SafetyPin: Encrypted Backups with Human-Memorable Secrets

Emma Dauterman  
UC Berkeley

Henry Corrigan-Gibbs  
EPFL and MIT CSAIL

David Mazières  
Stanford

Abstract. We present the design and implementation of SafetyPin, a system for encrypted mobile-device backups. Like existing cloud-based mobile-backup systems, including those of Apple and Google, SafetyPin requires users to remember only a short PIN and defends against brute-force PIN-guessing attacks using hardware security protections. Unlike today’s systems, SafetyPin splits trust over a cluster of hardware security modules (HSMs) in order to provide security guarantees that scale with the number of HSMs. In this way, SafetyPin protects backed-up user data even against an attacker that can adaptively compromise many of the system’s constituent HSMs. SafetyPin provides this protection without sacrificing scalability or fault tolerance. Decentralizing trust while respecting the resource limits of today’s HSMs requires a synthesis of systems-design principles and cryptographic tools. We evaluate SafetyPin on a cluster of 100 low-cost HSMs and show that a SafetyPin-protected recovery takes 1.01 seconds. To process 1B recoveries a year, we estimate that a SafetyPin deployment would need 3,100 low-cost HSMs.

1 Introduction

Modern mobile phones and tablets back up sensitive data to the cloud. To protect users’ privacy, this data must be encrypted under keys that are not available to the cloud provider. Unfortunately, with 3.8 billion smartphone users, it is impractical to expect them all to store, say, a 128-bit AES backup key. Not everyone has a computer, or trustworthy friends who can keep shares of a backup key, or even a safe place to store a backup key on paper. As a result, mobile OSes have fallen back to protecting backups with the least common denominator: device screen-lock PINs. Using PINs is good for security because a user’s screen-lock PIN never leaves her device (so the cloud provider never learns it). Using PINs is good for usability because users generally remember them.

Unfortunately, PINs have such low entropy (e.g., six decimal digits) that no feasible amount of key stretching can protect against brute-force PIN-guessing attacks. Instead, modern backup systems—such as those from Apple [49], Google [87], and Signal [57]—rely on hardware-security modules (HSMs) in their data centers to thwart brute-force attacks. Specifically, devices encrypt their backup keys under the public keys of HSMs, but each device includes a hash of its screen-lock PIN as part of the plaintext. HSMs return decrypted plaintext only to clients that can supply this PIN hash. Furthermore, HSMs limit the number of decryption attempts for any given user account. For fault tolerance, a device typically encrypts its backup key to the public keys of five HSMs, allowing any one of the five to recover the backup key.

This status quo still falls short of acceptable privacy for two reasons. First, HSMs are not perfect, yet each HSM in these systems is a single point of security failure for millions of users’ backup keys. Second, these systems make it difficult for clients to detect security breaches. For instance, if a malicious insider working in a data center physically steals an HSM, then to anyone outside the company it looks like an unremarkable single hardware failure. Alternatively, if an insider successfully guesses someone’s PIN, the victim may have no idea her backup was ever compromised.

This paper presents SafetyPin, a PIN-based encrypted-backup system with stronger security properties. The key idea behind SafetyPin is that recovering any user’s backed-up data either requires (a) guessing the user’s PIN or (b) compromising a very large number of HSMs—e.g., 6% of all HSMs operated by a provider. (The 6% figure here is a tunable system parameter.) Such large-scale attacks would typically need to span multiple data centers, be harder for insiders to pull off undetected against physical devices, cost more, and also likely cause service disruptions visible to end users.

One way to achieve SafetyPin’s security goal would be to threshold-encrypt the client’s hashed PIN and backup key in such a way that decrypting the client’s backup key would require the participation of 6% of all HSMs in the system. Unfortunately, this approach lacks scalability. If each client recovering a backup must interact with 6% of the system’s HSMs, adding more HSMs improves security without improving throughput. As the number of HSMs in the system increases, we would like the system’s overall throughput to increase in tandem with its security (i.e., the attacker’s cost).
To achieve scalability, SafetyPin takes a different approach: devices threshold-encrypt their backup keys to a small cluster of \( n \) HSMs such that decryption requires the participation of most HSMs in the cluster. The cluster size \( n \) is independent of the total number of HSMs in the system, and depends on both the fraction of compromised HSMs the system can tolerate and the fraction of HSMs that can fail-stop. (For example, to tolerate the compromise of 6% of HSMs where half of a cluster is allowed to fail-stop, we can set the cluster size \( n = 40 \).) This design achieves our scalability goal, since each device need only communicate with a small fixed number of HSMs during recovery. This design also achieves our security goal because the cluster of \( n \) HSMs that can decrypt a client’s backup depends on the client’s secret PIN, via a primitive we introduce called location-hiding encryption. Hence, even if an attacker compromises 6% of the HSMs in the system as a whole, the chances that the attacker compromises a “useful” set of HSMs—i.e., at least half of the HSMs in the device’s chosen cluster—is very small. More precisely, we show that if the total number of HSMs in the system is large enough (a few hundred or more), the probability that an attacker can decrypt a backup via HSM compromise is not much higher than the probability of simply guessing the client’s PIN.

In modern backup systems, each HSM only needs to monitor the number of PIN attempts for a small subset of users, but because of our location-hiding encryption primitive, every HSM needs to be able to verify the number of PIN attempts for every user. To maintain this information scalably, the HSMs use a new type of distributed log. Third parties can monitor this log to alert users whenever a backup-recovery attempt is underway. Since a compromised service provider may see which HSMs a mobile device interacts with during recovery (and could compromise those HSMs to recover the users’ backed-up data), HSMs revoke their ability to decrypt backups after completing the recovery process. Implementing this revocation requires adapting “puncturable encryption” [40] to storage-limited HSMs. While our prototype is focused on PIN-protected backups, these primitives have potentially broader applicability to problems such as private storage in peer-to-peer systems and cryptocurrency “brain wallets.”

We implemented SafetyPin on low-cost SoloKey HSMs [77]. We evaluate the system using a cluster of 100 SoloKeys (Figure 1) and an Android phone (representing the client device). Generating a recovery ciphertext on the client, excluding the time to encrypt the disk image, takes 0.37 seconds. To process 1B recoveries a year, or 123K recoveries per hour, we estimate that we would need 3,100 SoloKeys. In a SafetyPin deployment of 3,100 HSMs, tolerating the compromise of 6% of the HSMs (i.e., 194 HSMs), the client must interact with a cluster of 40 HSMs during recovery. Running our backup-recovery protocol across a cluster of this size takes 1.01 seconds.

**Limitations.** A limitation of SafetyPin is that the set of HSMs a device uses for recovery can leak information about the user’s PIN. In particular, an attacker who controls the data center can learn a salted hash of the user’s PIN during recovery. This is unfortunate in the common case that people re-use the same PIN after recovery [24, 44, 33, 74]. We discuss one mitigation in Section 8. Also, while it is possible to detect when PINs can safely be re-used, we have not yet implemented this functionality.

In addition, SafetyPin is more expensive than today’s PIN-based backup systems. SafetyPin requires the data center operator to operate a much larger fleet of HSMs (roughly 50–100x larger) than the standard HSM-based backup systems require. SafetyPin clients must also download roughly 2MB of keying material per day in a SafetyPin deployment supporting one billion recoveries per year, due to the periodic rotation of large HSM keys. Even so, we expect that the cost of storing and transferring disk images (GBs/user) will dwarf these costs.

## 2 The setting

**Entities.** Our encrypted-backup system involves three entities, whose roles we describe here.

**Client.** Initially, the client holds (1) a username with the service provider, (2) a human-memorable passphrase or PIN, (3) a disk image to be backed up, and (4) the public keys of the service provider’s HSMs. Later on, the client should be able to recover her backed-up data using only her username, her PIN, and access to the other components of the backup system.

In SafetyPin, as in today’s PIN-based backup systems, security depends on the client having access to the HSMs’ true public keys: If a malicious service provider can swap out the HSMs’ true public keys for its own public keys without detection, the service provider can immediately break security. Using a distributed log (Section 6) can ensure that all clients see a common set of HSM public keys, to prevent targeted attacks. Hardware-attestation techniques, as used in the FIDO [70] and SGX [46] specs, can provide another defense.

We also assume the provider has traditional account authentication (e.g., Gmail passwords) to prevent random third parties from consuming PIN guesses, but we omit this from the discussion for simplicity.

**Service provider.** The service provider offers the encrypted-backup service to a pool of clients and it maintains the data centers in which the backup system runs. For example, the service provider could be a mobile-phone vendor, such as Apple or Google. The service provider’s data centers contain the network infrastructure that connects the HSMs. They also contain large amounts of (potentially untrustworthy) storage and computing resources. Our security properties will hold against a service provider that becomes compromised at any point after the system is set up.

**Hardware security modules (HSMs).** The service provider’s data centers contain thousands of hardware security modules. An HSM is a tamper-resistant computing device meant for storing cryptographic secrets. HSMs have fully programmable
processors but are typically resource-poor (see Table 2). It is possible to lock an HSM’s firmware before deployment, which makes remote compromise and key-extraction attacks more difficult. Each HSM has a public key and stores the corresponding secret key in its secure memory.

The attack scenario. The service provider (Apple, Google, etc.) spends vast amounts of money acquiring a large user base for products that store user data in the cloud. The provider risks reputational damage and journalistic scrutiny if it cannot ensure the durability and confidentiality of user data.

A service provider can deploy SafetyPin as a way to build trust among its user base and to protect its own infrastructure against future compromise. By enlisting third-party organizations to monitor the SafetyPin deployment’s public distributed log, the provider can build further public trust in the system. At some point after the provider deploys SafetyPin, a powerful attacker wishes to steal user data. The attacker may have malicious insiders working for the provider. It may physically compromise data centers to steal HSMs. It may intercept shipments to tamper with some of the HSMs on their way to the data center. The attacker could also be a state actor employing legal pressure to gain access to data centers. Nonetheless, the attacker is sensitive to both the cost of attacks and the risk of public exposure.

Both the attack cost and risk of exposure increase with the number of HSMs the attacker must compromise. For instance, while a malicious insider working at a data center may be able to eavesdrop with a single HSM—passing the missing device off as a hardware failure—removing 100 HSMs is a much riskier proposition. A state actor who can order the provider to hand over HSMs may be dissuaded if doing so will attract press coverage either by making non-targeted clients’ data unrecoverable or creating a damming public audit trail.

The attacker may compromise clients as well as the provider. For instance, the attacker may have a good guess at a target user’s PIN, perhaps because of CCTV footage showing the user unlocking a mobile device. While SafetyPin cannot prevent the attacker from gaining access to the data with the correct PIN, the risk will be higher to the attacker if stolen PINs cannot be used without exposing the attack in SafetyPin’s public distributed log.

Table 2: Hardware security modules offer physical security protections but are computationally weak compared to a standard CPU.

| Device            | Price | $^2/sec | Storage | FIPS |
|-------------------|-------|---------|---------|------|
| SoloKey [77]      | $20   | 8       | 256 KB* |      |
| YubiHSM 2 [89]    | $650  | 14      | 126 KB  |      |
| SafeNet A700 [72] | $18,468 | 2,000   | 2,048 KB| ✔    |
| Intel i7-8569U (CPU) | $431 | 22,338 | n/a     |      |

The $^2/sec is NIST P256 elliptic-curve point-multiplications per second. "FIPS-2" refers to whether the device meets the FIPS 140-2 standard for HSMs. (* The 256 KB storage on the SoloKey is shared between code and data.)

Notation. The set $\mathbb{Z}^{>0}$ refers to the set of natural numbers \(\{1, 2, 3, \ldots\}\). For a positive integer \(n\), we let \([n] = \{1, \ldots, n\}\) and we use \(\perp\) to denote a failure symbol. For strings \(a\) and \(b\), we write their concatenation as \(a\|b\). Throughout, we use \(\lambda\) to denote the security parameter, and we typically take \(\lambda = 128\) (i.e., for 128-bit security).

3 System goals

SafetyPin implements an encrypted-backup functionality, which consists of two routines:

- the backup algorithm, which the client uses to produce its encrypted backup, and
- the recovery protocol, in which the client uses HSMs to recover the backup plaintext from ciphertext.

We define these protocols with respect to a number of HSMs \(N \in \mathbb{Z}^{>0}\) and a finite PIN space \(\mathcal{P} \subseteq \{0, 1\}^*\). For convenience, we define the master public key \(mpk\) for a data center to be all \(N\) HSMs’ public keys: \(mpk = (pk_1, \ldots, pk_N)\). The syntax of an encrypted-backup system is then as follows:

Backup\((mpk, user, pin, msg) \rightarrow ct\). Given the master public key \(mpk\), a client username \(user\), the client’s PIN \(pin\) ∈ \(\mathcal{P}\), and a message \(msg \in \{0, 1\}^*\) to be backed up, output a recovery ciphertext \(ct\). This routine runs on the client and requires no interaction with HSMs. The client uploads the resulting ciphertext \(ct\) to the service provider.

Recover\(^S.H_1,\ldots,H_N\)\((mpk, user, pin, ct) \rightarrow msg \text{ or } \perp\). The client initiates the recovery routine, which takes as input the master public key \(mpk\), a client username \(user\), a PIN \(pin\) ∈ \(\mathcal{P}\), and a recovery ciphertext \(ct\). During the execution of Recover, the client interacts with the service provider \(S\) and a subset of the HSMs \(H_1, \ldots, H_N\). Each HSM \(H_i\) holds the master public key \(mpk\), and its secret decryption key \(sk_i\). During recovery, the data center provides the client’s username user to each HSM.

The recovery routine outputs a backed-up message \(msg \in \{0, 1\}^*\) or a failure symbol \(\perp\).

We now describe the security properties that such a system should satisfy. We work in an asynchronous network model; we use standard cryptographic primitives to set up authenticated and encrypted channels between the client, service provider, and HSMs.

