CTng: Secure Certificate and Revocation Transparency

In God we Trust; Others we Monitor

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***Abstract*—We present CTng, an evolutionary and practical PKI design that efficiently addresses multiple key challenges faced by deployed PKI systems. CTng ensures strong security properties, including *guaranteed transparency* of certificates and *guaranteed, unequivocal revocation*, achieved under *NTTP-security*, i.e., without requiring trust in any single CA, logger, or relying party. These guarantees hold even in the presence of arbitrary corruptions of these entities, assuming only a known bound (f) of corrupt monitors (e.g., $f = 8$), with minimal performance impact. CTng also enables *efficient certificate validation* and preserves *relying-party privacy*, while providing scalable and efficient distribution of revocation updates.**

These properties significantly improve upon current PKI designs. In particular, while Certificate Transparency (CT) [35], [36], [37] aims to eliminate single points of trust, the existing specification [36] still assumes benign loggers. Addressing this through log redundancy is possible, but rather inefficient, limiting deployed configurations to $f \leq 2$.

We present a security analysis and an evaluation of our open-source CTng prototype, showing that it is efficient and scalable under realistic deployment conditions.

I. Introduction

The *Public Key Infrastructure (PKI)* facilitates the secure use of public keys. PKI is critical for the security of open, distributed systems such as the Internet. Typically, a *relying party* obtains a public key and validates it using a *certificate* signed by a trusted *Certificate Authority (CA)*. The PKI defines how certificates are issued and revoked (by the CAs) and validated (by relying parties).

Most deployed PKIs follow the X.509 standard [8], [24]. X.509 certificates are used in protocols such as TLS, SSH, S/MIME, IPsec, and others. The most common application of PKI is to secure web and other forms of communication over the Internet using the TLS protocol. In particular, web communication is typically protected using HTTPS, which runs HTTP over TLS, with the browser acting as the relying party and validating the server's certificate. We refer to the PKI used to secure web communication as *Web-PKI*. In Web-PKI, relying parties (browsers) inherently trust a broad set of *root CAs*. Each root CA can issue certificates for any domain or CA, effectively acting as a *Trusted Third Party (TTP)*, either directly or indirectly, by certifying another CA and facilitating a certificate chain.

There have been multiple *PKI failures* [48], [45], [10], [4]. Typically, an attacker obtains a *rogue certificate*, i.e., a certificate that appears valid to relying parties, contains a public key corresponding to a private key known to the attacker, and includes the identifier (e.g., a domain) of a benign victim entity. The attacker then exploits the rogue certificate to impersonate the victim, typically as a trusted website.

These failures and attacks motivated numerous proposals and efforts to improve the security of PKI schemes, including [46], [16], [41], [29], [54], [5], [62], [52], [53], [40], [19], [56], [31], [38], [37], [18], [39], [27], [13]. Among these, *Certificate Transparency (CT)* [35], [36], [37] stands out as the only 'post-X.509' PKI scheme that has been deployed and used in practice. The main goal of CT is to make the set of issued certificates publicly visible (transparent), enabling detection of rogue certificates, e.g., allowing domain owners to discover unauthorized certificates issued for their domain. In principle, this could be achieved by requiring CAs to publish every certificate they issue. However, a rogue CA could simply choose not to publish certain certificates.

CT addresses this by introducing *public logs* operated by entities known as *loggers*. In CT, a certificate is considered valid only if it comes with a *Signed Certificate Timestamp (SCT)*, a signed commitment from a logger to include the certificate in its public log. This approach ensures transparency even in the presence of misbehaving CAs. The broader objective of CT was to eliminate reliance on any single trusted party, a principle termed the *No Trusted Third Party (NTTP)* goal [37]. However, CT, as standardized by the IETF in CTv1 [35] and CTv2 [36], ensures transparency only under the assumption

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that all certificates are logged by an honest logger [58]. In other words, neither version satisfies the NTTP goal. If a CA can be compromised or act maliciously, so can a logger.

CT has been adopted by major browsers, including Chrome, Safari, and Brave, all of which require that valid certificates include SCTs. Currently, there are six deployed CT loggers operated by organizations such as Google, Cloudflare, and Let's Encrypt. In addition, 13 organizations run monitors that help detect maliciously or mistakenly issued certificates. Most CAs now issue CT-compliant certificates.

In practice, browsers do attempt to ensure security against rogue loggers, but in a constrained and inefficient way, requiring each certificate to be logged with multiple loggers, as recommended in CTv2 [36]. Typically, due to overhead, only two[1] loggers are required. While they are chosen from a set approved by browser vendors, the selection is made by the (potentially corrupt) CAs, who decide whether to log certificates with a specific log, limiting the security benefits. Furthermore, although CT logs are append-only and verifiable by CT monitors, malicious loggers can still maintain and present separate logs (i.e., Merkle trees), making CT vulnerable to *split-world attacks*. Currently, CT does not employ any mechanism to verify that a logger does not equivocate [36].

Another concern is that, while CT has significantly improved transparency of certificates, CT does not address the critical area of *certificate revocation*. Standardized revocation mechanisms, such as CRLs and OCSP, have been largely abandoned by browsers due to performance and privacy concerns. Today, browsers rely primarily on proprietary mechanisms, such as Google Chrome's CRLSets, which typically cover only a small subset of certificates (see §VII).

Changing the status quo is never easy, particularly when it involves the Web-PKI, a large-scale system under the control of various stakeholders. Some might argue that the current state of Web-PKI is "good enough", while others might quickly point out the significant practical challenges involved in implementing even minor changes. Both criticisms have validity; however, settling for "good enough" is not a viable option when it comes to the security of such a crucial component of our online infrastructure, especially since it is plausible to transition to a secure, yet performance-oriented PKI. Recent regulatory developments, such as the EU's eIDAS 2 [12], which requires browsers to trust government-approved CAs and limits their ability to remove unsafe or malicious ones, underscore that the need for a stronger PKI is not merely a theoretical concern, but a practical necessity.

In response, we present CTng, a Web-PKI design and prototype system. CTng improves *security* by achieving NTTP-secure transparency and revocation, while also supporting *efficient certificate validation* and ensuring *privacy* for relying parties. It allows relying parties to prefetch certificate validation data, removing reliance on real-time checks. CTng also improves *efficiency* and *scalability*; in particular, clients can validate certificates and their revocation status efficiently as part of their connection to a server, with minimal bandwidth overhead and without needing to communicate with any other entities. Finally, the CTng design is *evolutionary*, preserving most aspects of the existing PKI, including CT.

[1]The exact number depends on factors like issuance date and validity period. It also varies between browsers.

CTng expands the role of CT monitors, empowering them to monitor the logs and efficiently provide the information required for relying parties to validate certificates. CTng achieves NTTP-security *efficiently*: a CA needs only to log certificates with a single logger, and a relying party needs only one low-bandwidth interaction with a single monitor to receive the periodic transparency and revocation updates. This approach is efficiently implemented using well-established and widely available *threshold signatures*, which have open-source implementations, as essential for successful deployment.

CTng benefits are especially evident in the context of revocation, where CTng ensures *NTTP-secure guaranteed and unequivocal revocation*, allowing monitors to provide timely updates (e.g., daily or even hourly) to relying parties. By prefetching this information, relying parties can validate certificates without depending on additional real-time checks, avoiding the costly over-provisioning required to handle traffic spikes. In contrast, current revocation approaches require CAs or browser vendors to serve requests from arbitrary clients and CT leaves loggers similarly exposed to peak load conditions. To further improve efficiency, CTng incorporates the compact CRV design from [50] to minimize the size of revocation information distributed to clients.

We present an open-source implementation [25] of two versions of CTng: a base version (§IV) and an optimized version with two optional design enhancements that reduce bandwidth (§IV-D1 and §IV-D2), allowing CTng to support more monitors and to be resilient to a larger number of faulty monitors.

Our evaluations (§VI) confirm the practicality and efficiency of CTng across all entities. On modest hardware, the base version of CTng with just the broadcast optimization (§IV-D1) exceeds current global-scale precertificate throughput and supports 32 monitors while tolerating up to 8 faulty monitors ($f = 8$). With the additional erasure encoding optimization (§IV-D2), performance further improves for higher number of faulty monitors (f) . Increasing the number of monitors has negligible impact on system throughput.

