme for access structure $\mathcal{A}$ with rate R. We get $m_1 = O(m/R)$.

3. We instantiate $(\mathsf{LRShare}_{(3,n)}, \mathsf{LRRec}_{(3,n)})$ and $(\mathsf{LRShare}_{(2,n)}, \mathsf{LRRec}_{(2,n)})$ from Section 3.2.2. We get $m_2 = O(m)$ and $m_3 = O(m)$ by setting ε_3 and ε_4 to be $2^{-\Omega(m/\log m)}$.

Thus, the size of a share is $m_1 + m_2 + m_3 = O(m/R)$ and hence the rate is $O(R)$. The error of our construction is $2^{-\Omega(m/\log^{1+\rho}(m))}$.

We obtain the following corollary.

Corollary 4.3 *For any $n \in \mathbb{N}$, $\rho > 0$ and 4-monotone access structure $\mathcal{A}$, if there exists a statistically private (with privacy error ε) secret sharing scheme for $\mathcal{A}$ that can share m-bit secrets with rate R, there exists a non-malleable secret sharing scheme for sharing m-bit secrets for the same access structure $\mathcal{A}$ against $\mathcal{F}_{ind}$ with rate $\Omega(R)$ and simulation error $\varepsilon + 2^{-\Omega(m/\log^{1+\rho}(m))}$.*

5 Leakage Tolerant MPC for General Interaction Patterns

In this section, we will construct a leakage tolerant secure multiparty computation protocol for any interaction pattern. We will first recall some basic definitions from [HIJ+16].

5.1 Basic Definitions

This subsection is taken verbatim from [HIJ+16].

We begin by defining the syntax for specifying a communication pattern $\mathcal{I}$ and a protocol Π that complies with it. In all the definitions below, we let $\mathcal{P} = \{P_1, \ldots, P_n\}$ denote a fixed set of parties who would participate in the protocol. When we want to stress the difference between a protocol message as an entity by itself (e.g., "the 3rd message of party P_1") and the content of that message in a specific run of the protocol, we sometime refer to the former as a "message slot" and the latter as the "message content." To define an N-message interaction pattern for the parties in $\mathcal{P}$, we assign a unique identifier to each message slot. Without loss of generality, the identifiers are the indices 1 through N. An interaction pattern is then defined via a set of constraints on these message slots, specifying the sender and receiver of each message, as well as the other messages that it depends on. These constraints are specified by a message dependency graph, where the vertices are the message slots and the edges specify the dependencies.

Definition 5.1 (Interaction pattern [HIJ+16]) *An N-message interaction pattern for the set of parties $\mathcal{P}$ is specified by a message dependency directed acyclic labeled graph,*

$$\mathcal{I} = ([N], D, L : V \to \mathcal{P} \times (\mathcal{P} \cup \mathsf{Out}))$$

The vertices are the message indices $[N]$, each vertex $i \in [N]$ is labeled by a sender-receiver pair $L(i) = (S_i, R_i)$, with $R_i = \mathsf{Out}$ meaning that this message is output by party S_i rather than

sent to another party. The directed edges in D specify message dependencies, where an edge $i \to j$ means that message j in the protocol may depend on message i. The message-dependency graph must satisfy two requirements:

- *$\mathcal{I}$ is acyclic. We assume without loss of generality that the message indices are given in topological order, so $i < j$ for every $(i \to j) \in D$.*
- *If message j depends on message i, then the sender of message j is the receiver of message i. That is, for every $(i \to j) \in D$, we have $S_j = R_i$ (where $L(i) = (S_i, R_i)$ and $L(j) = (S_j, R_j)$).*

We assume without loss of generality that each party $P \in \mathcal{P}$ has at most one output, namely at most one $i \in [N]$ such that $L(i) = (P, \mathsf{Out})$. For a message $j \in [N]$, we denote its incoming neighborhood, i.e. all the messages that it depends on, by $\mathsf{DepOn}(j) := \{i : (i \to j) \in D\}$.

An n-party, N-message interaction pattern, is an N-message pattern for $\mathcal{P} = [n]$. We will interchangeably denote the i-th party as either using i or P_i.

We next define the syntax of an MPC protocol complying with a restricted fixed interaction pattern. Importantly, our model includes general correlated randomness set-up, making protocols with limited interaction much more powerful.

Definition 5.2 ($\mathcal{I}$ compliant protocol [HIJ+16]) *Let $\mathcal{I} = ([N], D, L)$ be an n-party N-message interaction pattern. An n-party protocol complying with $\mathcal{I}$ is specified by a pair of algorithms $\Pi = (\mathsf{Gen}, \mathsf{Msg})$ of the following syntax:*

- Gen *is a randomized sampling algorithm that outputs an n-tuple of correlated random strings $(r_1, \ldots, r_n)$.*
- Msg *is a deterministic algorithm specifying how each message is computed from the messages on which it depends. Concretely, the input of Msg consists of the index $i \in [N]$ of a vertex in the dependency graph, the randomness r_{S_i} and input x_{S_i} for the sender S_i corresponding to that vertex, and an assignment of message-content to all the messages that message i depends on, $M : \mathsf{DepOn}(i) \to \{0,1\}^*$. The output of Msg is an outgoing message in $\{0,1\}^*$, namely the string that the sender S_i should send to the receiver R_i.*

The execution of such a protocol Π with pattern $\mathcal{I}$ proceeds as follows. During an offline set-up phase, before the inputs are known, Gen is used to generate the correlated randomness $(r_1, ..., r_n)$ and distribute r_i to party P_i . In the online phase, on inputs $(x_1, \ldots, x_n)$, the parties repeatedly invoke Msg on vertices (message-slots) in $\mathcal{I}$ to compute the message-content they should send. The execution of Π goes over the message slots in a topological order, where each message is sent after all messages on which it depends have been received. We do not impose any restriction on the order in which messages are sent, other than complying with the depend-on relation as specified by $\mathcal{I}$. Once all messages (including outputs) are computed, the parties have local outputs $(y_1, \ldots, y_n)$, where we use $y_i = \perp$ to indicate that P_i does not have an output.