Property 1: Security. If the client obtains the HSMs’ true public keys, then even an attacker that:

- controls the service provider (in particular, is an active network attacker inside the data centers and has control of the service provider’s servers and storage),
- compromises an \(f_{\text{secret}}\) (e.g., \(f_{\text{secret}} = \frac{1}{16}\)) fraction of HSMs in the data center before the client begins the recovery process, and
4 Architecture overview

We now describe our encrypted-backup protocol (Figure 3) and explain how it satisfies the design goals of Section 3. We will discuss possible extensions and deployment considerations in Section 8.

4.1 The back-up process

The client begins the back-up process holding:

• the public keys of all HSMs in the data center,
• its secret PIN, and
• a disk image to be backed up (the “message”).

To back up its disk image, the client samples a subset of \( n \) HSMs out of the \( N \) total HSMs in the data center where \( n \ll N \). The client chooses this subset by hashing (a) public information: the service name, its username, and a public salt the client chooses at random, and (b) its secret PIN. The client then encrypts its message with a random AES encryption key, and then splits this AES key into \( n \) threshold shares using Shamir secret sharing [73], such that any threshold \( t \) of the shares suffice to recover the AES key. The client prepends each share with the client’s username to ensure that the ciphertexts are bound to the client’s username. The client then encrypts one share to the public key of each HSM in its chosen subset.

The client’s recovery ciphertext then consists of: its public salt, the AES-encrypted message, the \( n \) encrypted shares of the AES key, and a configuration-epoch number that the service provider can use to identify the set of HSMs that were in service at the time the client created its backup. The client computes the ciphertext locally and uploads it to the backup service provider, with no HSM interactions required.

To explain why this construction is scalable: since only a constant number of HSMs \( n \ll N \) participate in the decryption process, the system scales well as the number of HSMs in the data center increases.

To explain why this construction should be secure: if the attacker cannot guess the client’s PIN, the attacker does not know which set of \( n \) HSMs (out of the \( N \) total) it needs to compromise to recover the client’s AES key. So, the best attacks are either to: guess the client’s PIN or compromise a large fraction of the data center.

This argument requires that each individual key-share ciphertext leak no information about which HSM can decrypt it—a cryptographic property known as “key privacy” [7]. However, even key-private encryption schemes do not always remain secure against an adversary that adaptively compromises secret keys, which leads to our first technical challenge:

**Challenge 1.** How can we ensure that the client’s recovery ciphertext “leaks nothing” about which HSMs are required to decrypt the client’s message, even against an attacker who can adaptively compromise HSMs?

In Section 5, we explain how to solve this problem using location-hiding encryption, a new cryptographic primitive.

4.2 The recovery process

The client begins the recovery process holding:

• the public keys of all HSMs in the data center,
• its secret PIN, and
• its recovery ciphertext (which the client can fetch from the service provider).

• compromises all of the HSMs in the data center after the recovery protocol completes,

still should learn nothing about any honest client’s encrypted message (in a semantic-security sense [36]) beyond what it can learn by guessing that client’s PIN.

**Discussion:** The adversary can inspect all clients’ recovery ciphertexts and then choose to compromise a large set of HSMs that depends on those ciphertexts. Such attacks are relevant when, for example, a state actor with the power to compromise many HSMs targets the backed-up data of a specific set of users.

Two important caveats are: (1) SafetyPin does not protect against an attacker compromising HSMs while recovery is in progress (see Figure 4) and (2) as implemented, SafetyPin does not protect the PIN: an adversary that observes which HSMs the client contacts during recovery may learn a salted hash of the PIN after recovery completes. Section 6.3 discusses how to detect and mitigate this leakage by protecting the salt.

**Property 2: Scalability.** The recovery protocol should require the client to interact with a constant number of HSMs, independent of the number of HSMs in the data center. (This constant may depend on the security parameter and on the fraction of HSMs whose compromise the system can tolerate.) Hence, providers can deploy additional HSMs to scale capacity. Concretely, when we configure the system to tolerate the compromise of \( f_{\text{secret}} = \frac{1}{16} \) of the data center’s HSMs, our protocol requires the client to communicate with 40 HSMs during recovery.

**Property 3: Fault tolerance.** Every client should be able to recover her encrypted message even if a constant fraction \( f_{\text{live}} \) (e.g. \( f_{\text{live}} = \frac{1}{64} \)) of the HSMs in the data center fail-stop.

**Setting parameters.** For the remainder of this paper, we set the fraction of compromised HSMs that the system can tolerate to \( f_{\text{secret}} = \frac{1}{16} \) and the fraction of HSMs that can fail while still allowing the client to recover her backup to \( f_{\text{live}} = \frac{1}{64} \). This choice is reasonable because large companies have more than 16 data centers, while smaller companies can collaborate on a shared deployment with 16 physical security perimeters. By adjusting the other parameters, it is possible to achieve any \( 0 < f_{\text{secret}} < 1 \) or \( 0 < f_{\text{live}} < 1 \). (In Section 9.2, we discuss how the choice of these values affects other system parameters.)
First, the client asks the service provider to record its recovery attempt in the append-only log, implemented collectively by the service provider and HSMs. The log holds a mapping of identifiers to values. The service provider can insert new identifier-value pairs into the log but the service provider cannot modify or delete the values of defined identifiers, ensuring that there is at most one immutable value for each identifier.

The recovery attempt is logged as follows. The client begins by using public information (service name, username, and salt in the recovery ciphertext) along with its secret PIN to recover the subset of $n$ HSMs it picked during backup. The client then hashes these values together with some randomness to produce a cryptographic commitment $h$ to the identities of these HSMs and to its recovery ciphertext. The client then asks the service provider to insert the identifier-value pair $(\text{user}, h)$ into the log, where user is the client’s username. (In this discussion, we use the client’s username as the key for simplicity. In practice, to preserve privacy, we might use an opaque device-install UUID.)

The service provider collects a batch of these log-insertion requests, produces a Merkle-tree digest over the updated log, and runs a log-update protocol with the HSMs. At the end of this protocol, the HSMs hold the updated log digest. The service provider then returns to the client a Merkle proof $\pi$ proving that the pair $(\text{user}, h)$ appears in the latest log digest.

Since the service provider and HSMs run the log-update protocol periodically (e.g., every 10 minutes), the client will have to wait a few minutes on average to decrypt its backup. The client already has to download its large encrypted disk image, which will likely take minutes, so these steps can proceed in parallel.

The client then contacts its chosen set of $n$ HSMs over an encrypted channel, such as TLS. The client sends to each HSM: its username, the opening of its commitment $h$ (i.e., the values and randomness used to construct the commitment $h$), and the Merkle inclusion proof $\pi$. Each HSM

- recomputes the commitment $h$ and checks the inclusion proof $\pi$ (to confirm that the recovery attempt is logged), and
- decrypts its share of the client’s AES key, confirms that the username in the decrypted plaintext matches the one provided by the client (which prevents user $A$ from attempting to decrypt user $B$’s ciphertext, in collusion with a malicious service provider).

If both of these checks pass, the HSM returns the AES-key share to the client.

Given any $t$ of these decryption-key shares, the client can recover the AES key used to encrypt its backup. The client can then use this AES key to decrypt its backed-up message.

Since at most one log entry can exist per username, the use of the log ensures that each user can make at most one recovery attempt. In this way, the system defeats brute-force PIN-guessing attacks. With a slight modification, it is possible to allow each user to make a fixed number (e.g., 5) guesses, or a fixed number of guesses per time period (e.g., 5 per month).

A counter-intuitive property of this scheme is that the client never explicitly provides its PIN to the HSMs. The fact that the client knows which subset of the HSMs to contact implicitly proves the client’s knowledge of the PIN because the set of $n$ HSMs is much smaller than the total number of HSMs $N$.

This overview leaves some technical details unexplained. In particular:

**Challenge 2. How do the HSMs implement the append-only log without sacrificing scalability or security?**

A straightforward way to implement the log would be to have each HSM store the entire state of the log. But then every HSM would have to participate in every recovery attempt, which would not meet our scalability goals. Another implementation would be to have the data-center operators maintain the log, but then malicious data centers could violate the append-only
property, and thus mount brute-force PIN-guessing attacks, without HSMs noticing.

In Section 6, we explain how the HSMs can collectively maintain such an append-only log in a scalable and secure manner. At a high level, the (potentially adversarial) data center maintains the state of the log, which we represent as a list of identifier-value pairs. Every time the data center wants to insert an identifier-value pair into the log, the data center must prove to a random subset of the HSMs that the identifier to be inserted is undefined in the current log. Provided that at least one honest HSM audits each log-insertion, we can guarantee that the values associated with log identifiers are immutable (i.e., that we maintain the log’s append-only property). In this way, (a) each HSM needs to participate in only a vanishing fraction of the recovery attempts and (b) even an attacker who can compromise many of the HSMs cannot break the append-only nature of the log.

One remaining issue is that an attacker who observes the data center network may see which HSMs a client interacts with during recovery and decide to compromise that exact set of HSMs after recovery completes.

**Challenge 3.** *For scalability, the client should only communicate with a small number of HSMs during recovery. But then how can we protect against an attacker who compromises these HSMs after recovery completes?*

Our idea is as follows: after a client runs the recovery protocol, each participating HSM revokes its ability to decrypt that client’s recovery ciphertext. So, even if an after-the-fact attacker compromises the HSMs that participated in recovery, the attacker learns no useful information. The only window of vulnerability is at the moment after the client contacts its HSMs and before the HSMs complete revocation (Figure 4). Making this work on resource-limited HSMs requires new technical tools, which we describe in Section 7.

## 5 Protecting the mapping of users to HSMs with location-hiding encryption

In this section, we define and construct location-hiding encryption, which the client uses to encrypt its backup data.

The location-hiding encryption routine takes as input (1) a set of $N$ public keys, (2) a short PIN, and (3) a message, and outputs a ciphertext. In our application, the $N$ public keys are the public keys of the $N$ HSMs in the data center.

The cryptosystem has three main properties, which we formalize in Appendix A:

1. **Security.** To successfully decrypt the ciphertext, an attacker must either (a) guess the PIN or (b) control more than a constant fraction $f_{\text{secret}}$ of the $N$ total secret keys. This security property must hold even if the attacker can adaptively compromise an $f_{\text{secret}}$ fraction of the $N$ secret keys. In our application, this implies that unless an adversary can guess the PIN or compromise a constant $f_{\text{secret}}$ fraction of the HSMs in the data center, it learns nothing about the client’s backed-up data.

2. **Scalability.** Given the PIN used to encrypt the message, it is possible to decrypt the message using a small subset of the $N$ secret keys corresponding to the $N$ public keys used during encryption. In our application, a client who knows the correct PIN can recover its backup by interacting with only a small cluster of $n$ HSMs (for some parameter $n \ll N$) out of the $N$ total HSMs. So as $N$ grows, each HSM needs to participate in a vanishing fraction of the total recovery attempts.

3. **Fault tolerance.** Given the PIN, it is possible to decrypt a ciphertext even if a random fraction $f_{\text{live}}$ of all secret keys are unavailable. In our application, this implies clients can recover their backups even if an $f_{\text{live}}$ fraction of all HSMs fail.

We call this primitive “location-hiding encryption” because there is a small set of $n$ HSMs that the attacker could compromise to decrypt the ciphertext, but the cryptosystem hides the location of these HSMs within the larger pool of $N$ HSMs.

## Our construction

Our construction of location-hiding encryption is just a careful composition of existing primitives. However, it takes some analysis to prove that the composition provides the desired security properties. We describe our construction here in prose and we include the security definitions and proofs in Appendix A. The construction makes use of a public-key encryption scheme (hashed ElGamal encryption [28, 15]) and an authenticated encryption scheme (e.g., AES-GCM).

**Setup.** In our construction, each HSM $i$, for $i \in [N]$, holds a keypair $(pk_i, sk_i)$ for the public-key encryption scheme. Let $t \in \mathbb{Z}^{\geq 0}$ be a threshold such that if each HSM fails with probability $f_{\text{live}}$, then in a random sample of $n$ HSMs, there are at least $t$ non-failed HSMs with extremely high probability. Our instantiation takes $t = n/2$ for $f_{\text{live}} = \frac{1}{64}$.
Encryption. The encryption routine takes as input a list of $N$ public keys $(p_{k1}, \ldots, p_{kN})$, a PIN, and a message $msg$. To encrypt the message using our location-hiding encryption scheme:

1. Sample a random AES key $k$ and a random salt.
2. Split $k$ into $t$-out-of-$n$-Shamir secret shares $k_1, \ldots, k_n$ [73].
3. Hash the PIN and salt and use the result as a seed to generate a list of $n$ random indices $i_1, \ldots, i_n \in [N]$.
4. Encrypt each key-share $k_j$ with public key $p_{kj}$.
5. Finally, return (a) the salt, (b) the $n$ public-key ciphertexts, and (c) the AES encryption of $msg$ under key $k$.

Decryption. To decrypt given the ciphertext and PIN:

1. Hash the salt and PIN to reconstruct the set of indices $i_1, \ldots, i_n \in [N]$ used during encryption.
2. Use secret keys $sk_{i_1}, \ldots, sk_{i_n}$ to decrypt the $n$ shares of the AES key $k$. (In fact, only $t$ of the shares are necessary.)
3. Using the recovery routine for Shamir secret sharing, recompute the AES key $k$ from its shares.
4. Decrypt and return $msg$ using the AES key $k$.

Notice that the decryption routine only uses the PIN to sample the set of secret keys used for decryption. In our application, this implies that the client never needs to explicitly provide its PIN (or even a hash of its PIN) to the HSMs; contacting the right subset of HSMs is enough to ensure that the client provided the correct PIN.