CONTRIBUTIONS:

- We present CTng, an evolutionary extension of the current Web-PKI based on PKIX [6] and CT, utilizing well-established cryptographic primitives and approaches.
- CTng efficiently achieves NTTP-secure transparency and revocation. In particular, it prevents *logger omission attacks*, where a logger fails to include a certificate in the log within the promised timeframe, and provides a defense against *split-world attacks*, where a logger might present different log views to different clients. Further, CTng ensures guaranteed and unequivocal revocation and consequently prevents the *Zombie certificate attack*, where a certificate may appear as non-revoked to some

relying parties after its revocation.

- CTng offers benefits such as efficient certificate validation and privacy for relying parties. In particular, relying parties do not need to perform real-time signature validations. CTng is efficient even when using high-overhead signature schemes, such as post-quantum signature schemes.
- CTng achieves its security goals. We present a security analysis that demonstrates its NTTP-security guarantees for both transparency and revocation.
- CTng is efficient and scalable. We present a performance evaluation of our open-source implementation, showing its practicality for realistic deployments.

II. From CT to CTng

We now provide an overview of the current Web-PKI, highlighting how CT addresses some of the limitations of the trusted CA model. We then examine the attacks that remain possible despite CT and discuss how they are mitigated by CTng. Table I summarizes these attacks and the corresponding defenses in CT and CTng, while Figure 1 offers a high-level comparison of the two systems.

A. Current Web-PKI

CT was introduced to mitigate the risk of misbehavior by CAs by introducing two new entities: loggers, responsible for ensuring the transparency of certificates, and monitors, responsible for auditing these logs to ensure correctness, detect problematic certificates and help identify rogue or negligent CAs. Adding these entities has changed the way certificates are issued and used.

To establish a secure communication channel between a relying party (i.e., a browser) and a subject (i.e., a server) s, the subject provides a PKIX certificate to the relying party, which certifies the public key pk_s of s. To obtain the certificate, the subject contacts a CA, which first verifies that the subject controls s[2], and then generates a *precertificate*, essentially a PKIX-compliant certificate that includes a dedicated 'poison' extension, which prevents it from being treated as a valid certificate by relying parties, as specified by CTv2 [36].

The CA then submits the precertificate to multiple loggers. Each logger verifies that the precertificate is PKIX-compliant and has not been previously logged. If so, it issues a Signed Certificate Timestamp (SCT), a promise to include the precertificate in its log, implemented using a Merkle tree to ensure auditability, within the *Maximum Merge Delay (MMD)*.

The CA aggregates the SCTs and, using the X.509v3 extensions mechanism [24], embeds them into the final certificate, which is then sent back to the subject. Relying parties accept a certificate only if it includes a sufficient number of valid SCTs issued by loggers they trust. The certificate issuance process is illustrated in Figure 1a, steps I.1–I.4.

Periodically, monitors retrieve the newly logged certificates and the current *Signed Tree Head* (STH) from loggers and ensure that: (1) the log is append-only, i.e., all past certificates are still in the log; and (2) the log is transparent, i.e., all newly logged certificates *that were reported* by the logger were added to the log, and *only* them. Monitors also analyze the newly added certificates to detect any possible mis-issuance, impersonation, or phishing attempts, and may notify affected subjects. The monitoring process is depicted in steps P.1–P.2.

To revoke a certificate, the subject requests revocation from the issuing CA. Once revoked, the CA can publish the revocation status via Certificate Revocation Lists (CRLs) or provide it via the Online Certificate Status Protocol (OCSP). Browser vendors periodically retrieve CRLs from trusted CAs and propagate (select) revocation information to browsers. This allows relying parties to reject revoked certificates, provided the relevant revocation data has been supplied to them. CT does not play a role in the revocation process. The revocation process is illustrated in Figure 1a, step R.1, and the vendor-assisted propagation of revocation is shown in steps U.1–U.3.

B. Remaining Web-PKI Attacks and CTng Defenses

We now discuss several attacks that adversaries can carry out in the current Web-PKI, focusing on threats posed by different entities within the ecosystem, and compare how these attacks are addressed by deployed and proposed defense mechanisms in CT with the defenses of CTng (Table I).

Misbehaving Subjects (Websites): A rogue website can launch a *stealthy corrupt certificate* attack by obtaining a valid but fraudulent certificate, either by deceiving a benign CA or colluding with a malicious CA. Without CT, this attack could remain undetected indefinitely. With CT, the attack window is limited, as the certificate must be publicly logged within the MMD period and then can be reported by monitors. However, the effectiveness of CT relies on active monitoring and a timely response by domain owners. CTng strengthens this defense via its *efficient validation* process (§IV-E), which enables monitors to distribute *verified* information to relying parties ahead of time. This removes the need for real-time checks and allows relying parties to independently validate certificates.

Misbehaving CAs: Although CT provides transparency for certificate issuance, it does not extend it to certificate revocation. This enables two distinct split-world attacks. In a *stealthy revocation DoS* attack, a malicious CA can falsely but selectively indicate that a valid, non-revoked certificate has *been revoked*, causing denial-of-service for the targeted website since the certificate would be rejected. In a *Zombie certificate* attack, a malicious CA can falsely indicate that a revoked certificate has *not been revoked*, enabling attackers who control the corresponding private key (e.g., from a past compromise) to impersonate legitimate domains. CT offers no built-in defense against either of these attacks, and existing revocation mechanisms (CRLs, OCSP) are insufficient due to CA control over the information these mechanisms rely on, and the lack of transparency in how proprietary versions of CRLs are implemented across different browsers. CTng mitigates both attacks through its *Periodic Consistent Broadcast*

[2]CTng, like CT and other PKI schemes, does not mandate how this validation must be performed.

protocol (§IV-D), which ensures that revocation information is distributed transparently and consistently to relying parties.

Misbehaving Loggers: Loggers in CT are assumed to act honestly, but this assumption can be violated in two main ways. In a *logger omission* attack, the logger can issue an SCT but never include the corresponding certificate into the log within the MMD. This prevents the monitors from discovering the certificate. CT attempts to address this through SCT auditing mechanisms [51], [26], which are optional and raise privacy concerns - although Google Chrome audits a small proportion of SCTs using k-anonymous queries [15]. In contrast, CTng's *efficient certificate validation* allows clients to confirm that the certificate is indeed logged without relying on SCT audits or any other real-time queries and without sacrificing privacy. In a *logger split-world* attack, a rogue logger can present inconsistent Merkle tree views of its log to different clients (or at different times), enabling selective suppression of certificates. Although gossip protocols [44], [43] have been proposed for CT, they have not been standardized or deployed. CTng directly addresses this through its Periodic Consistent Broadcast, ensuring all monitors receive and verify consistent views of all logs.

Misbehaving Monitors: In CT as deployed, monitors operate independently, checking logs for suspicious certificates but playing no active role in verifying logger behavior, particularly for split-world attacks. A misbehaving monitor may fail to report suspicious certificates or collude with attackers to ignore them, an issue that CT does not address. In contrast, CTng assigns monitors an active, collaborative role: they collectively ensure log integrity and distribute verified certificate information to relying parties. CTng assumes that some monitors may be arbitrarily malicious and adopts a quorum-based model (§III-A) that tolerates up to f faulty monitors. A relying party needs to contact only one monitor to obtain information; a misbehaving monitor cannot provide incorrect information, it can at most fail to respond. In that case, the relying party can try another monitor (see §IV-E) and is guaranteed to succeed after at most f attempts.

III. CTng: Model and Goals

We now describe the CTng system and adversary models, and present CTng's security, privacy, and system goals.

A. System and Adversary Models

In CTng, we assume the same five types of entities as in CT: CAs, loggers, monitors, subjects (websites), and relying parties (browsers). Each relying party has a set of root (anchor) CAs and a set of trusted loggers, along with their known public keys (or certificates). In practice, browser vendors define these trusted entities.

We assume a computationally-bounded adversary that controls any number of subjects, CAs, loggers, and relying parties, but only up to f monitors. The adversary gains full control of these entities, including their private keys. This attack model allows the adversary to perform any of the attacks listed in Table I. We assume a majority of benign monitors; that is, there must be at least $2f + 1$ monitors and their connections should be $(f+1)$-connected; in particular, each benign monitor should be connected to at least one other benign monitor.

CTng's monitors use threshold signatures; for simplicity, we assume that the distributed generation of the threshold key is completed before CTng begins running and that the relying parties know the corresponding threshold verification key.

We assume loosely-synchronized clocks where the drift between any entity's local clock C and the real time τ is bounded by Δ_{clk} at any time. We also assume that communication between pairs of *benign* entities is bounded by Δ_{com}. Together, these assumptions ensure that a benign relying party can reliably contact at least one benign monitor within the expected time bounds. For simplicity, we assume that each logger and CA is *monitored* by all monitors.