For a set $T \subset [n]$ of corrupted parties, let view_T denote the entire view of T during the protocol execution. This view includes the inputs x_T , correlated randomness r_T , and messages received by T. (Sent messages and outputs are determined by this information.) The view does not include messages exchanged between honest parties. Security of a protocol with communication pattern $\mathcal{I}$ requires that for any subset of corrupted parties $T \subset \mathcal{P}$, the view view_T reveals as little about the

inputs $x_{\overline{T}}$ of honest parties as is possible with the interaction pattern $\mathcal{I}$. We formulate this notion of "as little as possible" via the notion of fixed vs. free inputs: If parties P_i , P_j are corrupted and no path of messages from P_i to P_j passes through any honest party, then the adversary can learn the output of P_j on every possible value of x_i . However, if there is some honest party on some communication path from P_i to P_j , then having to send a message through that party may be used to "fix" the input of P_i that was used to generate that message, so the adversary can only learn the value of the function on that one input.

Definition 5.3 (Fixed vs. free inputs.) *For an interaction pattern $\mathcal{I}$, parties $P_i, P_j \in \mathcal{P}$ (input and output parties), and a set $T \subset P$ of corrupted parties, we say that P_i has fixed input with respect to $\mathcal{I}, T$ and P_j if either*

- *$P_i \notin T$ (the input party is honest), or*
- *there is a directed path in $\mathcal{I}$ starting with some message sent by P_i, ending with some message received by P_j , and containing at least one message sent by some honest party $P_h \notin T$.*

We say that P_i has free input (with respect to $\mathcal{I}, T, P_j$) if $P_i \in T$ and its input is not fixed. We let $\mathsf{Free}(\mathcal{I}, T, P_j) \subseteq T$ denote the set of parties with free inputs, and $\mathsf{Fixed}(\mathcal{I}, T, P_j) = P \setminus \mathsf{Free}(\mathcal{I}, T, P_j)$ is the complement set of parties with fixed input (all with respect to $\mathcal{I}, T$ and P_j).

Using the notion of fixed inputs, we can now capture the minimum information available to the adversary by defining a suitable restriction of the function f that the protocol needs to compute.

Definition 5.4 *For an n-party functionality f, interaction pattern $\mathcal{I}$, corrupted set $T \subset P$, input $x = (x_1, \ldots, x_n)$ and output party $P_j \in P$, the residual function $f_{\mathcal{I},T,x,P_j}$ is the function obtained from f_j by restricting the input variables indexed by $F = \mathsf{Fixed}(\mathcal{I}, T, P_j)$ to their values in x. That is, for input variables $x'_{\overline{F}} = \{x'_i\}_{i \notin F}$, we define $f_{\mathcal{I},T,x,P_j} = f_j(x_1, \ldots, x_n)$, where $x'_i = x_i$ for all $i \in F$.*

We formalize our notion of security in the semi-honest model below. To get around general impossibility results for security with polynomial-time simulation [HLP11, GGG+14, BGI+14], we will allow by default simulators to be unbounded (but will also consider bounded simulation variants). We start by considering perfectly/statistically/computationally secure protocols.

Definition 5.5 *(Security with semi-honest adversaries). Let f be a deterministic n-party functionality, $\mathcal{I}$ be an n-party, N-message interaction pattern, and $\Pi = (\mathsf{Gen}, \mathsf{Msg})$ be an n-party protocol complying with $\mathcal{I}$. We say that Π is a perfectly T-secure protocol for f in the semi-honest model for a fixed set $T \subset P$ of corrupted parties if the following requirements are met:*

- ***Correctness:*** *For every input $x = (x_1, \ldots, x_n)$, the outputs at the end of the protocol execution are always equal to $f(x)$ (namely, with probability 1 over the randomness of Gen).*
- ***Semi-honest security:*** *There is an unbounded simulator $\mathcal{S}$ that for any input x is given x_T and the truth tables of the residual functions $f_{\mathcal{I},T,x,P_j}$ for all $P_j \in T$, and its output is distributed identically/statistically close/computationally indistinguishable to $\mathsf{view}_T(x)$.*

Remark 5.6 (Efficient Simulation) *For the case where we require the simulator to be efficient, we provide the simulator with oracle access to the residual function $f_{\mathcal{I},T,x,P_j}$.*

5.2 Known Protocols for Star Interaction Pattern

We will first recall the star interaction pattern and recall some results for this case.

Star Interaction Pattern. A $n+1$-party, $n+1$-message interaction pattern is called a star interaction pattern, if for each $i \in [n]$, $L(i) = (P_i, P_{n+1})$, $(i \to n+1) \in D$ and $L(n) = (P_{n+1}, \mathsf{Out})$. In other words, for every $i \in [n]$, P_i sends a single message to P_{n+1} who computes the output from all the messages received.