The intuition behind the security analysis is straightforward: with hashed ElGamal encryption, the ciphertext reveals no information about which $n$ public keys (out of the $N$ total where $n < N$) were used during encryption. Thus, the ciphertext reveals no information about which secret keys the attacker must compromise unless the attacker can guess the PIN. Without these secret keys, the attacker cannot learn anything about $k$, and therefore cannot decrypt the message.

The following theorem, which we prove as Theorem 10 in Appendix A.6 makes this argument precise:

**Theorem (Informal).** The location-hiding encryption scheme of Figure 15 instantiated with the hashed ElGamal encryption scheme (Appendix A.4) over a group $\mathbb{G}$ is secure (in the sense of Definition 3) for certain values of $n$ and $N$, provided that:

- the computational Diffie-Hellman problem is hard in $\mathbb{G}$,
- the authenticated-encryption scheme is secure, and
- we model the hash functions used in the construction as random oracles.

There are two reasons why the security analysis is non-trivial: First, we must ensure that the ciphertext leaks nothing about the $n$ keys to which it was encrypted (i.e., that it is key-private [7]). Second, we must ensure that the encryption scheme remains secure even if an attacker can adaptively compromise secret keys. This is known as security under selective-opening attack [8, 30, 45]. Showing that both properties hold at once is the source of the technical complexity.

6 The distributed log

In SafetyPin, the HSMs collectively maintain a distributed log, which any external party can read and replay. The service provider maintains the log state and the HSMs monitor log insertions to ensure that the service provider does not violate the log’s append-only property.

We use this log for two primary purposes:

1. **Limiting PIN guesses.** To prevent an attacker from brute-force guessing a client’s PIN, we use the log (as described in Section 4) to enforce a global limit on the number of recovery attempts that the HSMs allow per username.
2. **Monitoring recovery attempts.** The service provider logs each recovery attempt, so any SafetyPin client can inspect the log to learn whether someone (e.g., a foreign attacker or snooping acquaintance) has tried to recover their backed-up data. A client could then take mitigating action—such as contacting their service provider, a law-enforcement agency, or the press.

A third use for the log—which comes directly from related work [31] and which we have not yet implemented—is to manage HSM group membership. Whenever the service provider wants to add or remove an HSM from the data center, the service provider operator could record this information in the log before the other HSMs will accept the change. All SafetyPin clients can thus verify that they are communicating with the same set of HSMs. In addition, clients can also detect suspicious changes in the set of HSMs in the data center. (For example, if the service provider replaces all HSMs in the data center over the course of a day.)

The log is simply a list of identifier-value pairs maintained by the service provider. Clients can insert identifier-value pairs in order to record recovery attempts, and HSMs maintain a digest of the log state. Our distributed log must satisfy the following key property:

*If any honest HSM ever accepts that an identifier-value pair $(id, val)$ is included in the log, the HSM should never accept that $(id, val')$ is included in the log for any value $val' \neq val$.*

6.1 Underlying data structure

**Terminology.** The log $L$ is a list of key-value pairs. Since we use the word “key” in this paper to refer to cryptographic keys, we call log keys “identifiers.” We say that a log $L'$ “extends” a log $L$ if (a) $L$ is a prefix of $L'$ and (b) every identifier in $L'$ appears at most once.

Our distributed log uses an authenticated data structure [80, 67, 82] that implements the following five routines:
Choose we now explain how to use the primitives of Section 6.1 to build our distributed append-only log. Appendix B.2. At a very high level: the digest of the log is build our distributed append-only log.

### 6.2 Building a distributed log

We now explain how to use the primitives of Section 6.1 to build our distributed append-only log.

**Initializing the log.** The service provider maintains the entire state of the log \( L \). Each HSM stores a log digest \( d \) which, in steady state, is the digest of the log \( L \) that the service provider holds. Initially, the log \( L \) is empty and each HSM holds the digest of the empty log.

**Inserting into the log.** A client can insert an entry \((\text{id}, \text{val})\) into the log by simply sending the pair to the service provider. The service provider adds this entry to its log state \( L' \).

**Proving log membership to HSMs.** Before the HSMs allow a client to begin the recovery process, the HSMs require proof that the client’s recovery attempt is logged. Assume for the moment that the service provider holds a log \( L \) and all HSMs hold the up-to-date digest \( d = \text{Digest}(L) \). (We will explain how the HSMs get the latest log digest in a moment.) Then, a client can prove inclusion of any pair \((\text{id}, \text{val})\) in the log by asking the service provider for an inclusion proof. The service provider computes \( \pi_{\text{inc}} = \text{ProveIncludes}(L, \text{id}, \text{val}) \) and returns the inclusion proof to the client. The client then sends \((\text{id}, \text{val}, \pi_{\text{inc}})\) to the HSM, which can check \( \text{DoesInclude}(d, \text{id}, \text{val}, \pi_{\text{inc}}) \) to be convinced that \((\text{id}, \text{val})\) is in the log represented by its digest \( d \). This inclusion check is fast—logarithmic in the log length.

**Updating the log digest at the HSMs.** After a sequence of log-insertions, the service provider holds a log state \( L' \). The HSMs will be holding a digest \( d = \text{Digest}(L) \) of a stale log \( L \). If the service provider is honest, the new log \( L' \) extends the old log \( L \).

To update the log digest at the HSMs, the service provider will first send the new digest \( d' = \text{Digest}(L') \) to every HSM. Next, the data center must convince each HSM that this new digest \( d' \) represents a log that extends the log \( L \) that the old digest \( d \) represents.

One non-scalable way to achieve this would be for the service provider to send an extension proof \( \pi_{\text{Ext}} = \)
ProveExtends\((L, L')\) to every HSM. The problem is that the time required to check this extension proof grows linearly with the number of new log entries. So if every HSM checked the entire extension proof, the throughput of the system would not increase as the number of HSMs increases.

Instead, we use a randomized-checking approach, as in Figure 5. If there have been \(I\) insertions to the log since the last update, the service provider divides the updates into \(N\) chunks, each containing \(I/N\) insertions. The service provider then applies these chunks of updates to the old log \(L\) one at a time, producing a digest \(d_i\) and extension proof \(\pi_i\) for each of the \(N\) intermediate logs \((i \in \{1, \ldots, N\})\). The service provider then sends the root \(R\) of a Merkle-tree commitment to these digests to each HSM.

Each HSM then asks the service provider for a random \(\lambda\)-size subset of the intermediate digests and extension proofs, where \(\lambda\) is a security parameter. The service provider returns the requested digests and extension proofs and proves that these values are included in the Merkle root \(R\). Each HSM checks its requested intermediate extension proofs using \(\text{DoesExtend}(\cdot)\) and checks the Merkle proof relative to the root \(R\). The HSMs auditing the first and last chunks also ensure that the intermediate digests match the old digest \(d\) and the new digest \(d'\), respectively.

If these extension and Merkle proofs are valid, each HSM signs the tuple \((d, d', R)\) using an aggregate signature scheme [14], and returns the signature to the service provider. Once all online HSMs have signed, the service provider aggregates these signatures and broadcasts the aggregated signature to all HSMs. If any HSM fails during this process, the service provider notifies the HSMs and they restart this log-update process. (In Appendix B.3, we describe how the log can make progress even if HSMs fail during the log-update protocol.)

The HSMs check the aggregate signature on \((d, d', R)\) relative to the HSMs’ aggregate public key. If the signature is valid, the HSMs accept the new digest \(d'\).

**Security.** If there are at most \(f_{\text{secret}}\) compromised HSMs, then even if \(f_{\text{secret}}\) honest HSMs are slow, \((1 - 2f_{\text{secret}})N\) honest HSMs will participate in any successful protocol execution. If each of these HSM audits \(C\) chunks, then the probability that no honest HSM audits a particular log chunk is

\[
\Pr[\text{fail}] = (1 - \frac{1}{N})^{(1-2f_{\text{secret}})NC} \leq \exp((2f_{\text{secret}} - 1) \cdot C).
\]

(Here, we use the fact that \((1 - x) \leq \exp(-x)\).) If each HSM audits \(C = \lambda \approx 128\) chunks, this failure probability is \(\ll 2^{-128}\). In other words, some honest HSM will catch a cheating service provider with overwhelming probability. In addition, since all honest HSMs will expect a signature from all honest HSMs, this will cause the updating operation to fail and the system to halt. For this analysis, we assume that the adversary cannot adaptively compromise HSMs while the recovery protocol is running without taking them offline.

**Scalability.** Each HSM must check the extension proofs on \(\lambda\) chunks, where each chunk contains a \(1/N\) fraction of the total updates in each epoch. Thus each HSM checks a vanishing fraction \((\frac{1}{N})\) of log insertions. Each HSM checks one aggregate signature, which requires time independent of the number of HSMs [14]. Thus, the total work that each HSM performs per epoch decreases as the number of HSMs \(N\) increases.

Because we use the log primarily to limit the number of PIN attempts, garbage collection is straightforward. The service provider simply creates a new empty log, effectively resetting the number of PIN attempts for every user (old copies of the log can still be inspected to monitor recovery attempts). To ensure that the service provider does not run garbage collection and clear the state too frequently, each HSM will run garbage collection for a fixed number of times (e.g. the expected number of garbage collections over two years) before refusing to respond to further requests. This bounds the number of times the service provider can garbage collect the log.

### 6.3 Transparency and external auditability

Our log design allows anyone to audit the log to ensure that the service provider correctly maintains the log’s append-only property. Additional auditors only add to the security of the system by adding another layer of protection, as they can detect log corruptions in the event that more than \(f_{\text{secret}}\) HSMs are compromised. In particular, for any two log digests \(d\) and \(d'\), an auditor can ask the data center for the entire logs \(L\) and \(L'\) corresponding to both of these digests. The auditor confirms that \(d\) is the root of the log tree for \(L\) and that \(d'\) is the root of the log tree for \(L'\). Finally, the auditor checks that \(L'\) extends \(L\).

As an extra precaution, users could specify external parties (e.g., Let’s Encrypt) as designated auditors during backup. During recovery, the HSMs would only complete the recovery if these auditors sign the latest log digest. In this way, mounting a brute-force PIN-guessing attack against a user would require compromising the user’s external auditors as well.

The transparency log can also help with PIN re-use. As discussed in Section 8, instead of storing the salt directly with the service provider, the salt itself can be encrypted using a second round of location-hiding encryption and a null PIN. After recovery, the salt will be destroyed as discussed in the next section. Once the salt has been destroyed, the device restoring a backup can use the log to determine if anyone else has ever fetched the salt. If not, then it is safe for the user to re-use the old PIN.

As described in Section 4, the log contains usernames, which could be sensitive. To prevent leaking usernames, we would replace usernames with random device identifiers that are rerandomized when the device is factory reset. However, even with this modification, the log still leaks information about when and how often users restore backups, which the service provider may not wish to make public. While we hope that organizations would make their logs public, we acknowledge that some may only share their logs with several
hand-picked organizations for auditing or may not share their logs at all. In these cases, our security guarantees still hold, although some of the transparency benefits are lost.

7 Forward security by puncturable encryption

We would like our encrypted-backup system to provide forward secrecy [18]. During the recovery process, the client reveals the identity of the $n \ll N$ HSMs that can decrypt its backup. Without forward secrecy, an attacker can break into these $n$ HSMs to recover the client’s backed-up data. Forward secrecy ensures that after recovery, an attacker, even one who compromises all HSMs in the data center, learns no information about the client’s backup.

One seemingly straightforward way to provide forward secrecy would be to use a new keypair for each backup. However, because the client cannot interact with the HSMs it is encrypting to during backup (as this would reveal their identities), using a unique keypair for every backup would require every HSM in the data center to generate a new keypair for every backup, running counter to our scalability goals.

7.1 Background: Puncturable encryption

We instead achieve forward secrecy using puncturable public-key encryption [40, 41, 25, 21, 19, 26]. A puncturable encryption scheme is a normal public-key encryption scheme (KeyGen, Encrypt, Decrypt), with one extra routine:

\[ \text{Puncture}(sk, ct) \rightarrow sk_{ct}. \]

Given a decryption key $sk$ and a ciphertext $ct$, output a new secret key $sk_{ct}$ that can decrypt all ciphertexts that $sk$ could decrypt except for $ct$.

**Puncturable encryption for forward secrecy.** To achieve forward security in SafetyPin, after an HSM decrypts its share of a client’s recovery ciphertext $ct$, the HSM punctures its secret decryption key. The punctured key allows the HSM to decrypt all ciphertexts except for $ct$. Thus, if an attacker compromises all HSMs in the data center after a client has recovered its backup, the attacker will be unable to decrypt any backup images that clients have already recovered. Furthermore, if an attacker compromises at most $f_{\text{secret}} \cdot N$ HSMs total, where $f_{\text{secret}}$ is a parameter of the system that we define in Section 3, then the attacker will not be able to recover any backed-up data whatsoever.

**Existing tool: Bloom-filter encryption.** Our implementation uses a puncturable encryption scheme called Bloom-filter encryption [25]. There are only two details of Bloom-filter encryption that are important for this discussion.

1. **The secret key is large.** If a key supports $P \in \mathbb{Z}^{+\infty}$ punctures and we want decryption to fail with probability at most $2^{-\lambda}$, then the secret key for Bloom-filter encryption is an array of roughly $\lambda P$ elements of a cryptographic group $G$.
2. **Puncturing is simple.** Puncturing the secret key just requires deleting $\lambda$ elements in the data array that comprises the secret key.

Concretely, when we set the Bloom-filter-encryption parameters to suitable values for experimental evaluation, each Bloom-filter encryption secret key has size over 64 MB. Even high-end HSMs have only 1–2 MB of storage (Table 2), so storing such large keys on an HSM would be impossible.

7.2 Outsourced storage with secure deletion

We show how to efficiently outsource the storage of this large secret key in a way that preserves forward secrecy of the punctured key. In particular, the HSM can outsource the storage of its secret-key array to the untrustworthy service provider, while still retaining the ability to delete portions of the key. Our technique applies to outsourcing the storage of any data array—not just secret keys—so we describe our secure-deletion approach in general terms.