B. Security and Privacy Goals

In [58], Wrótniak et al. formally defined four PKI security requirements, and analyzed whether these requirements are satisfied by PKIX and CT under specific assumptions. We set these four requirements as security goals for CTng, and include intuitive definitions of these requirements below. For the formal definitions, see [58].

- **G1:** *Existential unforgeability*, i.e., every valid certificate ψ was either issued by the entity designated as the issuer of the certificate ($\psi.issuer$), or was issued by a (rogue) entity that managed to obtain a valid (but fraudulent) certificate $\hat{\psi}$ to a key it controls where $\hat{\psi}.subject = \psi.issuer$.
- **G2:** *Accountability*, i.e., every valid certificate has a root CA that is accountable for it (identified unequivocally as responsible for that certificate).
- **G3:** *Guaranteed transparency*, i.e., every certificate ψ that was logged at a logger ℓ at time t, every benign monitor that monitors ℓ (prior to t) is aware of ψ no later than $t+\Delta$, where Δ reflects the maximum delay allowed.
- **G4:** *Guaranteed revocation*, i.e., every certificate ψ that was revoked at time *t by its benign issuer*, will not be considered valid at any time after $t+\Delta$, where Δ reflects the maximum delay allowed for the revocation to be known.

That said, CTng not only aims to satisfy goals **G1-G4**, it also aims to do so under strict model assumptions. For example, as shown in Table II, both CTv1 and CTv2 satisfy the guaranteed transparency goal, but only under a weaker model assumption that loggers are benign (addressed through logger redundancy). In contrast, CTng assumes a stronger adversary model in which the adversary can control any number of loggers[3] and up to f monitors.

Because CTng targets a stronger adversary model, we identified an additional important property not defined in [58], and we define this property as an explicit goal for CTng:

- **G5:** *Unequivocal revocation*, i.e., an attacker cannot cause some relying parties to consider the certificate as

[3]For liveness, there needs to be at least one benign logger.

Attack model	Attack (§II-B)	CT defense	Current defenses
Subject, CA	Stealthy corrupt certificate	SCT validation	Efficient certificate validation (§IV-E)
	Stealthy revocation DoS	Short-lived certificates	Periodic Consistent Broadcast (§IV-D)
	Zombie certificate		
Subject, logger(s)	Logger omission	SCT auditing [51], [26] (privacy concerns)	Efficient certificate validation (§IV-E)
	Logger split-world	STH gossiping [44], [43] (not deployed)	Periodic Consistent Broadcast (§IV-D)
	Logger split-world or Logger omission	SCTs from multiple loggers (overhead: log and cert size, validation)	
Subject, monitor(s)	Monitor omission	Use multiple monitors (overhead)	

TABLE I: Comparison of CT and CTng defenses against different attacker models and relevant attacks, including: **stealthy corrupt certificate** (attacker uses a fake identity to deceive victims), **stealthy revocation DoS** (non-revoked certificate 'appears' revoked to victims), **Zombie certificate** (revoked certificate 'appears' non-revoked to victims), **logger omission** (a certificate with an SCT is expected to be in the log, but is not added), **logger split-world** (the logger presents different views of the log to different log clients), and **monitor omission** (monitor does not report a certificate to client). Detailed discussion in §II-B.

	X.509 w/CRL	X.509 w/OCSP	CTv1, CTv2 [37],[38]	CT-VendorRev e.g., Chrome, Safari	CT-VendorRev-wAudit Chrome w/Audit[30]	**CTng**
Security & Privacy Goals						
G1: Existential unforgeability	✓	✓	✓	✓	✓	✓
G2: Accountability	✓	✓	✓	✓	✓	✓
G3: Guaranteed transparency	✗	✗	✓✗[1]	✓✗[1]	✓✗[1]	✓
G4: Guaranteed revocation	✗	✗	✗	✓	✓	✓
G5: Unequivocal revocation	✗	✗	✗	✗	✗	✓
G6: Relying-party privacy	✓	✗	✗	✓	✓✗[2]	✓
Systems Goals						
G7: Evolutionary design	✓	✓	✓	✓	✓	✓
G8: Efficient certificate validation	✓	✗	✗	✓	✓	✓

TABLE II: Comparison of relevant (evolutionary) PKI schemes with respect to goals described and discussed in §III. See §VII for other schemes. **Additional comments:** [1]Both CTv1 and CTv2 assume loggers are benign; however, CTv1 wrongfully states that loggers are not assumed to be trusted; this statement was remediated in CTv2, which explicitly states the benign loggers assumption, but suggests logger redundancy. For more information, see [58]. [2]SCTs by default, are audited using k-anonymous lookup. Privacy exposure only if SCT not known to Google (but should be), or if using Enhanced safe browsing.

valid, while other relying parties consider the certificate revoked.

To understand why unequivocal revocation (G5) is necessary in addition to guaranteed revocation (G4), it is important to clarify that G4 assumes that the issuing CA is *benign*. However, as discussed in §II-B, a misbehaving CA can launch attacks such as the stealthy revocation DoS attack, selectively making a non-revoked certificate appear revoked. Since CTng assumes a stronger adversary model in which the adversary may control any number of CAs, it is important to also define the unequivocal revocation goal (G5), which ensures that even misbehaving CAs cannot carry out such attacks.

An additional concern in PKI design is the risk of compromising the privacy of relying parties. For example, both CTv1 and CTv2 describe an auditing process in which relying parties can query loggers for an STH and proofs of inclusion (PoI) for previously received certificates, in order to verify that the SCT promises made by loggers have been fulfilled. However, this process can reveal the browsing history of relying parties, since it exposes the certificates of the websites they visit. As a result, some relying parties avoid auditing altogether (e.g., Safari), while others (e.g., Chrome) support auditing select SCTs using k-anonymous lookup queries. Unfortunately, auditing all SCTs for all relying parties through this mechanism is not feasible from performance point of view. A similar concern exists with the OCSP revocation mechanism [47], which also requires relying parties to send a request that exposes the certificate they are validating. Thus, we set the following privacy goal:

- **G6:** *Relying party privacy*, i.e., certificates validated by the relying parties are never disclosed to any third party.

C. Systems Goals

One of the key factors that contributed to the widespread adoption of CT was the fact that CT was carefully designed to limit changes within the Web-PKI ecosystem. In general, proposals are more likely to be adopted if they extend existing mechanisms rather than replace them, especially when those mechanisms are already deployed. Moreover, augmenting an existing system allows for targeted improvements while preserving components that function well. For example, the classical PKI mechanisms based on X.509 and PKIX are highly efficient: certificates are compact, certification requires only a single signature operation, and basic certificate validation involves just a few signature verifications with no additional communication. Therefore, we set the following goal:

- **G7:** *Evolutionary design*, i.e., do not introduce new entities and instead propose reasonable modifications to the processes of existing entities that support deployment and facilitate transition from the current system.

That said, some of the currently deployed mechanisms, e.g., revocation, transparency etc., can be improved. Such improvements can help mitigate many challenges; for example, the challenges of scalability and flash crowds, where entities must provide services to an unpredictable yet large number of relying parties. This could affect the certificate validation process, which normally requires real-time operations against such entities. Thus, we set the following goal:

- **G8:** *Efficient certificate validation*, i.e., each certificate can be validated by relying parties using locally available data, without requiring any (real-time) requests to other entities.

IV. CTng Design

In this section, we first provide a high-level overview of how CTng addresses the shortcomings of Web-PKI and the changes it requires. We then detail the design of CTng, focusing on its core functions: certificate issuance (§IV-B), certificate revocation (§IV-C), monitoring and broadcasting (§IV-D), and certificate validation (§IV-E).

A. High Level Overview

CTng introduces changes to the current Web-PKI design that apply to all existing entities: CAs, loggers, monitors, and relying parties. However, aside from the changes introduced for monitors, all other modifications are relatively simple to implement. Most importantly, they can be deployed alongside existing infrastructure, allowing for a manageable transition to CTng. In fact, some of the changes are not mandatory for the deployment of CTng, and CTng can be initially deployed without them. However, these changes can significantly improve the efficiency of the system and, therefore, we believe that they should be considered for deployment, possibly gradually.

The decision to introduce more substantial changes to the monitors is deliberate. Among all entities, monitors are the easiest to modify, given their current deployment status, without disrupting the operation of existing Web-PKI.

We now describe CTng's design by explaining the changes to each type of entity. The design is illustrated in Figure 1b, where the changes with respect to the current Web-PKI are highlighted in red.