Theorem 5.7 ([BGI+14]) *For any function $f : (\{0,1\}^m)^n \to \{0,1\}^*$, there exists a star compliant protocol that computes f with perfect security tolerating upto $n-1$ corruptions. The communication complexity of the protocol is exponential in nm.*

Theorem 5.8 ([BKR17]) *For any function $f : (\{0,1\}^m)^n \to \{0,1\}^*$ that is computable in NC^1, there exists an efficient, star compliant protocol that computes f with perfect security tolerating a constant number of corruptions. Furthermore, assuming the existence of one-way functions, for any function $f : (\{0,1\}^m)^n \to \{0,1\}^*$ that is computable by a circuit, there exists an efficient star compliant protocol that computes f with computational security tolerating a constant number of corruptions.*

Theorem 5.9 ([GGG+14]) *Assuming the existence of indistinguishability obfuscation and one-way functions, for any function $f : (\{0,1\}^m)^n \to \{0,1\}^*$ that is computable by circuits, there exists an efficient, star compliant protocol that computes f with computational security tolerating upto $n-1$ corruptions.*

5.3 Definition: Leakage Tolerant MPC for an Interaction Pattern

We now give the definition of leakage-tolerant MPC for an interaction pattern $\mathcal{I}$.

We consider $(n+1)$-party $\mathcal{P} = \{P_1, \ldots, P_n, P_{n+1}\}$ protocol Π that is compliant with an interaction pattern $\mathcal{I}$ with a single output party (namely, P_{n+1} is the output party and does not have any inputs) that computes a function of $f : (\{0,1\}^m)^n \to \{0,1\}^*$, where the party P_i gets input $x_i \in \{0,1\}^m$ for each $i \in [n]$. Suppose at the end of a protocol Π, the party P_i's view $view_i$ is from a domain $\mathcal{V}_i$. Let us denote $\Pi(x)$ as the joint distribution of the views of every party during the execution of the protocol. We are interested in adversaries that statically corrupt t $(< n)$ of the parties, obtaining their entire states, and also obtain some leakage on the states of the other uncorrupted parties. More formally, we represent the view of such adversaries as families of functions of the form $\mathcal{G}_{t,\mu} = \{g_{T,\vec{\tau}} : T \subseteq [n], |T| \leq t, \tau_i : \mathcal{V}_i \to \{0,1\}^\mu\}$; where $g_{T,\vec{\tau}}(\Pi(x))$ outputs v_i for every $i \in T$, and $\tau_i(view_i)$ for $i \notin T$, when the protocol Π is run with input x – we refer to such a function as a (T, μ)-leakage function.

Definition 5.10 (Leakage Tolerance against Semi-Honest Adversaries) *Let f be a deterministic n-party functionality, $\mathcal{I}$ be an n-party, N-message interaction pattern, and $\Pi = (\mathsf{Gen}, \mathsf{Msg})$ be an n-party protocol complying with $\mathcal{I}$. We say that Π is a (T, μ)-* leakage tolerant protocol for f *in the semi-honest model for a set $T \subseteq \mathcal{P}$ if the it satisfies the following properties:*

- ***Correctness:*** *The protocol Π computes $f(x)$ correctly for any input $x = (x_1, \ldots, x_n)$.*

- ***Leakage Tolerance:*** *For any (T, μ)- leakage function $g_{T,\vec{\tau}}$, there is an unbounded simulator $\mathcal{S}$ satisfying the following.*
  - *For any input $x = (x_1, \ldots, x_n)$, the simulator $\mathcal{S}$ is given the inputs of the corrupted parties x_T and the truth tables of the residual functions $f_{\mathcal{I},T,x,P_j}$ for all $P_j \in T$ as input. It is allowed a single query to an oracle $\mathcal{O}[x_{\overline{T}}]$, which takes as input a tuple of functions $(\sigma_i)_{i\in\overline{T}}$, where each function is of the form $\sigma_i : \{0,1\}^m \to \{0,1\}^\mu$, and outputs $(\sigma_i(x_i))_{i\in\overline{T}}$.*
  - *We require that:*
$$g_{T,\vec{\tau}}(\Pi(x)) \approx \mathcal{S}^{\mathcal{O}[x_{\overline{T}}]}(f_{\mathcal{I},T,x,P_j}, x_T)$$
    where $\approx$ might indicate identical/statistically close/computationally indistinguishable.

We say that Π is a (t, μ)-(adaptive) leakage tolerant protocol for *f if it is (T, μ)-(adaptive) leakage tolerant for all $T \subseteq \mathcal{P}$ and $|T| \leq t$.*

5.4 Construction

In this subsection, we give a construction of a leakage-tolerant semi-honest MPC for any interaction pattern $\mathcal{I}$. The construction we give is the same as the one given in [HIJ+16] with the only change being that we use our strong local leakage-resilient scheme instead of any secret sharing scheme.

Before we describe the construction, we introduce the following notation. For a function $f : (\{0,1\}^m)^n \to \{0,1\}$, we denote by $f^{bit} : \{0,1\}^{mn} \to \{0,1\}$ the function that takes mn bits as inputs, groups them together in order into n strings of length m each, and applies f on them.

Building Blocks. The construction uses the following building blocks:

- A star compliant, semi-honest protocol $\Pi^* = (\mathsf{Gen}^*, \mathsf{Msg}^*, \mathsf{Eval}^*)$ that securely (either perfect/statistical/computational) computes the function f^{bit}. Here, Msg^* denotes the next message function of the parties $P_1, \ldots, P_{mn}$ and Eval^* is the function computed by the evaluator (or in other words, party P_{mn+1}).
- A $(n+1, n+1, 0, 0)$ threshold secret sharing scheme $(\mathsf{LRShare}, \mathsf{LRRec})$ that is ε-strong leakage resilient for some negligible ε against the function family $\mathcal{H}_{n-1,n,\mu}$.