**Desired functionality.** At a high level, the HSM has access to (a) a small amount of internal storage and (b) a large external block store, run by the service provider. The HSM wants to store an array of $D$ data blocks at the provider (data$_1$, ..., data$_D$). The HSM should be able to subsequently read or delete these blocks.

The following security properties should hold, even if the attacker, controlling the service provider, may choose the data-array and sequence of operations the HSM performs:

- **Integrity.** If the service provider tampers with the stored data in a way that could cause a read to return an incorrect result, the read operation outputs ⊥. Otherwise, the read operation for a block $i$ returns the value of the last data that the client wrote to block $i$.
- **Secure deletion.** If the service provider compromises the HSM after the HSM has run the delete operation for the $i$th data block, the attacker learns nothing about the data stored in block $i$. (This property implies a confidentiality property: the service provider learns nothing about the outsourced data.)

For efficiency, the HSM storage requirements must be small (constant size) and the read and delete routines should run quickly (in time logarithmic in the size $D$ of the data array). Unlike in ORAM [34, 35], our goal is not to hide the HSM’s data-access pattern from the service provider. We aim only to hide the contents of the array.

7.3 Our secure outsourced storage scheme

We explain our construction here in prose. See Appendix C for a more formal description.

**Running the setup phase.** During the setup phase, our outsourced-storage scheme builds a binary tree with $D$ leaves. Every node of the tree contains a fresh symmetric encryption
key. During setup, for each node in the tree with key sk_i, we encrypt the keys of the child nodes sk_i0 and sk_i1 with sk_i and store this ciphertext AE.Encrypt(sk_i, sk_i0∥sk_i1) in outsourced storage. At the leaves of the tree, we encrypt the i-th data block with the key sk_i at the i-th leaf and we store the ciphertext AE.Encrypt(sk_i, data_i) in outsourced storage.

For example, in Figure 6, we use sk_0 to encrypt sk_00 and sk_01 and we store the result in outsourced storage. We use key sk_01 to encrypt data item 2. Thus, knowing the root key sk is enough to decrypt the entire tree and access every data element in the array.

Reading a data block. To retrieve the data block at index i, the HSM reads in the ciphertexts along the path from the tree root to leaf i. The HSM then decrypts the chain of ciphertexts from the root down to recover the data block at index i. For example, in Figure 6, to retrieve data block 3, the HSM can use sk_0 to decrypt sk_1, and sk_1 to decrypt sk_10, which it can use to decrypt data item 1.

Deleting a data block. To delete the data block at index i, the HSM recovers (as in retrieval) the keys along the path from the root to leaf i. At the node containing the key to decrypt data block i, the HSM deletes the key. It then chooses a fresh key and re-encrypts the other key at that node using the fresh key. To maintain the ability of the parent key to decrypt the child ciphertext, the HSM updates the parent of that node to contain the fresh key for its child and re-encrypts the parent’s keys under a new key. It continues this up the path to the root, where the HSM chooses a new key sk’ to encrypt the root. The HSM replaces sk with sk’, deleting the old sk, and then sends the new ciphertexts along the path from the root to leaf i back to the service provider. For example, in Figure 6, to delete data item 3, the HSM decrypts the keys (sk_0||sk_1) and (sk_10||sk_11). The HSM then deletes sk_10, chooses a new key sk’ to encrypt sk_11, and then chooses a new key sk’ to encrypt sk_0 and sk’_1. The HSM then replaces sk with sk’.

Efficiency. The setup time is linear in the size of the data array D. The runtimes of retrieval and deletion are both logarithmic in D, and require only symmetric-key operations. The HSM stores only the constant-sized root encryption key sk.

Security intuition. An HSM can always recover the keys necessary to decrypt a data item, provided the HSM did not previously delete any of the keys necessary for decryption. Integrity follows immediately from the security of the underlying authenticated encryption scheme. Finally, we ensure secure deletion by deleting the key necessary to decrypt a certain data item and updating the root key. Without the old root key, it is impossible to access the key necessary to decrypt the deleted data item.

Putting it together. To summarize: the HSMs use a puncturable encryption scheme to prevent the compromise of HSM secrets at time T from allowing an adversary to learn about backed-up data that was recovered any time before T. We implement puncturable encryption using Bloom-filter encryption and outsource the storage of the large secret decryption key using our new technique for outsourcing with secure deletion.

8 Extensions and deployment considerations

The full SafetyPin implementation has to deal with a number of additional issues, which we discuss now.

Failure during recovery. As discussed in Section 7, after participating in recovery, HSMs revoke their ability to decrypt the recovered ciphertext. One consequence is that a client cannot recover the same backup ciphertext twice. This raises the question of what happens if a replacement device fails during or shortly after recovery, or if a communication failure during recovery prevents the new device from receiving the replies from the HSMs.

To solve this problem, when a client initiates recovery, it first generates a fresh per-recovery keypair (sk, pk) for a public-key encryption scheme. The client backs up this secret key sk using SafetyPin before initiating its recovery. Next, the client sends the public key pk to each HSM and then begins the backup-recovery process. Each HSM encrypts its replies to the client under pk, and each HSM sends a copy of each reply to the data center. If a client device fails during recovery, a second, replacement client device can retrieve the backed-up secret key sk and use these to decrypt the replies stored at the data center. This scheme nests arbitrarily, thereby handling any number of consecutive device failures during recovery.

Incremental backups. In practice, mobile devices often generate incremental backups rather than encrypting the entire disk image for each backup. SafetyPin supports incremental backups in the following way. The user uses SafetyPin to store a single AES key, which the user also keeps on her phone. The user can then encrypt incremental backups under this AES key and upload the resulting ciphertext to the data center. When
the user recovers, she recovers her AES key and can use this key to decrypt the incremental updates.

Multiple recovery ciphertexts. Clients back up their phones regularly (e.g., every three days), and will thus generate a series of recovery ciphertexts. We want to ensure that after a client recovers her backup from time $t$, the HSMs involved in recovery puncture their secret decryption keys so that they cannot decrypt that client’s backups from earlier times $t' < t$, even if an attacker compromises all HSMs in the data center. To achieve this, in the puncturable-encryption step (Section 7.1), we have the client use the same salt for each recovery ciphertext it generates. In this way, the client will encrypt its series of backups to the same set of HSMs. When these HSM puncture their secret keys during the recovery process, they will destroy their ability to decrypt any previous recovery ciphertexts from the given client. After recovery, the client chooses a new salt to generate subsequent backups on its new device.

Preventing post-recovery PIN leakage. As we have discussed, an attacker that watches the client recover can learn a salted hash of the user’s PIN, which can be used to mount an offline brute-force attack to learn the user’s PIN.

One approach to protect against this attack would be to have each user store their salt in secret-shared form at a random set $S_{salt}$ of HSMs, where $S_{salt}$ is included in the client’s recovery ciphertext. Then, provided that the attacker does not compromise this set of HSMs, the attacker would learn no useful information on the user’s PIN, even after recovery. An attacker could always compromise every HSM in $S_{salt}$, but an attacker that can compromise only a $f_{secret}$ fraction of HSMs in the data center would not be able to mount this attack against too many clients’ salts. We hope to model and prove this multi-user PIN-protection property in future work.

## 9 Implementation and evaluation

We implemented SafetyPin on an experimental data cluster of 100 hardware security devices (Figure 1).

**HSM.** For the HSMs, we used SoloKeys [77], a low-cost open-source USB FIDO2 security key. SoloKeys use a STM32L432 microcontroller with an ARM Cortex-M4 32-bit RISC core clocked at 80MHz and 265KB of memory. The device is not side-channel resistant, but has a true random number generator and can lock its firmware. We add roughly 2,500 lines of C code to the open-source SoloKey firmware [76].

By default, SoloKeys communicate with the USB host via USB HID, an interrupt-based USB class used typically for keyboards and mice that has a maximum throughput of 64KBps. To improve performance, we rewrote parts of the firmware to use USB CDC, a high-throughput USB class commonly used for networking devices. This gave a roughly 32× increase in I/O throughput (Table 7).

For the puncturable-encryption scheme (Section 7.1), we use a variant of Bloom-filter encryption [25] that avoids the need for pairings [13] but increases the size of the HSMs’ public keys. For the aggregate signature scheme needed for the log, we use BLS-style multisignatures [12] over the JEDI [51] implementation of the BLS12-381 curve.

Our implementation does not encrypt communication between the client and HSMs. Based on the time to run AES-128 and ElGamal encryption on the SoloKeys, we estimate that transport-layer encryption would add two ElGamal decryptions and 2KB of AES operations per recovery, increasing recovery time by approximately 0.3 seconds, or 30%. This overhead is comparatively high because processing a recovery only requires a handful of symmetric and public key operations.

**Service Provider.** Our service provider host is a Linux machine with an Intel Xeon E5-2650 CPU clocked at 2.60GHz. Our service-provider implementation is roughly 3,800 lines of C/C++ code (excluding tests) and uses OpenSSL.

**Client.** Our client device is a Google Pixel 4. Our implementation is roughly 2,300 lines of C/C++ code (excluding tests) and uses OpenSSL.

| Operation    | Ops/sec | Operation    | Ops/sec |
|--------------|---------|--------------|---------|
| Pairing      | 0.43    | HMAC-SHA256  | 2,173.91|
| ECDSA ver    | 5.85    | AES-128      | 3,703.70|
| ElGamal dec  | 6.67    | RTT, HID (32b)| 71.43   |
| $g^x \in G_{P256}$ | 7.69    | RTT, CDC (32b)| 2,277.90|
|              |         | Flash read (32b)| ≈166,000|

Table 7: Microbenchmarks on SoloKey. Pairing is on BLS12-381 curve using the JEDI library [51]. Other public-key operations use NIST P256 curve.
with our scheme for outsourced storage with secure deletion. Each HSM can process 1,503.9 recoveries per hour on average. HSM spends roughly 56% of its cycles rotating its keys, and maintaining the log between key rotations. Therefore, each HSM spends approximately 139.4 hours processing recoveries and on the number of public-key operations required) that key has been deleted. Key rotation is expensive: we estimate (based on the size of the data center N and PINs with six decimal digits, and is dictated by Theorem 10. We set the puncturable encryption keys to allow 2^{20} punctures, as we found this provides a reasonable tradeoff between the time to decrypt and puncture and the time between key rotations. With these parameters, we maintain secrecy if at most an \( f_{secret} = \frac{1}{35} \) fraction of the HSMs are compromised (or \( f_{secret} \cdot N \approx 194 \) total). We allow data recovery if at most an \( f_{live} = \frac{1}{48} \) fraction fail due to benign hardware failures (or \( f_{live} \cdot N \approx 48 \) total).

Baseline. We compare against an encrypted-backup system modeled on the ones that Google and Apple use [87, 49]. To backup, the client selects a fixed cluster of five HSMs and encrypts her recovery key and a hash of her PIN under the cluster’s public key. At recovery, the client sends the recovery ciphertext and a hash of her PIN to the cluster, and any HSM in the cluster can decrypt the ciphertext, check that the PIN hashes match, and return the recovery key. To defeat brute-force PIN-guessing attacks, each HSM independently limits the number of recovery attempts allowed on a given ciphertext.

Client overhead. Figure 10 gives the overhead of generating a backup in SafetyPin, compared to the baseline. The backup process takes 0.37 seconds. SafetyPin recovery ciphertexts are 16.5KB, versus 130B for our baseline, though we expect encrypted disk image to dominate the ciphertext size.

SafetyPin increases the bandwidth cost at the client. In the baseline scheme, the client downloads five public keys—one from each of its five chosen HSMs. In SafetyPin, the client must fetch a copy of all HSMs’ public keys. (This way, the service provider does not learn the subset of HSMs to which the client is encrypting its backup.) So, when a client first joins the system, the client must download all these keys (11.5MB). Whenever an HSM rotates its puncturable-encryption keys, clients must download the HSM’s new public key. In a deployment of \( N = 3,100 \) HSMs supporting one billion recoveries annually, we estimate that each SafetyPin client must download 1.97MB of keying material daily. Increasing the puncturable encryption failure probability would decrease client bandwidth, although this would require decreasing the fraction of HSMs allowed to fail, \( f_{live} \). If a client goes offline for several days, it must download the rotated public keys for each day it spent offline (roughly 2MB/day), up to a maximum of 11.5MB (the size of all HSMs’ keys). However, the client only needs to store the public keys for the \( n \) HSMs comprising its chosen

9.1 Microbenchmarks

Log. Figure 8 demonstrates how increasing the number of HSMs reduces the log-digest update time. We assume that the log is periodically garbage collected (i.e., approximately once a month), so that it holds at most a hundred million recovery attempts at once. If the HSMs run the log-update process every 10 minutes, each HSM spends approximately 11% of its active cycles auditing the log. The choice of how often to update the log is a tradeoff between how long users must wait to recover their backups and the total number of write cycles to non-volatile storage permitted by the hardware.

Puncturable encryption. Figure 9 shows the cost of performing a decrypt-and-puncture operation as the number of supported punctures increases. The AES operations associated with our scheme for outsourced storage with secure deletion (Section 7.2) dominate the cost.

Another way to implement outsourced storage with secure deletion would be to have the HSM store the outsourced array encrypted under a single AES key \( k \). To delete an item, the HSM would read in the entire array, delete the item, and write out the entire array encrypted under a fresh key \( k' \). With this approach, a deletion takes 48 minutes for a 64 MB array (the size of our outsourced secret keys). Our scheme thus improves system throughput by roughly 4,423×.