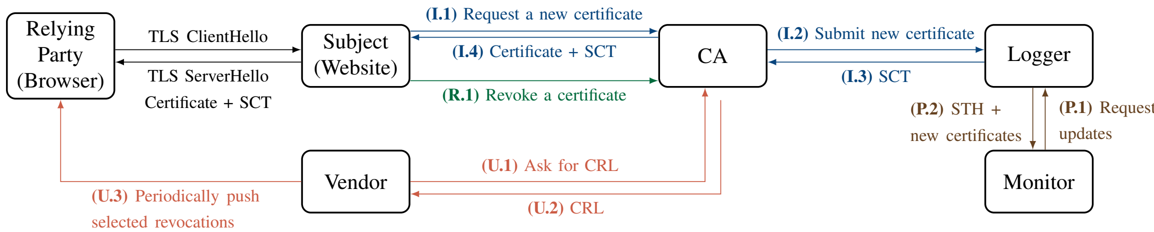
Certificate Authorities (CAs). The process of issuing a new certificate (I.1–I.4) in CTng remains unchanged from the process in CT [35], [36], except for one difference. Specifically, in CT, the CA receives from the logger, and includes in an extension in the certificate, a *Signed Certificate Timestamp (SCT)*, which is an attestation by the logger that the new certificate will be included in the log. In CTng, the SCTs are replaced by the CTng extension, which contains four values. Three of these values are used by relying parties to confirm that the certificate is indeed logged in log ℓ (see §IV-E): a log identifier ℓ, a period number p, and a Proof of Inclusion (PoI).

The fourth value in the CTng extension is called the Revocation Number (RN). The RN is used to implement CTng's improved revocation mechanism. The CTng design is based on the efficient design of [50], in which each CA maintains a bit vector, called the *Certificate Revocation Vector (CRV)*, where each bit represents the revocation status of a single certificate issued by that CA; the RN of a certificate is the index of the bit in the CRV corresponding to that certificate. In the original CRV design [50], the RN is added by the CA; and, like in other currently-deployed revocation mechanisms, correct provision of the revocation information depends on a single entity (the CA for CRV and 'classical' revocation mechanisms such as CRLs and OCSP, and the vendor for the widely-used OneCRL and CRLSet).

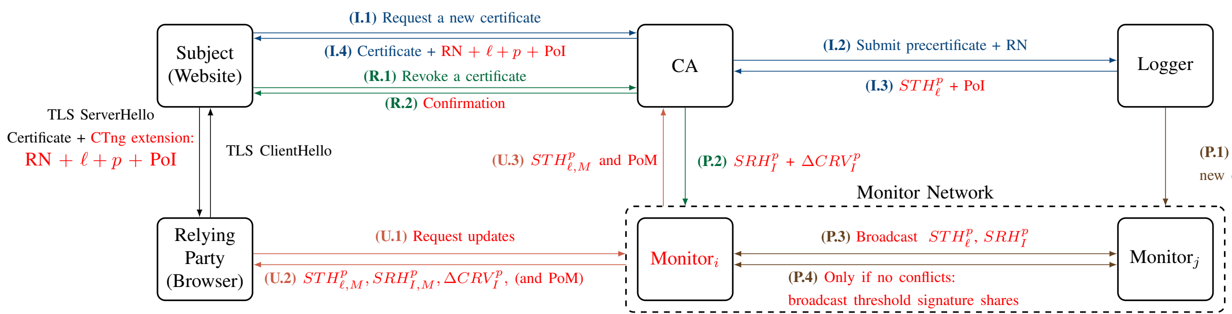
The revocation process (R.1, R.2 and P.2) in CTng differs: there is no single party which is responsible for distributing revocation information. Instead, in CTng, the revocation information, produced and signed by the CA, is also monitored, authenticated, and distributed by the CTng monitors. This avoids delays and failures associated with revocation queries against the CA, a major problem for CRLs and OCSP, and prevents revocation equivocation by a faulty CA/vendor; see §IV-D.

Loggers. Apart from their changed interaction with the CAs (SCTs are no longer issued), another key difference between CT and CTng is that loggers no longer maintain a single large Merkle tree for certificates. Instead, loggers create multiple smaller Merkle trees, each corresponding to certificates issued within a specific period. This approach keeps each tree relatively small, reducing the overhead of PoI verification, which is important since the PoI is part of the CTng extension and must be verified whenever a certificate is validated. Note that in CTng, we could have easily integrated the logging functions with the monitoring functions; the reasons to maintain separate loggers is mostly for backwards compatibility. Loggers also reduce the load on the monitor, allowing each monitor to receive the certificate from a small set of loggers rather than from many CAs; further savings in communication can be obtained using the erasure encoding optimization, see §IV-D2.

Monitors. In CTng, monitors no longer merely *passively* monitor logs (P.1 and P.2); instead, they actively participate in overseeing the behavior of loggers by broadcasting the certificates and STHs they receive from loggers to other monitors using the *Periodic Consistent Broadcast (PCB)* protocol (P.5 and P.6), introduced in §IV-D. The PCB protocol ensures that



(a) Current Web-PKI: X.509 with CT and vendor-assisted revocation information



(b) CTng

Fig. 1: High-level comparison of CT and CTng. **I** - Issuance, **R** - Revocation, **P** - Periodic Consistent Broadcast (Monitoring and Broadcasting), **U** - Update. In the CTng design, illustrated in Figure 1b, we marked in red the changes from the CT design. This graph only presents one possible deployment scenario where the relying party fetches all updates directly from the monitors.

all monitors maintain a consistent view of the logs, which enables them to provide verified periodic updates to the relying parties. PCB also allows benign monitors to generate a Proof of Misbehavior (PoM) against any logger that either fails to provide updates in a timely manner or sends conflicting information (i.e., equivocates), and to ensure correct operation in spite of possible faulty monitors. In addition to retrieving the certificates and transparency information (STHs) from the logs, monitors also retrieve revocation information directly from the CAs (P.2), allowing the monitors to efficiently and securely provide relying parties with up to date revocation information.

Relying parties. In CTng, a relying party periodically retrieves the transparency and revocation updates that were threshold-signed by the monitors (U.1 and U.2). These can be obtained directly from the monitors and can also be cached and served by ISPs or vendors, as they are third-party verifiable. The updates, alongside a CTng-compliant certificate that the client receives during a TLS handshake, allow the client to immediately validate the certificate, in particular, validate that the certificate was logged and was not revoked, without requiring any real-time interactions with the CA, the logger, or the monitors (see §IV-E).

We now provide more details on each of the CTng mechanisms.

B. Issuing and Logging Certificates

For an issuing CA I to generate a certificate for a subject, I first generates and sends a precertificate with a Revocation Number (RN) to a logger ℓ[4]. The RN is a unique sequential identifier[5] that maps a certificate to a revocation status bit in the CRV maintained by I.

Let t_ℓ denote the time on ℓ's clock when ℓ receives the precertificate. The logger ensures that the precertificate is valid and not already in the log, then adds it to a list of pending precertificates to be logged. Periodically, when the time on ℓ's clock is $t_\ell^{STH}(p) \equiv p \cdot MMD + 2 \cdot \Delta_{clk}$, logger ℓ computes the *head* (digest) of the Merkle tree whose leaves are the pending precertificates, each augmented with the RN. Since we bound the maximum clock drift to Δ_{clk} (see §III), in the worst case, an entity's clock may run up to $2 \cdot \Delta_{clk}$ ahead of another entity. Therefore, by waiting an additional $2 \cdot \Delta_{clk}$, we ensure that all benign monitors are ready to receive the update for the same correct period (all benign monitors have entered the pth INIT state, see §IV-D).

We adopt the terminology of CT [36] for the digest of the Merkle tree, referred to as the *Signed Tree Head* (STH). For period p and logger ℓ, we denote it as STH_ℓ^p, defined as:

$$STH_\ell^p = (p, \ell, head, size, \sigma)\,. \tag{1}$$

where $size$ is the number of precertificates and σ is ℓ's signature over the encoding of the other fields: $p \,||\, \ell \,||\, head \,||\, size$.

[4] In CTng, it normally suffices for the CA to use a single logger; the CA can quickly detect if the logger fails to add the precertificate to its publicly available log and, in this rare event, switch to a different logger.

[5] Certificate serial numbers, while also unique, are generated randomly and therefore cannot serve as efficient indices for revocation lookups.