Construction. Let $f : (\{0,1\}^m)^n \to \{0,1\}^*$ be a n-party functionality that depends on all its inputs and $\mathcal{I}$ be an interaction pattern with a single sink. Let $\mathcal{P} = \{P_1, \ldots, P_{n+1}\}$ be the set of parties with P_{n+1} being the evaluator who does not have any inputs. We give the construction of an $\mathcal{I}$ compliant protocol in Figure 4.

Theorem 5.11 *If Π^* computes f^{bit} with statistical/computational security and $(\mathsf{LRShare}, \mathsf{LRRec})$ is an ε-leakage resilient secret sharing scheme for some negligible ε then for any $\mu \leq m$, the construction in Figure 4 is a semi-honest, $\mathcal{I}$-compliant protocol for f that is (n, μ)-leakage tolerant with statistical/computational security. Furthermore, if each party gets R bits of correlated randomness and sends M bits in the protocol Π^*, then in $O(m(R + n^2M))$ bits of correlated randomness and sends $O(n^2Mm)$ bits.*

Proof The correctness of the protocol follows directly from the correctness of $(\mathsf{LRShare}, \mathsf{LRRec})$ and that of Π^*. We now argue leakage tolerance.

Gen : To generate the correlated randomness, do:

1. Run Gen^* to obtain the correlated randomness $(r_1, \ldots, r_{mn+1})$.
2. For each $i \in [mn]$ and $\sigma \in \{0,1\}$, compute $m_i^\sigma := \mathsf{Msg}^*(\sigma, r_i)$.
3. For each $i \in [mn]$ and $\sigma \in \{0,1\}$, compute $(m_{i,1}^\sigma, \ldots, m_{i,n+1}^\sigma) \leftarrow \mathsf{LRShare}(m_i^\sigma)$.
4. Choose random permutation strings $b_1, \ldots, b_n \leftarrow \{0,1\}^m$, one for each party P_i, $i \leq n$.
5. Let $c = b_1 \circ b_2 \circ \ldots \circ b_n$. For each $i \in [mn]$, let c_i denote the i-th bit of c.
6. For each $j \in [n]$, the correlated randomness for party j is $(\{m_{i,j}^{c_i}, m_{i,j}^{1-c_i}\}_{i\in[mn]}, b_j)$. The correlated randomness of the evaluator P_{n+1} is $(r_{mn+1}, \{m_{i,n+1}^{c_i}, m_{i,n+1}^{1-c_i}\}_{i\in[mn]})$

Msg : On input $x_j \in \{0,1\}^m$ and the correlated randomness, party P_j does the following:

1. Parses the correlated randomness as $(\{M_{i,j}^0, M_{i,j}^1\}_{i\in[mn],\sigma\in\{0,1\}}, b_j)$.
2. Computes $s_j = x_j \oplus b_j$ and sends s_j on every path to the evaluator in $\mathcal{I}$.
3. Then, for every P_k such that some path from P_k to the evaluator goes through P_j, party P_j waits until it receives the string s_k and then sends $\{M_{(k-1)m+\ell,j}^{s_{k,\ell}}\}_{\ell\in[m]} = \{m_{(k-1)m+\ell,j}^{x_{k,\ell}}\}_{\ell\in[m]}$ on every path to the evaluator.
4. For every P_k such that no path from P_k to the evaluator goes through P_j, party P_j sends both shares $\{M_{(k-1)m+\ell,j}^0, M_{(k-1)m+\ell,j}^1\}_{\ell\in[m]}$ on every path to the evaluator.
5. In addition, P_j forwards every message that it receives from other parties on some $\mathcal{I}$-path to the evaluator.

Eval: The evaluator uses its correlated randomness to reconstruct $M_{(k-1)m+\ell}^{s_{k,\ell}}$ for every $k \in [n]$ and $\ell \in [m]$. It then uses the function Eval^* on these reconstructed values to learn the output.

Figure 4: A $\mathcal{I}$ compliant protocol computing f. The construction is same as the one in [HIJ+16] except that we use our leakage resilient secret sharing.

Leakage Tolerance. To prove leakage tolerance of the protocol, we need to do the following. Given a (T, μ)-leakage function $g_{T,\vec{\tau}}$, we need to construct a simulator $\mathcal{S}$ such that,

$$g_{T,\vec{\tau}}(\Pi(x)) \approx \mathcal{S}^{\mathcal{O}[x_{\overline{T}}]}(f_{\mathcal{I},T,x,P_{n+1}}, x_T)$$

Depending on whether Π^* has statistical/computational security, we get $\approx$ to either be statistically or computationally close.

Description of Simulator. Let $T \subset [n]$ be the set of corrupted parties and let H be the set of honest parties i.e., $H = [n] \setminus T$. Partition T as T_{fixed} and T_{free} where $T_{fixed} = T \cap \mathsf{Fixed}(\mathcal{I}, T, P_{n+1})$ and $T_{free} = T \setminus T_{fixed}$. Let $H^* = T_{fixed} \cup H$ and $T^* = [n] \setminus H^*$. The simulator $\mathcal{S}$ does the following:

- It runs the simulator $\mathcal{S}^*$ for the star compliant protocol Π^* for computing f^{bit}. To be more precise, $\mathcal{S}$ runs $\mathcal{S}^*$ by specifying $\widehat{T} = \{(i-1)m + \ell : i \in T^*, \ell \in [m]\}$ as the set of corrupted parties and gives each bit of the string x_{T^*} as the corresponding corrupted parties' input. It sets the residual function to be given as input to $\mathcal{S}^*$ as $f_{\mathcal{I},T,x,P_{n+1}}$. Notice that by definition, the residual function $f_{\mathcal{I},T,x,P_{n+1}}$ fixes the inputs x_{H^*} and leaves x_{T^*} as free.