Each HSM punctures its secret key (Section 7.1) once after each decryption it performs. Since our puncturable-encryption scheme only supports a fixed number of punctures, each HSM must periodically rotate its encryption keys. We configure our puncturable-encryption scheme to allow each HSM to perform roughly \( 2^{18} \) decryptions before it must rotate its keys (rotation is triggered when half of the elements of the secret key have been deleted). Key rotation is expensive: we estimate (based on the number of public-key operations required) that key rotation on our HSMs will take roughly 75 hours. Each HSM spends approximately 139.4 hours processing recoveries and maintaining the log between key rotations. Therefore, each HSM spends roughly 56% of its cycles rotating its keys, and each HSM can process 1,503.9 recoveries per hour on average.

9.2 End-to-end costs

Parameters. We estimate that on average, each user will run recovery once a year. (There are 3.8B smartphone users [78] and 1.5B smartphones sold annually [79], so we expect \( 1.5/3.8 = 0.39 \approx 1 \) recovery/user/year.) We calculate that a SafetyPin deployment of \( N = 3,100 \) HSMs could support one billion users. So, we treat our small cluster of 100 HSMs as a representative slice of a larger data center of \( N = 3,100 \) HSMs. Within this larger data center, each client shares its recovery keys among a cluster of \( n = 40 \) HSMs. This choice of \( n \) is based on the size of the data center \( N \) and PINs with six decimal digits, and is dictated by Theorem 10. We set the puncturable encryption keys to allow \( 2^{20} \) punctures, as we found this provides a reasonable tradeoff between the time to decrypt and puncture and the time between key rotations. With these parameters, we maintain secrecy if at most an \( f_{secret} = \frac{1}{35} \) fraction of the HSMs are compromised (or \( f_{secret} \cdot N \approx 194 \) total). We allow data recovery if at most an \( f_{live} = \frac{1}{48} \) fraction fail due to benign hardware failures (or \( f_{live} \cdot N \approx 48 \) total).

Baseline. We compare against an encrypted-backup system modeled on the ones that Google and Apple use [87, 49]. To backup, the client selects a fixed cluster of five HSMs and encrypts her recovery key and a hash of her PIN under the cluster’s public key. At recovery, the client sends the recovery ciphertext and a hash of her PIN to the cluster, and any HSM in the cluster can decrypt the ciphertext, check that the PIN hashes match, and return the recovery key. To defeat brute-force PIN-guessing attacks, each HSM independently limits the number of recovery attempts allowed on a given ciphertext.

Client overhead. Figure 10 gives the overhead of generating a backup in SafetyPin, compared to the baseline. The backup process takes 0.37 seconds. SafetyPin recovery ciphertexts are 16.5KB, versus 130B for our baseline, though we expect encrypted disk image to dominate the ciphertext size.

SafetyPin increases the bandwidth cost at the client. In the baseline scheme, the client downloads five public keys—one from each of its five chosen HSMs. In SafetyPin, the client must fetch a copy of all HSMs’ public keys. (This way, the service provider does not learn the subset of HSMs to which the client is encrypting its backup.) So, when a client first joins the system, the client must download all these keys (11.5MB). Whenever an HSM rotates its puncturable-encryption keys, clients must download the HSM’s new public key. In a deployment of \( N = 3,100 \) HSMs supporting one billion recoveries annually, we estimate that each SafetyPin client must download 1.97MB of keying material daily. Increasing the puncturable encryption failure probability would decrease client bandwidth, although this would require decreasing the fraction of HSMs allowed to fail, \( f_{live} \). If a client goes offline for several days, it must download the rotated public keys for each day it spent offline (roughly 2MB/day), up to a maximum of 11.5MB (the size of all HSMs’ keys). However, the client only needs to store the public keys for the \( n \) HSMs comprising its chosen

Figure 10: Breakdown of time to save (on Android Pixel 4 phone) and recover (using our SoloKey cluster). We do not consider the time to encrypt or decrypt disk images.
recovery cluster which amounts to 9.02KB.

**Recovery time.** At a cluster size of \( n = 40 \) HSMs, Figure 11 shows that the end-to-end recovery time takes 1.01 seconds. Puncturable-encryption operations dominate recovery time (Figure 10), since these require expensive elliptic-curve operations for ElGamal decryption and many I/O and AES operations in order to perform secure deletion (Section 7.2).

**Tail latency.** In a deployment of SafetyPin, it will be important to consider not only the average throughput of the SafetyPin cluster, but also the request latency. Since recovery requests will arrive concurrently and in a bursty fashion, we will need to overprovision the system slightly to ensure that request tail latency does not grow too high, even under large transient loads. In Figure 13, we model how many HSMs are required to achieve various 99th-percentile latencies, while handling different average throughputs. We compute these values by modeling incoming requests using a Poisson process and each HSM using a M/M/1 queue with service times derived from our experimental results. As the figure demonstrates, by increasing the total number of HSMs, we can reduce the tail latency even when accounting for request contention. We anticipate that recovery time will in practice be dominated by the time to download the encrypted disk image, and so as long as the tail latency is less than or close to this time, any delay is unlikely to be noticed by the user.

**Financial cost.** Figure 12 shows how throughput scales as the outlay on HSMs increases and Table 14 presents dollar-cost estimates for SafetyPin deployments with different types of HSMs. For a configuration that tolerates the compromise of 50 high-quality HSMs, we estimate that adding SafetyPin to an unencrypted backup system would increase the system’s dollar cost by 2.5%.

### Table 14: Estimated cost of storing 4GB \( \times 10^9 \) users per year

| HSM Qty. | \( f_{secret} \) | \( N_{evil} \) | Cost         |
|----------|-------------------|--------------|--------------|
| SoloKey  | 3,037             | 1/16         | 189          | $600M        |
| YubiHSM2 | 1,732             | 1/16         | 108          | $1.1M        |
| SafeNet A700 | 40             | 1/20         | 2            | $738K        |
| – 10 evil HSMs | 320           | 1/32         | 10           | $3.0M        |
| – 50 evil HSMs | 800          | 1/16         | 50           | $14.8M       |

**Estimated cost of storing 4GB \( \times 10^9 \) users per year:** $600M

### 10 Related work

Today’s encrypted-backup systems rely either on the security of hardware security modules [39, 49], secure microcontrollers [2], or secure enclaves [57, 60]. Vulnerabilities in these hardware components leave encrypted-backup systems open to attack. And there is ample evidence of vulnerabilities in both HSMs [69, 63, 66, 43, 48, 17, 32, 1] and enclaves [84, 16, 38, 55, 29, 85, 86, 64, 20, 42, 54, 11] and reason for concern about hardware backdoors as well [81, 88, 6, 50].

Many companies including Anchorage [4], Unbound Tech [83], Curv [23], and Ledger Vault [53], offer systems for secret-sharing cryptocurrency secret keys across multiple hardware devices. Unlike SafetyPin, these solutions use a small fixed set of HSMs, so they cannot simultaneously provide scalability and protection against adaptive HSM compromise.

Mavroudis et al. [59] propose building a single trustworthy hardware security module from a large array of potentially faulty hardware devices. To achieve this, they use cryptographic protocols for threshold key-generation, decryption, and signing. Like Myst, SafetyPin distributes trust over a large number of hardware devices. Unlike Myst, SafetyPin focuses on hardware-protected PIN-based encrypted backups, rather than more traditional HSM operations, such as decryption and signing.

In recent theoretical work, Benhamouda et al. show how to scalably store secrets on proof-of-stake blockchains when...
an adversary can adaptively corrupt some fraction of the stake [10]. They face many of the same cryptographic challenges that we tackle in Section 5; their theoretical treatment complements our implementation-focused approach. While they use proactive secret sharing to periodically re-share the secret and hide the secret from an adversary controlling some fraction of the stake, our approach allows a party with some low-entropy secret to recover the high-entropy secret.

Transparency logs inspire our log design [5, 52, 61, 37, 56]. While these logs allow a powerful auditor to verify correctness, they do not easily allow distributing the work of auditing across many less powerful participants. The proofs we provide to the HSMs about the state of the log draw on work on authenticated data structures [80, 58, 68] and cryptocurrency light clients [65]. Kaptchuk et al. show how public ledgers can be used to build stateful systems from stateless secure hardware [47], and they show how their techniques can be applied to Apple’s encrypted-backup system. This work is complementary to ours, as they show how to securely manage state in cases where HSMs do not have secure internal non-volatile storage (an assumption we make in SafetyPin).

11 Conclusion

SafetyPin is an encrypted backup system that (a) requires its users to only remember a short PIN, (b) defeats brute-force PIN-guessing attacks using hardware protections, and (c) provides strong protection against hardware compromise. SafetyPin demonstrates that it is possible to reap the benefits of hardware security protections without turning these hardware devices into single points of security failure.

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A Analysis: Location-hiding encryption

Additional notation. We use $x \leftarrow 7$ to indicate assignment. For a finite set $S$, we use $x \leftarrow S$ to denote taking a uniform random sample from $S$. The notation $\text{poly}(\cdot)$ refers to a fixed polynomial function and $\text{negl}(\cdot)$ refers to a fixed negligible function.

A.1 Syntax

We first define the syntax of a location-hiding encryption scheme. The scheme is parameterized by a total number of HSMs $N$, a cluster size $n \ll N$, a PIN space $\mathcal{P}$, a message space $\mathcal{M}$, and two constants:

(a) the fraction $f_{\text{live}}$ of the $N$ HSMs in a cluster whose benign failure we can tolerate while still allowing message recovery, and

(b) the fraction $f_{\text{secret}}$ of the $N$ total HSMs whose compromise we can tolerate while still providing security.

In our construction, we take $f_{\text{live}} = \frac{1}{64}$ and $f_{\text{secret}} = \frac{1}{16}$, but any constants $0 < f_{\text{live}}, f_{\text{secret}} < 1$ would suffice with an adjustment to the parameters.

Formally, a location-hiding encryption scheme consists of the following five algorithms:

**KeyGen**($1^\lambda$) $\rightarrow$ (pk, sk). On input security parameter $\lambda$, expressed in unary, output a public-key encryption keypair.

**Encrypt**(pk$_1$, ..., pk$_N$, salt, pin, msg) $\rightarrow$ ct. Given a list of $N$ public keys, generated using KeyGen, a randomizing salt salt $\in \{0, 1\}^*$, a PIN pin $\in \mathcal{P}$, and a message msg $\in \mathcal{M}$, output a ciphertext ct.

**Select**(salt, pin) $\rightarrow$ (i$_1$, ..., i$_n$). Given a salt and PIN, output the indices (i$_1$, ..., i$_n$) $\in [N]^n$ of the $n$ secret keys needed to reconstruct the encrypted message.

**Decrypt**(sk$_{i_j}$, i$_j$, ct) $\rightarrow$ $\sigma$$_j$. Given a ciphertext ct, a secret key sk$_{i_j}$ and an index i$_j$, produce the $j$-th secret share $\sigma$$_j$ of the plaintext message.

**Reconstruct**(\(\sigma_1, ..., \sigma_n\)) $\rightarrow$ msg. Given $n$ shares of the plaintext message msg returned by Decrypt, reconstruct msg.

To understand the syntax, it may be helpful for us to explain how we use these routines in our encrypted-backup application:

1. **Setup.** During device provisioning, each HSM runs (pk$_i$, sk$_i$) $\leftarrow$ KeyGen($1^\lambda$) to generate its keypair (pk$_i$, sk$_i$). The HSM stores the secret sk$_i$ in its internal memory and it publishes pk$_i$.

2. **Backup.** To back up its message msg with salt and PIN pin, given HSM public keys (pk$_1$, ..., pk$_N$) the client runs

   \[
   \text{ct} \leftarrow \text{Encrypt}(pk_1, \ldots, pk_N, \text{salt}, \text{pin}, \text{msg}).
   \]

   The client uploads its recovery ciphertext ct to the service provider.

3. **Recovery.** During recovery, the client fetches its salt and recovery ciphertext ct from the service provider. It then computes

   \[
   (i_1, \ldots, i_n) \leftarrow \text{Select}(\text{salt}, \text{pin})
   \]

   to identify the cluster of $n$ HSMs it must communicate with during recovery.

   For each HSM $i_j \in \{i_1, \ldots, i_n\}$, the client asks HSM $i_j$ to decrypt the ciphertext ct. (Our full protocol requires interaction with the service provider, but we elide those details here.) The HSM, who holds the secret key sk$_{i_j}$, computes:

   \[
   \sigma_j \leftarrow \text{Decrypt}(sk_{i_j}, i_j, \text{ct}).
   \]

   Finally, the client recovers its plaintext as

   \[
   \text{msg} \leftarrow \text{Reconstruct}(\sigma_1, \ldots, \sigma_n).
   \]

A.2 Definitions

Intuitively, the correctness property states that if a random $f_{\text{live}}$ fraction of the secret keys are unavailable, decryption still succeeds.

**Definition 1** (Correctness). Formally, we say that a location-hiding encryption scheme on parameters $(N, n, \mathcal{P}, \mathcal{M}, f_{\text{live}}, f_{\text{secret}})$ is correct if, for all PINs pin $\in \mathcal{P}$, all messages msg $\in \mathcal{M}$, on security parameter $\lambda \in \mathbb{Z}_{>0}$ Experiment 2 outputs 1, except with probability $\text{negl}(\lambda)$.

This notion of correctness only guarantees message reconstruction if some shares $\sigma_i$, computed using the Decrypt routine, are deleted or unavailable. We do not consider the stronger notion of correctness, in which message reconstruction is possible even if some of the shares $\sigma_i$ are corrupted (rather than just missing).

The security property for location-hiding encryption states that no efficient adversary can distinguish the encryption of two chosen messages with probability much better than guessing the PIN after roughly $N$ guesses. This property should hold even if the adversary may adaptively corrupt up to $f_{\text{secret}} \cdot N$ of the $N$ total secret keys.
Experiment 2 (Location-hiding encryption: Correctness). We define the following correctness experiment, which is parameterized by a location-hiding encryption scheme with parameters \((N,n,P,M,f_{\text{live}},f_{\text{secret}})\), a PIN pin \(\in P\), a message \(msg \in M\), and a security parameter \(\lambda \in \mathbb{Z}^{>0}\).