The logger ℓ then sends STH_ℓ^p to I as well as a PoI for each of the precertificates sent by I. This allows CAs to generate a CTng-compliant certificate, i.e., a certificate which includes the PoI of the certificates after validating the PoI against STH_ℓ^p. Logger ℓ will also send STH_ℓ^p and all the precertificates included in STH_ℓ^p to all the monitors.

Similar to the CT logger split-world attack discussed in §II-B, rogue loggers in CTng could present different STH_ℓ^p values to the CA and to the monitors, causing certificates issued by the CA to be considered invalid by the relying party. To prevent a logger from indefinitely delaying the issuance process, the CA should periodically obtain either (i) the STH_ℓ^p threshold-signed by the monitors or (ii) a Proof of Misbehavior (PoM; see §IV-D) against the logger, similar to how relying parties prefetch updates from monitors (see §IV-E). If a PoM is available, or if the monitor threshold-signed STH_ℓ^p does not match the STH_ℓ^p previously received from the logger, the CA should immediately re-log its issued certificates with a different logger. In the case of a mismatch, the CA should also forward the STH_ℓ^p received from the logger to at least $f+1$ monitors to report logger equivocation.

C. Revoking Certificates

The CTng revocation process is initiated by the issuing CA I, usually upon request to revoke the certificate by its subject (the domain owner), and is based on the *Certificate Revocation Vector* (CRV) design of [50]. The vector CRV_I^p is maintained by I and last updated in period p. It contains one entry for each certificate issued by I, with all entries initially set to 0 (not revoked). To mark the revocation of a certificate ψ, the CA sets $CRV_I^p[\psi.CTng.RN] = 1$, where $\psi.CTng.RN$ denotes the *revocation number* of ψ, included in its CTng extension.

Similarly to how the logger informs the monitors periodically, the CA also informs the monitors about revocations periodically; that is, the CA sends these updates whenever its clock shows $p \cdot MMD + 2 \cdot \Delta_{clk}$ for an integer $p > 0$. I first computes its ΔCRV for period p, denoted ΔCRV_I^p, as:

$$\Delta CRV_I^p = CRV_I^p \oplus CRV_I^{p-1} \quad (2)$$

Then, I generates a Signed Revocation Hash (SRH), which contains I's signature over the revocation status of the certificates issued by I. Let h denote a Collision Resistant Hash Function (CRHF)[6]; the SRH of CA I for MMD period p, denoted as SRH_I^p, is defined as:

$$SRH_I^p = (p, I, h(CRV_I^p), h(\Delta CRV_I^p), \sigma) \quad (3)$$

where $h(CRV_I^p)$ and $h(\Delta CRV_I^p)$ are the CRHF outputs on CRV_I^p and ΔCRV_I^p respectively, and σ is I's signature over: $p \mid\mid I \mid\mid h(CRV_I^p) \mid\mid h(\Delta CRV_I^p)$. Finally, I will send both the SRH_I^p and the ΔCRV_I^p to all the monitors.

[6]For simplicity, we show the design using a keyless CRHF, as done by the CT specifications, including in its Merkle tree. Of course, this is secure only under the random oracle model. It is easy to adjust the design for security in the real model, by using a keyed CRHF (also referred to as Any Collision Resistant (ACR) hash function).

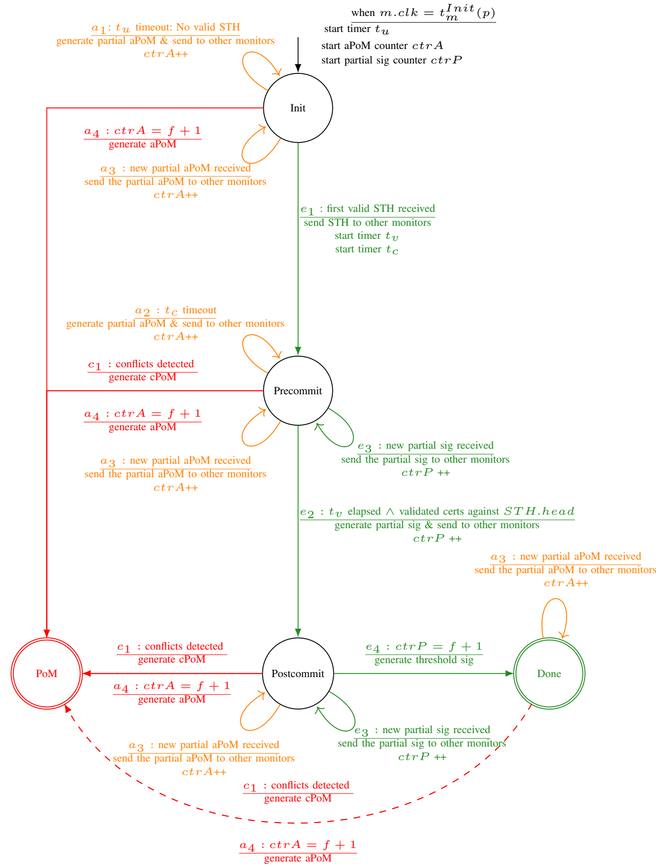


Fig. 2: State diagram of a monitor m running the PCB protocol in period p to process transparency updates from a logger. Green represents the processing of timely and valid STH, which must come with the corresponding set of certificates. Orange represents misbehavior accusations. Red represents the response when misbehavior is proven. *Dashed* transition represents one that may not occur within the same period. Revocation updates from CAs are handled basically in the same way, simply referring to CRVs and corresponding SRHs instead of to STHs.

D. Monitoring and Periodic Consistent Broadcast (PCB)

Each monitor m, wakes up periodically[7], whenever its clock value is $t_m^{Init}(p) \equiv p \cdot MMD$ for integer $p \geq 0$. At this time, the monitor begins to process the STHs and precertificates sent to it during this period by the loggers, as well as the SRHs and ΔCRVs sent to it by the CAs. We refer to this process as the *Periodic Consistent Broadcast (PCB) protocol*.

The goals of the PCB protocol are: (1) to validate receipt of a valid STH_ℓ^p and corresponding precertificates from each logger ℓ, and to produce a threshold-signed version of the STH_ℓ^p, denoted as $STH_{\ell,M}^p$, which is identical in every field

[7]We could instead use different periods for different pairs of monitor and logger/CA to reduce peak load. The modifications are simple but result in some 'writing clutter' (and a bit additional delay), so we simplify by using the same periods for all.

except that its signature field $STH^{p}_{\ell,M}.\sigma$ contains a joint signature by at least $f+1$ monitors, allowing validation by relying parties; (2) similarly, to validate receipt of a valid SRH^{p}_{I} and the corresponding ΔCRV^{p}_{I} from each CA I, and to generate $SRH^{p}_{I,M}$, signed jointly by $f+1$ monitors; and (3) to identify any loggers or CAs that send invalid or no periodic updates, and to generate a corresponding *Proof of Misbehavior (PoM)*.

The operation of the PCB protocol in each monitor is defined by a distinct state machine for every origin (logger or CA); see Figure 2. Since the PCB protocol handles both transparency and revocation information similarly, we focus on the process for handling the STHs and precertificates from a single logger, denoted ℓ, during a single MMD period p. The processing of the SRHs and CRVs is similar.

The monitor wakes up for period p, and enters the INIT state, when its local clock value is $p \cdot MMD$. Since the clock drift is at most Δ_{clk}, monitor m wakes up during the real-time interval $[p \cdot MMD - \Delta_{clk}, p \cdot MMD + \Delta_{clk}]$, which is *before* the time when a benign logger would send the STH and precertificates of period p to all monitors (see §IV-B). This means that the monitor will already be waiting for the STH and precertificates before they arrive from a benign logger.

Upon receiving the (first) valid STH, either from the logger or from another monitor, m sends that STH to the other monitors and transitions from the INIT state to the PRECOMMIT state. By forwarding the STH, m helps mitigate potential attacks in which it receives the STH, but some other benign monitor does not.

The INIT state also handles the case where the benign monitors, and in particular benign monitor m, do not receive the STH in a timely manner. When the PCB state machine starts, a dedicated timer for t_u seconds is begun, where[8]:

$$t_u \equiv \Delta_{com} + 4 \cdot \Delta_{clk} \tag{4}$$

If the t_u timer times out, i.e., when m's clock reaches $p \cdot MMD + t_u$, then m signs a *partial accusation Proof of Misbehavior (partial aPoM)* against logger ℓ and sends it to the other monitors as part of event a_1. If there are $f+1$ or more benign monitors that generate such partial aPoM against ℓ, i.e., that did not receive the STH from ℓ before t_u expired, then each benign monitor will receive at least $f+1$ partial aPoMs against ℓ. Each time a new partial aPoM against ℓ is received, event a_3 increments a counter. When the counter reaches $f+1$, event a_4 is invoked, namely, monitor m computes an aPoM, denoted $aPoM_M$, jointly signed by at least $f+1$ monitors, all stating that logger ℓ is faulty. Then, m transitions to the POM state, as there is no need to continue processing updates from logger ℓ that was attested as faulty.