- The output of $\mathcal{S}^*$ is the set of correlated randomness for the corrupted parties, namely, $\{r_{(i-1)m+\ell}\}_{\ell\in[m]}$ for each $i \in T^*$ and the set of honest party messages $\{m_{(i-1)m+\ell}\}_{\ell\in[m]}$ for each $i \in H^*$.

- $\mathcal{S}$ uses the correlated randomness $\{r_{(i-1)m+\ell}\}_{\ell\in[m]}$ to generate the messages for the corrupted parties. Specifically, for each $i \in T^*$ and $\ell \in [m]$, $\mathcal{S}$ computes $m^{\sigma}_{(i-1)m+\ell} = \mathsf{Msg}^*(\sigma, r_{(i-1)m+\ell})$ for $\sigma \in \{0,1\}$.

- For every party $i \in H^*$, $\mathcal{S}$ sets both $(m^0_{(i-1)m+\ell}, m^1_{(i-1)m+\ell})$ to be equal to $\{m_{(i-1)m+\ell}\}_{\ell\in[m]}$ obtained from $\mathcal{S}^*$.

- For each $i \in [mn]$ and $\sigma \in \{0,1\}$, $\mathcal{S}$ computes $(m^{\sigma}_{i,1}, \ldots, m^{\sigma}_{i,n+1}) \leftarrow \mathsf{LRShare}(m^{\sigma}_i)$.

- For each $i \in T \setminus \{n+1\}$, $\mathcal{S}$ chooses random permutation strings $b'_i \xleftarrow{\$} \{0,1\}^m$. For every honest party $i \in H$, $\mathcal{S}$ sets $b'_i = 0^m$.

- Let $c = b'_1 \circ b'_2 \circ \ldots \circ b'_n$. For each $i \in [mn]$, let c_i denote the i-th bit of c.

- For each $j \in T$, $\mathcal{S}$ sets the correlated randomness for party P_j for some $j \leq n$ as $R_j = (\{m^{c_i}_{i,j}, m^{1-c_i}_{i,j}\}_{i\in[mn],\sigma\in\{0,1\}}, b'_j)$. If $P_{n+1} \in T$, it sets the evaluator's correlated randomness to be $(r_{mn+1}, \{m^{c_i}_{i,n+1}, m^{1-c_i}_{i,n+1}\}_{i\in[mn]})$.

- $\mathcal{S}$ generates the protocol messages as follows. Let the interaction pattern $\mathcal{I}$ be specified by $([N], D, L : V \to \mathcal{P} \times (\mathcal{P} \cup \mathsf{Out}))$. For each k in 1 to N, $\mathcal{S}$ does the following:
  - Let $L(k) = (P_{r_1}, P_{r_2})$.
  - If $P_{r_1} \in T$ (or in other words, it is corrupted) then, $\mathcal{S}$ uses the input of the party P_{r_1} denoted by x_{r_1} and the correlated randomness R_{r_1} to generate the protocol message honestly.
  - If $P_{r_1} \in H$, then $\mathcal{S}$ chooses a random string $s_{r_1} \leftarrow \{0,1\}^m$ (or reuses s_{r_1} if it has been picked before). $\mathcal{S}$ sends the party P_{r_1}'s masked input as s_{r_1} and sends the rest of the messages (i.e., forwarding the appropriate set of shares) exactly as in the protocol.

- For every $k \in H$, let msg_k be the set of messages that P_k has received during the execution of the protocol. $\mathcal{S}$ defines the leakage function σ_k to be queried to the oracle $\mathcal{O}$ as follows. σ_k has the messages msg_k, the set of shares $\mathsf{shares}_k = \{m^{c_i}_{i,k}, m^{1-c_i}_{i,k}\}_{i\in[mn],\sigma\in\{0,1\}}$ and the string s_k hardwired and on input x_k, computes $b_k := x_k \oplus s_k$ and outputs $\tau_k(\mathsf{msg}_k, \mathsf{shares}_k, x_k, b_k)$.

- $\mathcal{S}$ outputs the protocol messages sent to and by the corrupted parties and $\{(R_i, x_i)\}_{i\in T}$ in the clear and for all $i \in H$, it outputs $\sigma_i(x_i)$.

We now argue that the output of the simulator is statistically/computationally close to the output of $g_{T,\overrightarrow{\tau}}(\Pi(x))$. We show this via a hybrid argument.

$\underline{\mathsf{Hyb}_1}$: In this hybrid, we make the following change with respect to $g_{T,\overrightarrow{\tau}}(\Pi(x))$. For each $i \in H^*$ and $\ell \in [m]$, we generate the shares $\{m^{\sigma}_{(i-1)m+\ell,j}\}_{j\in[n+1],\sigma\in\{0,1\}}$ which is part of the correlated randomness as a leakage resilient secret sharing of the secret $m^{x_{i,\ell}}_{(i-1)\ell+m}$ for both $\sigma \in \{0,1\}$. That is, both set of shares correspond to the secret sharing of the same secret, namely, $m^{x_{i,\ell}}_{(i-1)\ell+m}$. Notice that by definition of H^*, it follows that for every $i \in H^*$, there exists at least one party P_j (for some $j \leq n$) that does not reveal the share $m^{1-x_{i,\ell}}_{(i-1)\ell+m,j}$ for each $\ell \in [m]$ and hence, intuitively, it should follow from the local leakage resilience property of our secret sharing scheme that $\mathsf{Hyb}_1 \approx_s g_{T,\overrightarrow{\tau}}(\Pi(x))$. However, the proof of this claim is involved and we in fact, require the underlying leakage resilient secret sharing to be strong. We now give the details.