The experiment consists of the following steps:

- \((pk_i,sk_i) \leftarrow \text{KeyGen}(1^\lambda)\), for \(i = 1,\ldots,N\)
- salt \(\leftarrow \{0,1\}^t\)
- \(ct \leftarrow \text{Encrypt}(pk_1,\ldots,pk_N,\text{salt},\text{pin},\text{msg})\)
- \(F \leftarrow \emptyset\)
  - Add each element of \(\{1,\ldots,N\}\) to \(F\) with independent probability \(f_{\text{live}}\).
- \(\{i_1,\ldots,i_n\} \leftarrow \text{Select}(\text{salt},\text{pin})\)
- \(\sigma_j = \begin{cases} \bot & \text{for } j \in F \\ \text{Decrypt}(sk_{i_j},i_j,ct) & \text{otherwise} \end{cases}\)
- \(msg' \leftarrow \text{Reconstruct}(\sigma_1,\ldots,\sigma_n)\)

The output of the experiment is 1 if \(msg = msg'\) and 0 otherwise.

For a location-hiding encryption scheme on parameters \((N,n,P,M,f_{\text{live}},f_{\text{secret}})\), an adversary \(A\), and a security parameter \(\lambda \in \mathbb{Z}^{>0}\), let \(W_{A,i}\) denote the probability that the adversary outputs “1” in Experiment 4 with bit \(\beta \in \{0,1\}\). Then we define the advantage of \(A\) at attacking a location-hiding encryption scheme \(E\) as:

\[
\text{LHEncAdv}[A,E](\lambda) = |W_{A,0} - W_{A,1}|.
\]

Definition 3 (Security). Let \(E\) be a location-hiding encryption scheme, on parameters \((N,n,P,M,f_{\text{live}},f_{\text{secret}})\), such that \(N = \text{poly}(\lambda), n = \text{poly}(\lambda), \) and \(|P|\) and \(|M|\) grow as (possibly superpolynomial) functions of \(\lambda\). Then we say that the location-hiding encryption scheme \(E\) is secure if, for all efficient adversaries \(A\),

\[
\text{LHEncAdv}[A,E](\lambda) \leq O(N/|P|) + \text{negl}(\lambda).
\]

Remark 5 (Understanding the security definition.). Since our security definition allows the adversary to corrupt up to \(f_{\text{secret}} \cdot N\) of the secret keys, the adversary can always reconstruct the plaintext with probability roughly \(f_{\text{secret}} \cdot N / \text{poly}(\lambda)\) using the following attack to decrypt a ciphertext \(ct\) with salt \(\text{salt}\):

- Pick two messages \(msg_0, msg_1 \in M\).
- Pick a candidate PIN \(\text{pin'} \in P\).
- Run \(I = \{i_1,\ldots,i_n\} \leftarrow \text{Select}(\text{pin'},\text{salt})\).
- Corrupt the keys in \(I\) and use them to decrypt \(ct\).

- If \(ct\) decrypts to either \(msg_0\) or \(msg_1\), guess the corresponding bit.
- Otherwise, try again with another PIN.

If the attacker can corrupt \(f_{\text{secret}} \cdot N\) keys, it can try at least \(f_{\text{secret}} \cdot N/n\) PINs and its success probability is at least \(f_{\text{secret}} \cdot N/(n|P|)\). Our security analysis shows that this is nearly the best attack possible against our location-hiding encryption scheme, up to low-order terms.

Remark 6 (Choosing-ciphertext security). A stronger and more realistic definition of security would allow the adversary to make decryption queries to the HSMs in the penultimate stage of the attack game, as in the standard chosen-ciphertext attack (CCA) security game [71, 22]. Since our underlying public-key encryption scheme (hashed ElGamal) is already CCA secure, we believe that—at the cost of some additional complexity in the definitions and proofs—it would be able to achieve a CCA-type notion for location-hiding encryption.

A.3 A tedious combinatorial lemma

We will need the following lemma about random sets to prove Theorem 10. The proof of this lemma uses no difficult ideas,
but keeping track of the constants involved is a bit tedious.

**Definition 7.** Let $n, N$, and $\Phi$ be positive integers. Say that a set $S \subseteq [N]$ “$n/2$-covers” a list $L \in [N]^n$ if at least $n/2$ elements of the list $L$ appear in the set $S$.

Then, let $\text{Covr}_{n,\Phi}(\alpha, \beta)$ denote the probability, over the random choice of lists $L_1, \ldots, L_\Phi \sim [N]^\Phi$, that there exists a set $S \subseteq [N]$ of size $\alpha N$ that $n/2$-covers more than $\beta N$ of the lists.

**Lemma 8.** For $N > en \approx 2.71n$ and $\Phi < 2^{n/2}$, we have

$$\text{Covr}_{n,\Phi}(\frac{1}{n}, \frac{3}{n}) \leq 2^{-N/4}.$$  

**Proof.** Let $\alpha = 1/16$ and $\beta = 3/n$. Our task is then to bound the probability that there exists a set of size $\alpha N$ that $n/2$-covers more than $\beta N$ lists.

Fix a set $S \subseteq [N]$ of size $\alpha N$. We compute the probability that $S$ $n/2$-covers more than $\beta N$ of the chosen lists.

The probability that a fixed set $S$ $n/2$-covers a list $L$ is at most $(\binom{n}{N}) \cdot (2\alpha)^{n/2}$ over the random choice of $L \sim [N]^n$. Using the inequality $(\frac{1}{k}) \leq (ne/k)^k$, we can bound this probability by $(2e\alpha)^{n/2}$.

Then, the probability that a fixed set $S$ $n/2$-covers some subset of $\beta N$ of the $\Phi$ lists is then

$$\left( \frac{\Phi}{\beta N} \right) \left( 2(2e\alpha)^{n/2} \right)^{\beta N} \leq \left( \frac{\Phi}{\beta N} \right) \left( 2(2e\alpha)^{n/2} \right)^{\beta N}$$

Finally we apply the union bound over all $(\binom{N}{\alpha N})$ possible choices of the set $S$ to get the final probability:

$$\left( \frac{N}{\alpha N} \right) \left( \frac{\Phi}{\beta N} \right) \left( 2(2e\alpha)^{n/2} \right)^{\beta N} \leq \left( \frac{\Phi}{\beta N} \right) \left( 2(2e\alpha)^{n/2} \right)^{\beta N}$$

and since $(e/\alpha)^{\alpha N} = (16e)^{1/16} < 2^{1/2}$ and $\beta = 3/n$,

$$\leq 2^{N/2} \left( \frac{\Phi e n}{3N} \right) \left( 2(2e\alpha)^{n/2} \right)^{\beta N}$$

and since we have assumed $N > en, en/N \leq 1$, so

$$\leq \left( 2^{1/2} \left( \Phi \cdot (2e\alpha)^{n/2} \right)^{\beta N}$$

and letting $\alpha = 1/16$ implies $2e\alpha = e/16 < 2^{-3/2}$, so

$$\leq 2^{N/2} \left( \Phi \cdot 2^{-3n/4} \right)^{3N/n}$$

and since $\Phi < 2^{n/2}$,

$$\leq 2^{N/2} \left( 2^{-n/4} \right)^{3N/n} \leq 2^{-N/4}.$$

\[ \square \]

### A.4 Our construction

**Location-hiding encryption scheme.** The construction is parameterized by:

- a universe size $N \in \mathbb{Z}^{>0}$,
- a cluster size $n \in \mathbb{Z}^{>0}$,
- a PIN-space $\mathcal{P}$,
- a recovery threshold $t \in \mathbb{Z}^{>0}$,
- a public-key encryption scheme $(\text{PKE}, \text{KeyGen}, \text{PKE}.\text{Encrypt}, \text{PKE}.\text{Decrypt})$, and
- an authenticated encryption scheme $(\text{AE}.\text{Encrypt}, \text{AE}.\text{Decrypt})$ with keyspace $\mathbb{F}$ (some finite field), and
- a security parameter $\lambda \in \mathbb{Z}^{>0}$.

The message space is $\{0,1\}^*$. We use a hash function $\text{Hash}: \{0,1\}^t \times \mathcal{P} \to [N]^n$ (e.g., built from SHA-256) which we model as a random oracle [9]. We use $t$-out-of-$n$ Shamir secret sharing [73]; we denote the sharing and reconstruction algorithms over a finite field $\mathbb{F}$ by $(\text{Shamir}.\text{Share}_\mathbb{F}, \text{Shamir}.\text{Reconst}_\mathbb{F})$.

$\text{KeyGen}(1^\lambda) \to (pk, sk)$.

- On input security parameter $\lambda$, expressed in unary, run the key-generation routine for the underlying public-key encryption scheme: $(pk, sk) \leftarrow \text{PKE}.\text{KeyGen}(1^\lambda)$.

$\text{Encrypt}(pk_1, \ldots, pk_N, \text{pin}, \text{msg}) \to (\text{salt}, ct)$.

- Sample a random salt $\text{salt} \sim \{0,1\}^t$ and compute $(i_1, \ldots, i_n) \leftarrow \text{Hash}(\text{salt}, \text{pin})$.

- Sample a transport key $k \sim \mathbb{F}$ and split it using $t$-out-of-$n$ Shamir secret sharing [73]:

  $$k_1, \ldots, k_n \leftarrow \text{Shamir}.\text{Share}_\mathbb{F}(k).$$

- Encrypt each share of the key using the underlying public-key encryption scheme. That is, for $j \in \{1, \ldots, n\}$, set $C_j \leftarrow \text{PKE}.\text{Encrypt}(pk_j, k_i)$.

- Set $M \leftarrow \text{AE}.\text{Encrypt}(k, \text{msg})$, $ct \leftarrow (M, C_1, \ldots, C_n)$, and output $(\text{salt}, ct)$.

$\text{Select}(\text{salt}, \text{pin}) \to (i_1, \ldots, i_n) \in [N]^n$.

- Output $\text{Hash}(\text{salt}, \text{pin}) \in [N]^n$.

$\text{Decrypt}(sk_i, ct) \to \sigma_i \in \mathbb{G}$.

- Parse $ct$ as $(M, C_1, \ldots, C_n)$.

- Output $\sigma \leftarrow (M, \text{PKE}.\text{Decrypt}(sk, C_i))$.

$\text{Reconstruct}(\sigma_1, \ldots, \sigma_n) \to \text{msg}$.

- For $i \in [n]$, Parse each share $\sigma_i$ as a pair $(M_i, k_i)$.

- Let $k \leftarrow \text{Shamir}.\text{Reconst}_\mathbb{F}(k_1, \ldots, k_n) \in \mathbb{F}$.

- Let $M$ be the most common value in $\{M_1, \ldots, M_n\}$.

- Output $\text{AE}.\text{Decrypt}(k, M)$.

Figure 15: Our construction of location-hiding encryption.

Our construction, which is parameterized by a standard
public-key encryption scheme, appears in Figure 15. We instantiate the public-key encryption scheme with hashed ElGamal encryption, which we recall here.

**Hashed ElGamal** We instantiate the location-hiding encryption scheme of Section 5 with the “Hashed ElGamal” encryption scheme. The scheme uses a cyclic group $G = \langle g \rangle$ of prime order $p$, in which we assume that the computational Diffie-Hellman problem is hard. It also uses an authenticated encryption scheme $(AE, Encrypt, AE, Decrypt)$ with key space $K$ and a hash function $\text{Hash} : G \rightarrow K$.

A keypair is a pair $(x, g^x) \in \mathbb{Z}_p \times G$, where $x \in \mathbb{Z}_p$. To encrypt a message $m \in \{0, 1\}^\ell$ to public key $X \in G$, the encryptor computes $(g^r, AE(\text{Hash}(X^r), m))$.

A standard argument [15] shows that Hashed ElGamal satisfies semantic security against chosen-ciphertext attacks [71, 22].

In our use of hashed ElGamal, we can provide domain separation between different ciphertexts by prepending inputs to the hash function $\text{Hash}$ during encryption and decryption with: (1) the client’s username, (2) the salt associated with the ciphertext, and (3) the public keys of the $n$ public keys to which the client encrypted the ciphertext. All of these values are available to the client during encryption and to the HSMs during decryption.

### A.5 Proof of correctness

We now prove:

**Theorem 9.** Consider an instantiation of the encryption scheme $E$ of Figure 15 with parameters $(N, n, P, t)$, with cluster size $n = \Omega(\lambda)$, on security parameter $\lambda$, and recovery threshold $t = n/2$. Then the resulting scheme is a correct location-hiding encryption scheme (in the sense of Definition 1) for any fault-tolerance parameter $f_{\text{live}} \leq \frac{1}{8}$.

**Proof:** To prove the claim, we need to bound the probability that, out of a random size-$n$ subset of all $N$ keys, no more than $t = n/2$ are failed. Fix a size-$(n/2)$ subset $S$ of the $n$ chosen keys. The probability that the keys in $S$ are all failed is $(f_{\text{live}})^{n/2}$. We can now take a union bound over all $\binom{n}{n/2}$ choices of $S$ to bound the final failure probability:

$$\Pr[\text{fail}] \leq \binom{n}{n/2} \cdot (f_{\text{live}})^{n/2} \leq 2^n \cdot \left(\frac{1}{2^t}\right)^{n/2} = 2^{-n/2}.$$  

Since we have taken $n = \Omega(\lambda)$, this probability is negligible in $\lambda$. \hfill \Box

### A.6 Proof of security

We state the main theorem and then prove it here.

**Theorem 10.** The location-hiding encryption scheme $E$ of Figure 15 instantiated with the hashed-ElGamal public-key encryption system (Appendix A.4) is secure, in the sense of Definition 3, in the random-oracle model, for recovery threshold $t = n/2$ and corruption threshold $f_{\text{secret}} = 1/16$.