A rogue logger can also send *different (conflicting) STHs* to different monitors. This case is handled by the PRECOMMIT state. Let t^{Pre}_{m} denote the time on m's clock when it enters the PRECOMMIT state. The PRECOMMIT state has two main goals: to collect the set of precertificates corresponding to the STH and to detect an attack in which a rogue logger sends two conflicting STHs. A conflicting STH pair constitutes a *collision Proof of Misbehavior (cPoM)*; if such a collision occurs, m immediately transitions to the POM (error) state (as part of event c_1). If m reaches the PoM state as the result of a cPoM, m must also sign and collect $f+1$ partial signatures[9] over the logger and the period number[10], similar to an aPoM.

To make sure that m will not sign one STH while another benign monitor signs a different STH, m transitions from the PRECOMMIT state to the next valid state, POSTCOMMIT, only after receiving the set of precertificates corresponding to the STH (see §IV-D1) *and* after its clock shows $t^{Pre}_{m} + t_v$. The value of t_v depends on the connectivity among the monitors; if all monitors are directly connected, then $t_v = 2 \cdot \Delta_{com}$, and in general, $t_v = 2 \cdot \Delta_{com} \cdot d_M$, where d_M is the diameter of the $(f+1)$-connected monitor network. This ensures that when m transitions to POSTCOMMIT, all other monitors have received the same STH, preventing the case that two benign monitors move to POSTCOMMIT with different STHs. The PCB protocol allows some benign monitors to move to POSTCOMMIT and others to terminate with a PoM, as long as all benign monitors that reach POSTCOMMIT agree on the same STH. This behavior does not introduce a vulnerability in CTng; the rogue logger would be known to all benign monitors before the end of next period, and therefore also to all relying parties.

A rogue logger could also fail to send the set of precertificates corresponding to the STH. Monitor m detects this if it does not receive all precertificates by $t^{Pre}_{m}+t_c$, where the timer t_c is set to $t_c = \Delta_{com}$. Benign loggers send the precertificates together with the STH[11], so they should arrive no more than Δ_{com} after the STH. Upon such detection, m signs and sends a *partial aPoM* to the other monitors, since it has determined that ℓ is rogue (event a_3). Monitor m transitions to the POM (error) state if it collects $f+1$ validly signed partial aPoMs from different monitors against ℓ (event a_4).

When m transitions to POSTCOMMIT, it signs the STH using its share of the threshold signing key distributed among

[8]To understand why we use this value of t_u, see the correctness analysis (Claim 2 in §V-A).

[9]Suppose that the relying parties would take into account the cPoM even without a threshold signature. A rogue monitor could send a cPoM to some benign relying party RP, while not sharing the cPoM with benign monitors or with other relying parties. Note, however, that if RP would share the cPoM with the other monitors, this will cause global awareness that the logger is corrupted; in fact, if the exchange between monitor and RP is signed, this will also show that the monitor is corrupt. Attackers may prefer to avoid such exposure. The threshold signature on the cPoM ensures the detection of such attacks without depending on the relying party to submit the cPoM.

[10]The threshold signature only needs to cover the logger identifier and the period number, since a rogue logger could issue more than two conflicting STHs, thereby disrupting the partial signature collection and verification process.

[11]This holds under the simplifying assumption that all loggers provide STHs and certificates directly to all monitors. A slightly larger t_c may be needed if, for efficiency, STHs and certificates are relayed between monitors; details omitted.

the monitors; we refer to this as a *partial signature (sig)*.[12] Monitor m gathers the partial signatures it receives in event e_3, and the PCB completes successfully (transitions to DONE) once m is in the POSTCOMMIT state and has collected $f+1$ partial signatures and therefore has a complete threshold signature (event e_4).

Monitor m could transition to the POM error state either from the POSTCOMMIT state or from the DONE state. This transition is done once m detects a conflicting STH, or collects $f+1$ valid *aPoM* messages signed by different monitors. In any case where m does not reach DONE, it will transition to the POM error state with a valid *Proof of Misbehavior* (either an aPoM or a cPoM) against the rogue logger ℓ.

If m does reach the accepting DONE state, it will hold the threshold-signed STH, denoted $STH^p_{\ell,M}$, along with the corresponding precertificates.

Monitor m follows a similar process to obtain the valid threshold-signed revocation information (ΔCRV^p_I) and $SRH^p_{I,M}$, or a PoM against a rogue CA.

Monitor m provides the $STH^p_{\ell,M}$, ΔCRV^p_I, and $SRH^p_{I,M}$ to relying parties via either *periodic prefetching* (§IV-E) . It also informs any subscribing entity x (e.g., a domain owner) of any logged certificate that matches the profile to which x subscribed (e.g., domain names identical or similar to those owned by x).

We next describe two (optional) optimizations to the PCB protocol. The first is a simple *broadcast optimization*, that reduces the bandwidth usage between monitors, and is used by our implementation by default (§IV-D1). The second optimization, described in §IV-D2, uses erasure encoding to further reduce the amount of data sent from a logger to each monitor to $\frac{2 \cdot |certs|}{n}$ bytes (where n is the total number of monitors). Our evaluation shows mixed results for this optimization, therefore we did not make it the default in the implementation.

1) The Broadcast Optimization

In the PCB protocol as described so far, monitors immediately share the (validated) precertificates they received with other monitors. This would result in transmitting duplicate precertificates between monitors, leading to unnecessary bandwidth consumption. To mitigate this, our implementation deploys, by default, the following simple optimization:

- After a monitor receives from the logger the complete, valid set of precertificates for the current period, it notifies its neighbors.
- A monitor receiving the first such notification, sends back a request for precertificates and begins a timer (for at least $2 \cdot \Delta_{com}$).
- If the timer expires and the monitor still did not receive a complete, valid set of precertificates corresponding to the STH, then the monitor sends requests to any monitor that informed it of the availability of a complete, valid set of precertificates. Until then, the monitor does not send such requests to other monitors (except the first one), but does keep a record of the monitors from whom it received notification of available precertificates.
- A monitor that receives a request for precertificates, responds with the (complete, valid) set of precertificates it had received.
- Once the timer expires, the monitor will change its operation and immediately ask for the precertificates whenever it receives notice of their availability at a peer monitor.

This approach eliminates redundant fetches and sending of precertificates while still ensuring prompt delivery.

[12]Any threshold signature scheme can be used, e.g., [49]. In fact, it suffices to use any public key signature scheme, with each monitor generate its own signing and verification key pair; a set of $f+1$ validly-signed signatures by different monitors over the same message is considered a valid threshold signature.

2) The Erasure Encoding Algorithm (EEA) Optimization

In the EEA optimization of the PCB protocol, loggers and CAs break down the precertificates or compressed CRV data into $k = \lfloor \frac{n}{2} \rfloor$ data shards[13] and generate $n-k$ parity shards, where n is the number of monitors; each shard, say shard i, is sent to a corresponding monitor, denoted m_i. We referred to both the data shards and the generated parity shards as the EEA-encoded shards, whereas only a total of k shards are needed to recover the original data. Since our model requires $n \geq 2 \cdot f + 1$ (see §III-A), this ensures that the system can always tolerate f losses.

To ensure the authenticity of the EEA-encoded shards, the logger/CA constructs another Merkle tree where the n shards are leaves, computes PoI PoI_i for shard i and sends, as part of the update, to monitor m_i, along with the STH^p_ℓ (see Equation 1). The $STH^p_\ell.\sigma$ field is a signature over $p \mid\mid \ell \mid\mid h(head_{EEA} \mid\mid head) \mid\mid size$.

The EEA optimization requires a small change to the PCB protocol as illustrated in Figure 2. Namely, in event e_2, the monitor should reconstruct the precertificate file from k EEA-encoded shares before validation.

E. Efficient Certificate Validation by Relying Parties

Relying parties with reasonable resources and connectivity obtain, every MMD, the STHs of that period (from each logger) as well as the SRHs and ΔCRVs of that period (from each CA). Every logger and CA known to the relying party must be accounted for: if a monitor fails to provide a threshold-signed update for any logger or CA, it must either provide, or must already have provided, a Proof of Misbehavior against that entity.