Lemma 5.12 *If* $(\mathsf{LRShare}, \mathsf{LRRec})$ *is an* ε*-strong leakage resilient secret sharing scheme, then* $\mathsf{Hyb}_1 \approx_{nm\varepsilon} g_{T,\overrightarrow{\tau}}(\Pi(x))$.

Proof We prove this lemma by defining a sequence of sub-hybrids.

Let Γ be the set $\{(i,\ell) : i \in H^*, \ell \in [m]\}$. Let $\prec$ be a total ordering on the set Γ. For every element $\gamma \in \Gamma$, we define a hybrid distribution Hyb_γ where the correlated randomness is generated as follows:

- Run Gen^* to obtain the correlated randomness $(r_1, \ldots, r_{mn})$.
- For each $i \in [mn]$ and $\sigma \in \{0,1\}$, compute $m^{\sigma}_i := \mathsf{Msg}^*(\sigma, r_i)$.
- For each $i \in H^*$ and $\ell \in [m]$, if $(i,\ell) \prec \gamma$, reset $m^{\sigma}_{(i-1)m+\ell} := \mathsf{Msg}^*(x_{i,\ell}, r_{(i-1)\ell+m})$ for both $\sigma = \{0,1\}$.
- For each $i \in [mn]$ and $\sigma \in \{0,1\}$, compute $(m^{\sigma}_{i,1}, \ldots, m^{\sigma}_{i,n+1}) \leftarrow \mathsf{LRShare}(m^{\sigma}_i)$.
- Choose random permutation strings $b_1, \ldots, b_n \leftarrow \{0,1\}^m$, one for each party P_i, $i \leq n$.
- Let $c = b_1 \circ b_2 \circ \ldots \circ b_n$. For each $i \in [mn]$, let c_i denote the i-th bit of c.
- For each $j \in [n]$, the correlated randomness for party j is $(\{m^{c_i}_{i,j}, m^{1-c_i}_{i,j}\}_{i\in[mn]}, b_j)$. The correlated randomness of the evaluator P_{n+1} is $\{m^{c_i}_{i,n+1}, m^{1-c_i}_{i,n+1}\}_{i\in[mn]}$

The protocol is then run exactly as in Figure 4 with the above generated correlated randomness. We denote the joint distribution of the views of all the parties where the correlated randomness is generated as above using $\Pi_\gamma(x)$. The output of Hyb_γ is $g_{T,\overrightarrow{\tau}}(\Pi_\gamma(x))$.

Let γ_{first} be the first element as per the ordering $\prec$. We note that $\mathsf{Hyb}_{\gamma_{first}}$ is distributed identically to $g_{T,\overrightarrow{\tau}}(\Pi(x))$. We now prove the following claim:

Claim 5.13 *For any* $\gamma, \gamma' \in \Gamma$ *where* γ' *is the successor of* γ *as per the ordering* $\prec$*, we have that* $\mathsf{Hyb}_\gamma \approx_\varepsilon \mathsf{Hyb}_{\gamma'}$.

Proof Assume for the sake of contradiction that the statistical distance between Hyb_γ and $\mathsf{Hyb}_{\gamma'}$ is greater than ε. We will use this to contradict the security of the strong leakage resilience of $(\mathsf{LRShare}, \mathsf{LRRec})$. Let $\gamma = (i^*, \ell^*)$. We first define the concept of friend parties.

Definition 5.14 (Friend Party) *We define a party P_j to be a friend of P_i for an $i \in H^*$ as follows:*

- *If $i \in H$, then $P_j = P_i$.*
- *Else, if $i \in H^* \setminus H$, then $P_j \in H$ is a friend of P_i such that there is a directed path in $\mathcal{I}$ starting with some message sent by P_i, ending with some message received by the evaluator P_{n+1} , and containing at least one message sent by the party P_j.*

Intuitively, the friend party of P_{i^*} will not reveal its share $m^{1-x_{i^*,\ell^*}}_{(i^*-1)\ell^*+m,j^*}$ and we can make use of this to argue indistinguishability between Hyb_γ and $\mathsf{Hyb}_{\gamma'}$. We formalize this intuition by giving a reduction to the strong leakage resilience property of the secret sharing scheme. The reduction proceeds as follows:

1. Let P_{j^*} be the friend of P_{i^*}.
2. Let Send_{j^*} be the set of parties in $\{P_1, \ldots, P_n\}$ such that P_{j^*} receives both shares from this party for the ℓ^*-th bit of the input of party P_{i^*} during the execution of the protocol. We note that $|\mathsf{Send}_{j^*}| \leq n-1$, since Send_{j^*} does not include i^*.
3. Run Gen^* to obtain the correlated randomness $(r_1, \ldots, r_{mn})$.
4. For each $i \in [mn]$ and $\sigma \in \{0,1\}$, compute $m_i^\sigma := \mathsf{Msg}^*(\sigma, r_i)$.
5. For each $i \in H^*$ and $\ell \in [m]$, if $(i,\ell) \prec \gamma$, reset $m^\sigma_{(i-1)m+\ell} := \mathsf{Msg}^*(x_{i,\ell}, r_{(i-1)\ell+m})$ for both $\sigma = \{0,1\}$.
6. For each $i \in [mn]$ and $\sigma \in \{0,1\}$, compute $(m^\sigma_{i,1}, \ldots, m^\sigma_{i,n+1}) \leftarrow \mathsf{LRShare}(m_i^\sigma)$.
7. Reset for each $j \in [n]$, $m^{1-x_{i^*,\ell^*}}_{(i^*-1)m+\ell^*,j} = \perp$.
8. Interact with the strong leakage resilient challenger. Provide $(m_0, m_1) = (m^{x_{i^*,\ell^*}}_{(i^*-1)m+\ell^*}, m^{1-x_{i^*,\ell^*}}_{(i^*-1)m+\ell^*})$ as the two challenge messages and ask the challenger to reveal the shares of the parties in the set Send_{j^*}. The challenger replies with the set of shares share_k for each $k \in \mathsf{Send}_{j^*}$ where this set of shares correspond to a sharing of m_0 or m_1.
9. The reduction generates the permutation strings exactly as per the description of the protocol.
10. Using $\{\mathsf{share}_k\}_{k \in \mathsf{Send}_{j^*}}$, the set of shares it generated in Step 6, the parties inputs and the permutation strings, the reduction generates all the protocol messages that are sent. We note that this is possible since each party $j \in [n] \setminus \mathsf{Send}_{j^*}$, do not make use of missing shares in generating the protocol messages.
11. Let msg_{j^*} be the set of messages that P_{j^*} has received during the execution of the protocol. The reduction defines a leakage function σ_k (that has μ bits of output) to be sent to the challenger as follows. σ_{j^*} has the following values hardwired:
    - The messages msg_{j^*} that P_{j^*} received during the execution of the protocol.
    - All the shares generated as a part of the correlated randomness given to P_{j^*} except the share share_{j^*}.

- The input x_{j^*} and the permutation string b_{j^*}.

On input the share share_{j^*}, it computes τ_{j^*} applied to the secret state of P_{j^*}.

12. It provides this leakage function σ_{j^*} to the challenger and asks for the shares of the remaining parties in the clear.

13. Using the information received from the challenger, the reduction outputs the protocol messages sent to and by the corrupted parties and their private state and for all $i \in H$, it outputs the leakage function τ_i applied on their secret state.

We note that if the shares correspond to a sharing of m_0 then the output of the reduction is identical to $\mathsf{Hyb}_{\gamma'}$. Else, it is identical to Hyb_{γ}. This contradicts the strong leakage resilience of the secret sharing. ■

By repeated application of Claim 5.13, we infer that $\mathsf{Hyb}_{\gamma_{first}} \approx_{(nm-1)\varepsilon} \mathsf{Hyb}_{\gamma_{last}}$ where γ_{last} is the last element in Γ as per the ordering $\prec$. Note that $\mathsf{Hyb}_{\gamma_{last}} \approx_{\varepsilon} \mathsf{Hyb}_1$ via another application of Claim 5.13. Thus, we infer that $\mathsf{Hyb}_1 \approx_{nm\varepsilon} g_{T,\vec{\tau}}(\Pi(x))$. This completes the proof of the lemma. ■

$\underline{\mathsf{Hyb}_2}$: In this hybrid, for every $i \in H^*$ and $\ell \in [m]$, we set $m^{x_{i,\ell}}_{(i-1)\ell+m}$ (used in computing the secret shares) as the output of $\mathcal{S}^*$ on input x_{T^*} and the residual function $f_{\mathcal{I},T,x,P_{n+1}}$. That is, instead of computing $m^{x_{i,\ell}}_{(i-1)\ell+m}$ using the input bit $x_{i,\ell}$ and the correlated randomness $r_{(i-1)\ell+m}$, we will use the simulator $\mathcal{S}^*$ for the star compliant protocol Π^* to generate this message. Note that it follows from the security of Π^* that Hyb_1 is statistically/computationally close to Hyb_2 depending on whether Π^* was statistically/computationally secure.

$\underline{\mathsf{Hyb}_3}$: This hybrid is same as the output of the simulator $\mathcal{S}$. Note that the only difference between Hyb_2 and Hyb_3 is syntactic. In Hyb_2, for each $i \in H$, we use the actual inputs x_i for each $i \in H^*$ to generate the output of $\tau_i(\mathsf{msg}_i, \mathsf{shares}_i, x_i, b_i)$ whereas in Hyb_3, we use $\sigma_i(x_i)$. Notice that the output of $\tau_i(\mathsf{msg}_i, \mathsf{shares}_i, x_i, b_i)$ is the same as the output of $\sigma_i(x_i)$ in the description of the simulator. Thus, these two hybrids are identically distributed.

This completes the proof of Theorem 5.11. ■

Instantiation. Using the known protocols for the star interaction pattern (see Section 5.2), we obtain the following corollary.

Corollary 5.15 ([BGI+14, BKR17, GGG+14]) *Let $\mathcal{I}$ be a n-party interaction pattern with a single sink and let be $f : (\{0,1\}^m)^n \to \{0,1\}^*$ be a function which depends on all its inputs. Then,*

- *There is a statistical $\mathcal{I}$-compliant leakage tolerant protocol that securely computes f upto $n-1$ passive corruptions. The communication complexity is exponential in n, m.*

- *If f is computable by a circuit in NC^1, then there exists an efficient $\mathcal{I}$-compliant leakage tolerant protocol that computes f with statistical security upto a constant number of corruptions. Assuming one-way functions, every f that is computable by circuits has a computationally secure, efficient, $\mathcal{I}$-compliant leakage tolerant protocol upto a constant number of corruptions.*

- *Assuming indistinguishability obfuscation and one-way functions, every function computable by circuits has a computationally secure, efficient, $\mathcal{I}$-compliant leakage tolerant protocol upto $n-1$ passive corruptions.*

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A Background: Non-Malleable Codes

We start with the definition of a coding scheme.