More precisely, we consider an instantiation of our location-hiding encryption scheme $E$ with parameters $(N, n, P, t = n/2)$, where $N > \cdot n \approx 2.71n$ and $|P| < 2^{\ell/2}$, using hashed ElGamal encryption for the public-key encryption scheme, and using an arbitrary authenticated encryption scheme $AE$.

Then, let $A$ be an adversary that attack $E$ in the location-hiding encryptions security game with corruption threshold $f_{\text{secret}} = 1/16$. Denote $A$’s attack advantage as $L\text{HEncAdv}[A, E]$. Then if $A$ makes at most $Q$ queries to the random oracle $\text{Hash}'$, used in hashed ElGamal encryption, we construct:

- an efficient algorithm $B_{\text{CDH}}$ that breaks CDH in group $G$ with advantage $\text{CDHAdv}[B_{\text{CDH}}, G]$ and
- an efficient algorithm $B_{\text{AE}}$ that breaks the authenticated encryption scheme $AE$ with advantage $\text{AEAdv}[B_{\text{AE}}, AE]$ such that

$$L\text{HEncAdv}[A, E] \leq 2^{-N/4} + N \cdot Q \cdot \text{CDHAdv}[B_{\text{CDH}}, G] \cdot \frac{3N}{n|P|} + \text{AEAdv}[B_{\text{AE}}, AE].$$

Notice that if the number of HSMs $N$ grows much larger than the size of the PIN space $|P|$, the bound of Theorem 10 on the adversary’s advantage becomes vacuous. (In fact, this limitation is inherent—see Remark 5 in Appendix A.) To support extremely large data deployments, it would be possible to shard users into data centers of moderate size (e.g., $N \approx 50,000$).

The main technical challenge in proving Theorem 10 comes from the fact that the adversary may adaptively compromise a subset of the secret keys. This is very similar to the issues that arise when proving a cryptosystem secure against “selective opening attacks” [27, 8].

**Proof of Theorem 10 (Security).** We construct the following series of games and show that the difference between the adversary’s advantage from one game to the next is small. For $i \in \{0, \ldots, 4\}$, let $W_i$ the event that the adversary wins in Game $i$, where we define the winning condition for each game below.

**Game 0.** Game 0 proceeds according to the location-hiding security game where the challenger plays the role of the random oracle. For location-hiding encryption scheme and adversary $A$ in Game 0, by definition,

$$\Pr[W_0] = L\text{HEncAdv}[A, E].$$  (1)

**Game 1.** Game 1 proceeds in the same way as Game 0, except that for the adversary to win in Game 1, we also require that a certain “bad event” does not take place. This bad event is that many different PINs hash to the same set of HSMs. Formally,
**Game 0.** We define Game 0 by instantiating Experiment 4 our location-hiding encryption scheme, instantiated in turn with hashed-ElGamal encryption (Appendix A.4). Game 0 is parameterized by an adversary \( \mathcal{A} \), a group \( G \) of prime order \( p \) with generator \( g \), a universe size \( N \in \mathbb{Z}^{>0} \), a cluster size \( n \in \mathbb{Z}^{>0} \), a PIN-space \( \mathcal{P} \), a recovery threshold \( t \), a security parameter \( \lambda \), hash functions \( \text{Hash} : \{0,1\}^\lambda \times \mathcal{P} \to \{N\}^n \) and \( \text{Hash}' : \{0,1\}^\lambda \to \mathcal{F} \) (which we model as random oracles where the challenger plays the role of the random oracle), an authenticated encryption scheme \( \text{AE} \) (Encrypt, AE.Decrypt) with keyspace \( \mathcal{F} \) and message space \( \{0,1\}^* \), a security parameter \( \lambda \in \mathbb{Z}^{>0} \), and a bit \( \beta \in \{0,1\} \).

- The challenger runs:
  
  \[ \text{sk}_i \leftarrow \mathbb{Z}_p \quad \text{pk}_i \leftarrow g^{\text{sk}_i} \text{ for } i \in \{1,\ldots,N\} \]
  
  and sends \( (\text{pk}_1,\ldots,\text{pk}_N) \) to the adversary.
- The adversary chooses two messages \( \text{msg}_0,\text{msg}_1 \in \mathcal{M} \) and sends these to the challenger.
- The challenger then computes the ciphertext using hashed-ElGamal. That is, the challenger computes
  
  \[ \text{pk}_i \leftarrow \text{Hash}(\text{salt},\text{pin}) \]
  
  \[ k \leftarrow \mathcal{F} \]
  
  \[ (\text{sk}_1,\ldots,\text{sk}_n) \leftarrow \text{Shamir.Share}_G(k). \]

Here, Shamir.Share\(_G\) denotes \( t \)-out-of-\( n \) Shamir secret sharing over the field \( \mathcal{F} \). The challenger then encrypts the shares \( (\text{sk}_1,\ldots,\text{sk}_n) \) of the transport key \( k \) using hashed ElGamal encryption. For \( j \in \{1,\ldots,n\} \), the challenger computes:

\[ r_j \leftarrow \mathbb{Z}_p \]
\[ k_j \leftarrow \text{Hash}'((\text{pk}_j)^{r_j}) \]
\[ C_j \leftarrow (g^{r_j},k_j \oplus k_j). \]

Finally, the challenger encrypts the message \( \text{msg} \) with the transport key \( k \) and outputs the ciphertext:

\[ M \leftarrow \text{AE.Encrypt}(k,\text{msg}_\beta) \]
\[ \text{ct} \leftarrow (M,C_1,\ldots,C_n) \]

and sends \( (\text{salt},\text{ct}) \) to the adversary.
- The adversary may make adaptive \( f_{\text{secret}} \cdot N \) corruption queries and may perform computation between its queries. At each query:
  
  - the adversary sends to the challenger an index \( i \in [N] \)
  - the challenger sends to the adversary \( \text{sk}_i \).
- Finally, the adversary outputs a bit \( \beta' \in \{0,1\} \).

The output of the experiment is the bit \( \beta' \).

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Lemma 11. Let $W_1$ be the event that the adversary wins in Game 1 and $W_2$ be the event that the adversary wins in Game 2. In particular, for every adversary $A$ in Game 1, we construct a CDH adversary $B_{CDH}$ in group $G$ with advantage $CDHAdv[B_{CDH}, G]$ that runs in time linear in the runtime of $A$ and makes $Q$ Hash’-oracle queries such that
\[
|Pr[W_1] - Pr[W_2]| \leq N \cdot Q \cdot CDHAdv[B_{CDH}, G].
\]

Proof. Let $F$ be the event that the adversary makes a Hash’-oracle query at the point $pk_i^j$ before making a corruption query for key $i$. Then, by the definition of Games 1 and 2, and the Difference Lemma [75], we have
\[
|Pr[W_1] - Pr[W_2]| \leq Pr[F].
\]

So, to prove the lemma we need only bound $Pr[F]$.

Given an adversary $A$ in Game 1, we will construct a CDH adversary $B_{CDH}$ such that $Pr[F] \leq N \cdot Q \cdot CDHAdv[B_{CDH}, G]$. We construct $B_{CDH}$ as follows:

- The algorithm $B_{CDH}$ receives a challenge tuple $(g, g^x, g^y) \in G^3$ from the CDH challenger.
- The algorithm $B_{CDH}$ plays the role of the Game-1 challenger. The algorithm $B_{CDH}$ deviates from the normal behavior of the Game-1 challenger in two ways:
  1. The challenger chooses a random value $i^* \sim [N]$. For the $i^*$th keypair, algorithm $B_{CDH}$ sets $pk_i^* \leftarrow g^x$, where $g^x$ is the value from the CDH challenger. (Algorithm $B_{CDH}$ generates all other keypairs as the challenger in Game 1 does.)
  2. When constructing the final ciphertext, algorithm $B_{CDH}$ sets the encryption nonce in the $i^*$th ciphertext to be the value $g^x$, where $g^x$ is the value from the CDH challenger. For the value Hash’$(/[pk_i^{y^*}]^*)$ needed to construct the ciphertext, the algorithm $B_{CDH}$ just chooses a random element in $F$.
- The algorithm $B_{CDH}$ then proceeds in the same way as the Game 1 challenger with the following modification:
  - If $A$ makes a corruption query for $i^*$, $B_{CDH}$ outputs $\bot$.
  - If $B_{CDH}$ did not output $\bot$, then $B_{CDH}$ randomly chooses one of the points at which $A$ made a Hash’-oracle query and returns the queried point to the CDH challenger.

We now compute algorithm $B_{CDH}$’s CDH advantage. Whenever event $F$ occurs, the algorithm $A$ makes a random-oracle query to a point $(pk_i)^{y^*}$, for some $i \in [N]$, before issuing a corruption query at $i$. Notice that $B_{CDH}$ succeeds whenever:

1. Event $F$ occurs,
2. $i = i^*$, and
3. the algorithm $B_{CDH}$ guesses the correct random-oracle query to output.

These three events are independent. Furthermore, their probabilities are:

1. $Pr[F]$ – to be computed later,
2. $Pr[i = i^*] = 1/N$, and
3. $Pr[guesses\ correct\ r.o.\ query] = 1/Q$.

Therefore,
\[
CDHAdv[B_{CDH}, G] \geq Pr[F] \cdot \left(\frac{1}{N}\right) \cdot \left(\frac{1}{Q}\right),
\]
which proves the lemma.

\[\square\]

Lemma 12. Let $W_2$ be the event that the adversary wins in Game 2 and $W_3$ be the event that the adversary wins in Game 3. Then
\[
|Pr[W_2] - Pr[W_3]| \leq \frac{3N}{n|P|}.
\]

Proof. Let $F$ be the event that the adversary $A$ wins in Game 2 but not in Game 3. By construction of the games, we have
\[
|Pr[W_2] - Pr[W_3]| \leq Pr[F],
\]
so our task is to bound $Pr[F]$.

To analyze $Pr[F]$, we consider a modified game between $A$ and its challenger in Games 2 and 3. Here, we modify the challenger to halt the execution of $A$ as soon as event $F$ occurs. This modification cannot increase $Pr[F]$.

However, in the modified game, $A$’s view is independent of the uncorrupted elements of $I$ until event $F$ occurs. This is so because the only values that the adversary sees that depend on the set $I$ are the $k_i$ values. Since these are computed as the output of Hash’$(/[pk_i^{y^*}]^*$), until the adversary queries the random oracle Hash’ at the point $(pk_i)^{y^*}$, these $k_i$ values are also independent of $I$.

Since we have already argued (Game 2) that the challenger never queries the random oracle at a point $(pk_i)^{y^*}$ without having corrupted $i$, the ciphertext values that the adversary sees are independent of the uncorrupted elements of $I$.

Therefore, the answers to the corruption queries give the adversary no information on the uncorrupted elements of $I$. Then $Pr[F]$ is just the probability that the adversary makes $N/16$ non-adaptive corruption queries and is able to cause event $F$ to occur.

By the winning condition of Game 1, for any set of $N/16$ corrupted HSMs $S$, there are at most $3N/n$ PINs in $S$ such that $|Hash(salt, pin) \cap S| > n/2$. Therefore, the probability that $F$ occurs is $Pr[F] \leq (3N/n)|P| = 3N/(n|P|)$.

\[\square\]

Lemma 13. Let $W_3$ be the event that the adversary wins in Game 3 and $W_4$ be event that the adversary wins in Game 4. Then
\[
|Pr[W_3] - Pr[W_4]| = 0.
\]

Proof. We use the security of Shamir secret sharing to prove this lemma. For each of the indexes that the adversary does not corrupt, we can replace $k_i \oplus k_i'$ with a random element in $F$ in both games. We know that in both games, the adversary
corrupts fewer than \( n/2 \) of the keys in \( I \), and so the adversary learns fewer than \( n/2 \) of the shares \( k_1, \ldots, k_n \) of the transport key. Therefore, we can replace the transport key \( k \) with \( k' \in \mathbb{F} \), completing the proof.

\[ \Pr[A_4] = \text{AEAdv}[B_{AE}, \text{AE}] \ . \]

**Proof.** Given an adversary \( A \) in Game 4, we will construct an adversary \( B_{AE} \) attacking the authenticated encryption scheme \( AE \) such that the advantage of \( B_{AE} \) is identical to that of \( A \).

We construct \( B_{AE} \) as follows:

- The algorithm \( B_{AE} \) computes \( (pk_1, \ldots, pk_N) \) in the same way as the Game 4 challenger and sends them to \( A \).
- When \( A \) returns the messages \( m_{sg0}, m_{sg1} \in M \), \( B_{AE} \) forwards the messages to the AE challenger and receives the ciphertext \( c^* \) from the AE challenger.
- The algorithm \( B_{AE} \) then computes the ciphertext in the same way as the Game 4 challenger except that instead of computing the ciphertext as \((AE, \text{Encrypt}(k, \text{msg}), C_1, \ldots, C_n)) \), \( B_{AE} \) sends \( A \) \((c^*, C_1, \ldots, C_n)\).
- The algorithm \( B_{AE} \) responds to queries in the same way as the Game 4 challenger.
- When \( A \) outputs bit \( \beta' \in \{0, 1\} \), \( B_{AE} \) forwards the response to the AE challenger.

Because in Game 4, we sample \( k' \) independently from the rest of the messages, \( A \) cannot distinguish between interaction with \( B_{AE} \) and the Game 4 challenger. The advantage of \( A \) is exactly \( \text{AEAdv}[B_{AE}, \text{AE}] \), completing the proof. \( \Box \)

**B.2 Implementation using Merkle trees**

We now describe how to implement the routines defined in Section 6.1, using a construction as in Nissim and Naor [67]. In the following discussion, we define the “log tree” for a log \( L \) to be a binary search tree, ordered by identifiers \( id \). Each internal node in the tree—in addition to containing a value—contains the cryptographic hash of its two child nodes, as in a Merkle tree.

\( \text{Digest}(L) \rightarrow d \). Construct the log tree for \( L \) by inserting the elements of \( L \) into a binary-search tree one at a time. (We can use any type of self-balancing binary-search tree here.)