The relying party would then validate the threshold-signature on these values, ensuring that the values were validated by (at least) $f+1$ monitors—that is, by at least one benign monitor.

Specifically, for all transparency updates, the relying party confirms the validity of the threshold signature $STH^p_{\ell,M}.\sigma$

[13]We initially implemented the EEA version of PCB with $k = f+1$ but $k = \lfloor \frac{n}{2} \rfloor$ performs better under all settings. Results for both K values can be found in the README file of our code repository [25].

over the encoding of the rest of the fields, that is, over: $p \mid\mid \ell \mid\mid head \mid\mid size$ using the group public key of the monitors PK_M. Similarly, for all revocation updates, the relying party confirms that $SRH^p_{I,M}.\sigma$ is a valid threshold signature over the rest of the fields, that is, over: $p \mid\mid I \mid\mid h(CRV^p_I) \mid\mid h(\Delta CRV^p_I)$.

Relying parties should validate cPoMs and aPoMs, ensuring that each of them is signed by at least $f+1$ monitors (over $p \mid\mid \ell$ or $p \mid\mid I$, depending on the originator entity type). Each cPoM should contain two different updates for the same period and both signed by the same (faulty) logger or CA.

Obtaining the periodic updates can be done either directly from one of the monitors or from another source; since updates are threshold-signed and timestamped, their authenticity and freshness are ensured even if obtained from a source which is not trustworthy. For example, the information can be downloaded from a monitor, validated and then cached and provided by a browser vendor or the client's ISP, or distributed as DNS records; the resource requirements on the monitors would be minimal.

If any verification fails or if any logger/CA was not accounted for, the relying party considers the monitor or other source from which it received the update as faulty, and restarts the process with a different monitor (or other source).

After successful verification, the relying party updates its local storage with the newly received information. For each transparency update $STH^p_{\ell,M}$, the relying party stores the logger identifier $STH^p_{\ell,M}.\ell$, the period number $STH^p_{\ell,M}.p$, and the Merkle digest $STH^p_{\ell,M}.head$. For each revocation update ($SRH^p_{I,M}$ and ΔCRV^p_I), the relying party updates the CRV [14]corresponding to $SRH^p_{I,M}.I$ in its local storage:

$$localCRV[SRH^p_{I,M}.I] \leftarrow localCRV[SRH^p_{I,M}.I] \vee \Delta CRV^p_I .$$

The prefetched information allows the relying party to perform efficient certificate validation as follows. Upon establishing a TLS connection with a subject (website) that uses a CTng-compliant certificate ψ, the relying party first confirms that ψ is PKIX-valid, as per [11], e.g., correctly signed, trust-anchored, not expired, etc. The relying party then searches its local storage for the head $Localheads[\ell, p]$ corresponding to logger $\psi.CTng.\ell$ in period $\psi.CTng.p$ and extracts the precertificate portion Φ from ψ. Then, it computes $h(\Phi)$ and confirms that $Localheads[\ell, p]$ can be reconstructed from $h(\Phi)$ and $\psi.CTng.PoI$[15]. Finally, the relying party checks that ψ has not been revoked by inspecting its locally maintained CRV for the issuing CA $\psi.I$ at index $\psi.CTng.RN$, confirming that: $LocalCRV[\psi.I][\psi.CTng.RN] = 0$.

V. Analysis

In this section, we analyze how CTng, with the periodic prefetch mechanism (§IV-E), achieves the goals listed in §III against an attacker who controls an arbitrary set of loggers and CAs, as well as up to f monitors.

A. Timing and Correctness Analysis

We now analyze the timeline of events in CTng, showing that CTng ensures *correctness:* issued certificates become valid (until revoked or expired) within bounded time, and revoked certificates become invalid within bounded time (see Figure 3). The analysis provides time bounds, which we also use in the design of CTng. The analysis is mostly done in terms of real-time intervals; we denote different real-time values by adding subscripts and superscripts, as needed, to the symbol τ. The bound of Δ_{clk} on the clock bias allows us to bound the real-time when the local clock of an entity shows any given value, e.g., C, as:

$$\tau_C \in [C - \Delta_{clk}, C + \Delta_{clk}] \tag{5}$$

In particular, let $C_p \equiv p \cdot MMD + 2 \cdot \Delta_{clk}$ for an integer $p \geq 0$, and τ^ℓ_p denote the time when ℓ's local clock shows C_p. From Equation 5, we have:

$$\begin{aligned} \tau^\ell_p &\in [C_p - \Delta_{clk}, C_p + \Delta_{clk}] \\ &= [p \cdot MMD + \Delta_{clk}, p \cdot MMD + 3 \cdot \Delta_{clk}] \end{aligned} \tag{6}$$

For simplicity, the timing analysis, whose results are summarized by Theorem 1, assumes that the execution of all entities begins at real-time zero.

Theorem 1. Let I be a benign CA, ℓ be a benign logger used by I, m be a benign monitor, RP be a benign relying party using m, and d_M be the diameter of the $f+1$-connected monitor network. Then:

1) Let Φ be a precertificate issued by I at time τ_I, and suppose RP validates the corresponding certificate ψ at time $\tau_{RP} > \tau_I + 2 \cdot MMD + 4 \cdot \Delta_{clk} + (4 + 3 \cdot d_M) \cdot \Delta_{com}$. Then RP will determine ψ to be valid, provided that ψ has not expired (i.e. $\psi.to \geq \tau_{RP} + \Delta_{clk}$) and that ψ was not revoked by I until τ_{RP}.
2) Let ψ be a certificate that I issued (at τ_I) and revoked at time τ_R ($\tau_R > \tau_I$). Suppose RP validates ψ at time $\tau_{RP} > \tau_R + 2 \cdot MMD + 4 \cdot \Delta_{clk} + (3 + 3 \cdot d_M) \cdot \Delta_{com}$. Then RP will determine ψ to be invalid.

Proof. We only present the argument for the first statement; the second statement follows similarly as shown in the full version of the paper [30].

The proof is by a series of claims, following the events in the handling of Φ (and ψ) and the timeline as illustrated in Figure 3. In Claim 1, we analyze the period when ℓ handles and forwards Φ (and the corresponding STH) issued at τ_I. In Claim 2, we derive the bound for the update timer t_u and show that the t_u timeout event will never happen if the logger is benign. Claim 3 shows that all benign monitors move to DONE state, with a complete threshold signature over the STH, before $\tau_I + MMD + 2 \cdot \Delta_{clk} + (2 + 3 \cdot d_M) \cdot \Delta_{com}$.

Finally, let ψ be the full certificate corresponding to the precertificate Φ, and, in particular, containing the PoI generated by the benign logger ℓ. Claim 4 completes the proof, by

[14]Note that here we use $\vee$ instead of $\oplus$ because revoked certificate should stay revoked.

[15]Same verification method as the SCT auditing [36], but relying parties in CTng do not need to rely on the logger to provide the Merkle root and PoI.

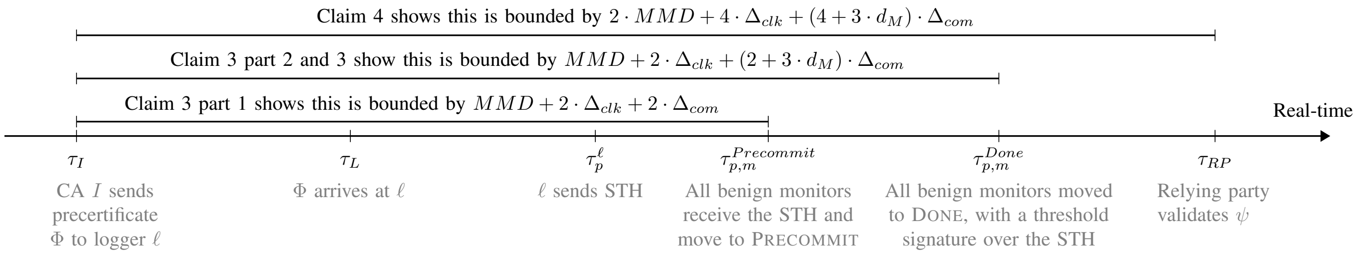


Fig. 3: Timeline visualizations used in the proof of correctness (Theorem 1).

showing that the benign relying party RP will consider ψ as valid at time $\tau_{RP} > \tau_I + 2 \cdot MMD + 4 \cdot \Delta_{clk} + (4+3 \cdot d_M) \cdot \Delta_{com}$, provided that $\psi.to \geq \tau_{RP} + \Delta_{clk}$ and that ψ was not revoked by I until τ_{RP}. □

Claim 1. Let τ_I denote the time when the CA sent Φ to logger ℓ and let p denote the period number included in the first STH which ℓ sends after receiving Φ. Then $\tau_p^\ell \in [\max\{\tau_L, C_p - \Delta_{clk}\}, C_p + \Delta_{clk}]$ and $p \in [p_0, p_1]$, where:

$$p_0 \equiv \left\lceil \frac{\tau_I - 3 \cdot \Delta_{clk}}{MMD} \right\rceil,\ p_1 \equiv 1 + \left\lfloor \frac{\tau_I - \Delta_{clk} + \Delta_{com}}{MMD} \right\rfloor \quad (7)$$

Argument: Immediately from the bound Δ_{com} on the delay, we know that logger ℓ receives Φ during $[\tau_I, \tau_I + \Delta_{com}]$; let τ_L denote the time when ℓ received Φ. From the design, we know that ℓ sends the next STH when its clock shows C_p, with the STH containing the period number p. From Equation 6 and the fact that this is the STH sent *after* τ_L, we obtained $\tau_p^\ell \in [\max\{\tau_L, C_p - \Delta_{clk}\}, C_p + \Delta_{clk}]$.