Definition A.1 (Coding scheme) *Let* $\mathsf{Enc} : \{0,1\}^m \to \{0,1\}^n$ *be a randomized algorithm and* $\mathsf{Dec} : \{0,1\}^n \to \{0,1\}^m \cup \{\perp\}$ *be a deterministic function. We say that* $(\mathsf{Enc}, \mathsf{Dec})$ *is a* coding scheme with code length n and message length m *if for all* $s \in \{0,1\}^m$, $\Pr[\mathsf{Dec}(\mathsf{Enc}(s)) = s] = 1$, *where the probability is taken over the randomness of* Enc. *The* rate *of the coding scheme is* $\frac{m}{n}$.

Dziembowski, Pietrzak and Wichs [DPW10] introduced the notion of non-malleable codes which generalizes the usual notion of error correction. In particular, it guarantees that when a codeword is subject to tampering attack, the reconstructed message is either the original one or something that is independent of the original message.

Definition A.2 (Non-Malleable Codes [DPW10]) *Let* $\mathsf{Enc} : \{0,1\}^m \to \{0,1\}^n$ *and* $\mathsf{Dec} : \{0,1\}^n \to \{0,1\}^m \cup \{\perp\}$ *be (possibly randomized) functions, such that* $\mathsf{Dec}\big(\mathsf{Enc}(s)\big) = s$ *with probability 1 for all* $s \in \{0,1\}^m$. *Let* $\mathcal{F}$ *be a family of tampering functions and fix* $\varepsilon > 0$. *We say that* $(\mathsf{Enc}, \mathsf{Dec})$ *is* $\varepsilon-$non-malleable *w.r.t.* $\mathcal{F}$ *if for every* $f \in \mathcal{F}$, *there exists a random variable* D_f *on* $\{0,1\}^m \cup \{\mathsf{same}^\star\}$, *such that for all* $s \in \{0,1\}^m$,

$$|\mathsf{Dec}(f(X_s)) - \mathrm{copy}(D_f, s)| \leq \varepsilon$$

where $X_s \leftarrow \mathsf{Enc}(s)$ *and* copy *is defined by* $\mathrm{copy}(x,y) = \begin{cases} x & \text{if } x \neq \mathsf{same}^\star \\ y & \text{if } x = \mathsf{same}^\star \end{cases}$. *We call* n *the* length *of the code and* m/n *the* rate.

Split-state Tampering Functions. We focus on the *split-state* tampering model where the encoding scheme splits s into c states: $\mathsf{Enc}(s) = (\mathcal{S}_1, \ldots, \mathcal{S}_c) \in \mathcal{S}_1 \times \mathcal{S}_2 \ldots \times \mathcal{S}_c$ and the tampering family is $\mathcal{F}_{split} = \big\{(f_1, \ldots, f_c) \big| f_i : \mathcal{S}_i \to \mathcal{S}_i\big\}$. We will call such a code as c-split-state non-malleable code.

Augmented Non-Malleable Codes. We recall the definition of augmented, 2-split-state non-malleable codes [AAG+16].

Definition A.3 (Augmented Non-Malleable Codes [AAG+16]) *A coding scheme* $(\mathsf{Enc}, \mathsf{Dec})$ *with code length* $2n$ *and message length* m *is an augmented 2-split-state non-malleable code with error* ε *if for every function* $f, g : \{0,1\}^n \to \{0,1\}^n$*, there exists a random variable* $D_{(f,g)}$ *on* $\{0,1\}^n \times (\{0,1\}^m \cup \{\mathsf{same}^\star\})$ *such that for all messages* $s \in \{0,1\}^m$*, it holds that*

$$|(\mathsf{L}, \mathsf{Dec}(f(\mathsf{L}), g(\mathsf{R}))) - \mathcal{S}(D_{(f,g)}, s)| \leq \varepsilon$$

where $(\mathsf{L}, \mathsf{R}) = \mathsf{Enc}(s)$*,* $(\mathsf{L}, \widetilde{m}) \leftarrow D_{f,g}$ *and* $\mathcal{S}((\mathsf{L}, \widetilde{m}), s)$ *outputs* (L, s) *if* $\widetilde{m} = \mathsf{same}^\star$ *and otherwise outputs* $(\mathsf{L}, \widetilde{m})$*.*

Explicit Constructions. We now recall the constructions of split-state non-malleable codes.

Theorem A.4 ([Li17]) *For any* $n \in \mathbb{N}$*, there exists an explicit construction of* 2*-split-state non-malleable code with efficient encoder/decoder, code length* $2n$*, rate* $O(\frac{1}{\log n})$ *and error* $2^{-\Omega(\frac{n}{\log n})}$*.*

Theorem A.5 ([KOS18, GMW17]) *For every* $n \in \mathbb{N}$ *and* $\rho > 0$*, there exists an explicit construction of* 3*-split-state non-malleable code with efficient encoder/decoder, code length* $(3 + o(1))n$*, rate* $\frac{1}{3+o(1)}$ *and error* $2^{-\Omega(n/\log^{1+\rho}(n))}$*.*

Additional Property. It was shown in [BS18] that the construction given in [KOS18, GMW17] satisfies the property that given two particular states of the codeword, the message remains statistically hidden.