As the digest \( d \), output the hash of the root of the log tree.

\( \text{ProveIncl}(L, \text{id}, \text{val}) \rightarrow \{\text{inc}, \perp\} \). If \((\text{id}, \text{val}) \notin L \), output \( \perp \). Otherwise, output the Merkle inclusion proof that proves that \((\text{id}, \text{val}) \) is in the log tree rooted at \( d = \text{Digest}(L) \).

\( \text{ProveExt}(L, L') \rightarrow \{\text{ext}, \perp\} \). We show how the routine works in the special case in which \( L' \) contains exactly one entry \((\text{id}, \text{val})\) that does not appear in \( L \). To generalize to the case in which there are many new entries in \( L' \), we run this routine once for each new entry and output the concatenation of all of the resulting proofs.

If \( L' \) does not extend \( L \), output \( \perp \). Otherwise, find the identifiers \( \text{id}_{\text{left}} \) and \( \text{id}_{\text{right}} \) that appear just before and after \( id \) in the lexicographical ordering of identifiers in the old log \( L \). Let their corresponding values be \( \text{val}_{\text{left}} \) and \( \text{val}_{\text{right}} \). The first portion of the proof \( \text{ext} \) is a Merkle proof of inclusion of \((\text{id}_{\text{left}}, \text{val}_{\text{left}})\) and \((\text{id}_{\text{right}}, \text{val}_{\text{right}})\) in the old log tree rooted at digest \( d = \text{Digest}(L) \). These proofs prove that the new identifier \( id \) is not in the log represented by the old digest \( d \).

Next, insert \((\text{id}, \text{val})\) into the log tree for \( L \) to get a log tree for \( L' \) and its corresponding digest. The second portion of the proof \( \text{ext} \) is a Merkle proof of inclusion of every node in the log tree for \( L' \) that does not appear in the log tree for \( L \). These proofs prove that the new digest \( d' \) represents the root of \( L' \)’s log tree, with the new pair \((\text{id}, \text{val})\) inserted.

\( \text{ProveExt}(d, d', \text{ext}) \rightarrow \{0, 1\} \). Parse \( \text{ext} \) as a series of Merkle inclusion proofs over log trees. Check that \( id \) lies between \( \text{id}_{\text{left}} \) and \( \text{id}_{\text{right}} \) in lexicographic order. Verify each of the Merkle proofs generated in \( \text{ProveExt} \).

Checking the Merkle proofs in this case also requires
verifying that the nodes in log trees satisfy the proper ordering constraints of a binary search tree: the left child’s value is less than the parent’s value and the right child’s value is greater than the parent’s value. Accept if all of these proofs accept.

**Security properties.** The completeness properties are immediate. Inclusion soundness follows directly from the analysis of Merkle proofs, which in turn rely on the collision-resistance of the underlying hash function. To sketch the argument for extension soundness: The first part of the proof $\pi_{\text{Ext}}$ convinces the verifier that the new identifier id does not appear in the log $L$ represented by the old digest $d$. Therefore, the log $L'$ that digest $d'$ represents must not contain any duplicate identifiers, since $L$ contains no duplicate identifiers. The second part of the proof $\pi_{\text{Ext}}$ convinces the verifier that log tree rooted at $d'$ is identical to the log tree rooted at $d$, except with the pair $(\text{id}, \text{val})$ added.

### B.3 Making progress in spite of failures

If frequent node failures prevent the log from making progress, the HSMs can run the following more complicated log-update protocol. In this variant, each HSM chooses which log chunks to audit as a deterministic function of the Merkle root $R$ and the HSM’s own node ID.

Provided that we choose the constant $C$ large enough, for every choice of the Merkle root, $R$ at least one honest HSM will audit each log chunk. In particular, by taking $C \geq 384$ the probability that no honest HSM audits a particular chunk is less than $2^{-384}$. The probability that there exists a 256-bit Merkle root $R$ that causes no honest HSM to audit a particular chunk is then at most $2^{236} \cdot 2^{-384} = 2^{-128}$. So, no matter how the service provider influences the Merkle root $R$, at least one honest HSM will attempt to audit each chunk.

Using this method, given the Merkle root $R$, every HSM can deterministically compute the set of log chunks that every other HSM will audit. Then, if any HSM fails during the audit process, the remaining non-failed HSMs can recursively run our randomized-checking protocol to check the log chunks that the failed HSMs would have checked.

In this way, the protocol can make progress even if HSMs fail during the log-update process.

### C Details on outsourced storage with secure deletion

We model the external service provider as an oracle $S$ to which the HSM has access. To read and write blocks from external storage at address $addr \in \mathbb{Z}^+$ the HSM can call $S.\text{Get}(addr) \rightarrow \text{block}$ or $S.\text{Put}(addr, \text{block})$.

The HSM needs to implement the following functions:

- **Setup** $(\text{data}, \text{addr}) \rightarrow \text{sk}$. Store the given data blocks in the external storage system $S$ encrypted under a secret key sk. The HSM stores only sk in its internal memory.
- **Read** $(\text{sk}, i) \rightarrow \text{data}_i$ or $\perp$. Use key sk to read the $i$th logical data block from the storage system $S$. Return $\perp$ on failure.
- **Delete** $(\text{sk}, i) \rightarrow \text{sk}'$ or $\perp$. Use key sk to delete the $i$th logical data block from the external storage system $S$. Return a new key sk’ to store in internal memory. Return $\perp$ on failure.

Let $(\text{AE.Encrypt}, \text{AE.Decrypt})$ be a symmetric-key authenticated encryption scheme such as AES-GCM, with key space $K$. We store data blocks encrypted as leaves of a binary tree of height $h = 1 + \lceil \log_2(D) \rceil$. We store leaf $i$ at address $2^h + i$, block $a$’s parent at address $\lfloor a/2 \rfloor$, and $a$’s left and right children at $2a$ and $2a + 1$, respectively.

For convenience, we have optional parameters that take a non-default value only on recursive calls. We bundle outsourced blocks into a 0-indexed vector that can be sliced. We assume slicing beyond the end of an array yields a zero-length slice, and assigning $\text{sk} / \text{msg} \leftarrow \text{msg}$, if $\text{msg}$ is the length of only one key, sets $\text{sk} \leftarrow \text{msg}$ and $\text{sk} \leftarrow \text{empty}$. We assume if AE.Decrypt fails, it raises an exception caught outside of these routines.

**Setup** $(\text{data}, \text{addr} = 1) \rightarrow \text{sk}$.

- Let $D \leftarrow \text{len(data)}$
- If $D = 0$, set $\text{msg} \leftarrow 0 \cdots 0$
- Else if $D = 1$, set $\text{msg} \leftarrow \text{data}[0]$
- Else ($D > 1$)
  - Run $\text{sk}_r \leftarrow \text{Setup}^S(\text{data}[0: \lfloor D/2 \rfloor], 2 \cdot \text{addr})$
  - Run $\text{sk}_l \leftarrow \text{Setup}^S(\text{data}[\lfloor D/2 \rfloor : n], 2 \cdot \text{addr} + 1)$
  - Let $\text{msg} \leftarrow \text{sk}_r$
- Let $\text{sk} \leftarrow \text{sk}_l$
- $S.\text{Put}(\text{addr}, \text{AE.Encrypt} (\text{sk}, \text{msg}))$
- Return $\text{sk}$

**Read** $(\text{sk}, i, \text{level} = h - 1) \rightarrow \text{msg}$.

- Let $\text{addr} \leftarrow \lfloor (2^h + i)/2^\text{level} \rfloor$
- Let $c \leftarrow S.\text{Get} (\text{addr})$
- If $\text{level} = 0$ then return $\text{AE.Decrypt} (\text{sk}, c)$
- If $i \& 2^\text{level} - 1 = 0$ (i.e., $i$ in left child), let $\text{sk} \leftarrow \text{AE.Decrypt} (\text{sk}, c)$
- Otherwise (right child), let $\text{sk} \leftarrow \text{AE.Decrypt} (\text{sk}, c)$
- Return $\text{Read}^S (\text{sk}, i, \text{level} - 1)$

**Delete** $(\text{sk}, i, \text{level} = h - 1) \rightarrow \text{sk}$.

- If $\text{level} = 0$ return $0 \cdots 0$ (useless encryption key)
- Let $\text{addr} \leftarrow \lfloor (2^h + i)/2^\text{level} \rfloor$
- Let $\text{sk}_r \leftarrow \text{AE.Decrypt} (\text{sk}, S.\text{Get} (\text{addr}))$
- If $i \& 2^\text{level} - 1 = 0$ (i.e., $i$ in left child)
  - Let $\text{sk}_l \leftarrow \text{Delete}^S (\text{sk}, i, \text{level} - 1)$
- Else (right child) let $\text{sk} \leftarrow \text{Delete}^S (\text{sk}, i, \text{level} - 1)$
- Let $\text{sk} \leftarrow \text{sk}$.
D Artifact Appendix

D.1 Abstract

The SafetyPin implementation is split into two components:
- **HSM**: The HSMs (hardware security modules) are used to recover user secrets. Our implementation uses SoloKeys, which are low-cost HSMs. We add roughly 2,500 lines of C code to the open-source SoloKey firmware.
- **Host**: The host implements functionality for the user and data center, including saving secrets, maintaining the log, and coordinating HSMs. Our implementation is roughly 3,800 lines of C/C++ code.

We implement the protocol described in the paper above. To improve performance, we rewrote parts of the SoloKey firmware to use USB CDC, a high-throughput USB class commonly used for networking devices. This results in roughly a 32x increase in I/O throughput. Our artifact is available at:

https://github.com/edauterman/SafetyPin

D.2 Artifact check-list

- **Hardware**:
  - ★ 100 SoloKeys
  - ★ 10 Anker SuperSpeed USB 3.0 hubs
  - ★ 2 4-port USB PCIe controller cards
  - ★ Linux machine with Intel Xeon E5-260 CPU clocked at 2.60GHz

- **Compilation**: The ARM compiler for the SoloKeys, and gcc for the host.
- **Metrics**: Latency
- **Experiments**: Log-audit time, puncturable encryption overhead, breakdown of recovery time, cluster size vs. recovery time
- **Required disk space**: 14MB
- **Expected experiment run time**: 50 minutes
- **Public link**: https://github.com/edauterman/SafetyPin
- **Code licenses**: Apache v2

D.3 Description

D.3.1 How to access

We provide reviewers with credentials to remotely access our system. Instructions for assembling a similar system are available here:

https://github.com/edauterman/SafetyPin/blob/master/SETUP.md

D.3.2 Hardware dependencies

Our artifact uses SoloKeys as low-cost HSMs. We use 100 SoloKeys for our experiments, although other deployments could use a different number of HSMs. Because of limitations in the Intel XCHI controller, which supports a maximum of 96 endpoints (each USB 3.0 device has 3 endpoints), we installed PCIe cards to support additional endpoints. This is not necessary for smaller-scale deployments, but larger deployments should choose a host that supports installing such PCIe cards or find another solution. We also recommend USB hubs with an external power source such as the Anker hubs.

D.3.3 Software dependencies

The firmware for the SoloKeys builds on the original SoloKey firmware, which already includes several libraries for cryptographic primitives on embedded devices:

- https://github.com/solokeys/solo
- To support pairings for aggregate signatures, we use the jedi-pairing library for embedded devices:
  - https://github.com/ucbrise/jedi-pairing/
- For USB HID support, we use the Signal11 library:
  - https://github.com/signal11/hidapi

We implement our cryptographic primitives that do not require pairings at the host using OpenSSL.

D.4 Installation

Instructions for building the host are available under host/. Instructions for building the firmware and flashing the SoloKeys are available under hsm/. The SoloKey documentation provides additional details and troubleshooting for building and flashing the SoloKeys:

https://docs.solokeys.io/

When experimenting with SafetyPin, you should not boot SoloKeys in DFU mode, as this locks the firmware and will prevent you from modifying the firmware later (e.g. to load an updated version of the SafetyPin source).

D.5 Experiment workflow

Reviewers can remotely access our machine and run all experiments by executing ./runAll.sh in bench/. More detailed instructions for running individual experiments are available here:

https://github.com/edauterman/SafetyPin#instructions-for-artifact-evaluation.
D.6 Evaluation and expected result

Run all experiments by executing `.runAll.sh` in `bench/`. This will produce figures in `bench/out` that match Figure 8, Figure 9, Figure 10, and Figure 11. Note that we only reproduce the recovery time breakdown in Figure 10. Additionally, the configuration we set up for the reviewers only uses 90 HSMs for Figure 8 and Figure 11. We do this to keep different firmware on the remaining 10 HSMs to measure the breakdown in puncturable encryption time as the secret key size increases (Figure 9). For the experiments we show in the body of the paper, we re-flashed HSMs between experiments so that we could use all 100 HSMs to generate Figure 8 and Figure 11.

D.7 Experiment customization

The experiment for Figure 8 can be modified to measure different data center sizes without changing the firmware on the HSMs. The experiment for Figure 9 can likewise be modified to measure different secret key sizes, although this requires changing HSM firmware. If we had more HSMs, we could easily expand Figure 11 to show the effect of larger cluster sizes. We do not measure cluster sizes less than 40 because our analysis shows that our security guarantees begin to break down below this point.

D.8 Notes

To switch between USB CDC and USB HID, change the HID flag on both the host and the HSMs (this requires loading new firmware on the HSMs). More detailed instructions are available here:

https://github.com/edauterman/SafetyPin/blob/master/SETUP.md.

Note that rather than generating puncturable encryption secret keys on the HSM (a process we estimate would take roughly 75 hours), to run our experiments efficiently, we generate the secret key on the host (for security, a real-world deployment would need to generate this secret key on the HSM).

D.9 AE Methodology

Submission, reviewing and badging methodology:

https://www.usenix.org/conference/osdi20/call-for-artifacts