We next show that $p \in [p_0, p_1]$ where p_0, p_1 are as in Equation 7.

Since $\tau_I \leq \tau_L \leq \tau_I + \Delta_{com}$, it follows from Eq. 5 that:

$$\tau_I \leq C_p + \Delta_{clk} = p \cdot MMD + 3 \cdot \Delta_{clk} \quad (8)$$

Since Φ must be received after the previous STH:

$$\begin{aligned} \tau_I &\geq \tau_{p-1}^\ell - \Delta_{com} \geq C_{p-1} - \Delta_{com} - \Delta_{clk} \\ &= (p-1) \cdot MMD + \Delta_{clk} - \Delta_{com} \end{aligned} \quad (9)$$

From Equations (8) and (9), we can derive bounds for p:

$$\frac{\tau_I - 3 \cdot \Delta_{clk}}{MMD} \leq p \leq 1 + \frac{\tau_I - \Delta_{clk} + \Delta_{com}}{MMD} \quad (10)$$

Since p is an integer, we have:

$$p_0 \leq p \leq p_1, \text{ defined in Equation 7} \quad (11)$$

□

We next show that monitor m will receive STH_ℓ^p and Φ before the t_u timer times-out, where $t_u \equiv \Delta_{com} + 4 \cdot \Delta_{clk}$.

Claim 2. Let $\tau_{p,m}^{Init}$ denote the time of the p^{th} INIT in m. Then:

1) $\tau_{p,m}^{Init} \in [p \cdot MMD - \Delta_{clk}, p \cdot MMD + \Delta_{clk}]$.
2) For every period $p > 0$ and benign logger ℓ, monitor m receives a valid STH_ℓ^p during $[\tau_{p,m}^{Init}, \tau_{p,m}^{Init} + t_u]$, that is, the t_u timeout event (action a_1 of the INIT state) is never invoked for the benign logger ℓ.

Argument: Monitor m begins period p when its local clock shows $p \cdot MMD$, i.e., following Eq. 5, at $\tau_{p,m}^{Init} \in [p \cdot MMD - \Delta_{clk}, p \cdot MMD + \Delta_{clk}]$, which is the first part of the claim. Equivalently:

$$p \cdot MMD - \Delta_{clk} \leq \tau_{p,m}^{Init} \leq p \cdot MMD + \Delta_{clk} \quad (12)$$

To prove the second item, we first note that a benign logger ℓ sends the STH for the p^{th} period, i.e., STH_ℓ^p, when its clock shows $p \cdot MMD + 2 \cdot \Delta_{clk}$, i.e., after real-time $p \cdot MMD + \Delta_{clk}$. From the RHS of Equation 12, this cannot happen *before* $\tau_{p,m}^{Init}$.

To prove the second item, it remains to show that the STH is not received *after* $\tau_{p,m}^{Init} + t_u$, where $t_u \equiv \Delta_{com} + 4 \cdot \Delta_{clk}$. The *latest* real-time at which ℓ will send the STH would be $p \cdot MMD + 3 \cdot \Delta_{clk}$, therefore, m receives the STH at or before real-time $p \cdot MMD + 3 \cdot \Delta_{clk} + \Delta_{com}$. By substituting $p \cdot MMD \leq \tau_{p,m}^{Init} + \Delta_{clk}$ (LHS of Equation 12) we see that m receives the STH before $\tau_{p,m}^{Init} + 4 \cdot \Delta_{clk} + \Delta_{com}$, as required.

□

We now bound the times for the different steps of the PCB protocol in benign monitors, providing us with the desired bound on the time until the monitors have a valid STH for certificates issued by benign CAs, using a benign logger.

Claim 3. Let $\tau_{p,m}^{Precommit}$, $\tau_{p,m}^{Postcommit}$ and $\tau_{p,m}^{Done}$ denote the time when benign monitor m enters its p^{th} PRECOMMIT, POSTCOMMIT and DONE state. Let d_M denote the diameter of the $(f+1)$-connected monitor topology. Then:

1) Any precertificate sent to a benign logger ℓ at time τ_I, and the corresponding STH, are received by m at time $\tau_{p,m}^{Precommit} \leq \tau_I + MMD + 2 \cdot \Delta_{clk} + 2 \cdot \Delta_{com}$.
2) All benign monitors generate their partial signature over STH_ℓ^p and move to POSTCOMMIT at time $\tau_{p,m}^{Postcommit} \leq \tau_I + MMD + 2 \cdot \Delta_{clk} + (2+2 \cdot d_M) \cdot \Delta_{com}$.
3) All benign monitors move to DONE state, with a complete threshold signature over the STH, at time $\tau_{p,m}^{Done} \leq \tau_I + MMD + 2 \cdot \Delta_{clk} + (2 + 3 \cdot d_M) \cdot \Delta_{com}$.

Argument: Recall that, by Equation 6, ℓ sends the STH at or before

$$p \cdot MMD + 3 \cdot \Delta_{clk}$$

From Claim 1, we have $p \leq p_1$, where (by Equation 7)

$$p_1 \equiv 1 + \left\lfloor \frac{\tau_I - \Delta_{clk} + \Delta_{com}}{MMD} \right\rfloor .$$

The time τ_p^ℓ at which ℓ sends Φ and the corresponding STH is therefore not later than

$$\tau_I + MMD + 2 \cdot \Delta_{clk} + \Delta_{com}.$$

Hence, we have an upper bound on the time at which monitor m receives the STH and certificates, denoted $\tau_{p,m}^{Precommit}$:

$$\tau_{p,m}^{Precommit} \leq \tau_I + MMD + 2 \cdot \Delta_{clk} + 2 \cdot \Delta_{com}.$$

This proves part 1 of the claim.

In the PRECOMMIT state, the benign monitors wait for $t_v = 2 \cdot d_M \cdot \Delta_{com}$ time for a potential conflicting STH. Such a conflicting STH will not be received, since ℓ is benign and will not send a conflicting STH (and the attacker cannot send a fake yet valid conflicting STH). Therefore, the time $\tau_{p,m}^{Postcommit}$ at which a benign monitor m moves to POSTCOMMIT is at most: $\tau_{p,m}^{Postcommit} \leq \tau_I + MMD + 2 \cdot \Delta_{clk} + (2 + 2 \cdot d_M) \cdot \Delta_{com}$, proving part 2 of the claim.

At this point, each benign monitor will generate its partial signature of the STH, and send it to all the other monitors. Let $\tau_{p,m}^{Done}$ denote the time at which benign monitor m will have the necessary $f+1$ partial signatures, and move to DONE; this would occur within only one more $d_M \cdot \Delta_{com}$, i.e., $\tau_{p,m}^{Done} \leq \tau_I + MMD + 2 \cdot \Delta_{clk} + (2 + 3 \cdot d_M) \cdot \Delta_{com}$. □

Claim 4. Let ψ be the certificate corresponding to the precertificate Φ, both issued by benign CA I using the benign logger ℓ. Suppose that a benign relying party RP, which uses a benign monitor m, validates ψ at time $\tau_{RP} > \tau_I + 2 \cdot MMD + 4 \cdot \Delta_{clk} + (4 + 3 \cdot d_M) \cdot \Delta_{com}$. Then RP will determine ψ to be valid, provided that $\psi.to \geq \tau_{RP} + \Delta_{clk}$ and that ψ was not revoked by I until τ_{RP}.

Argument: Since ℓ and the CA I are both benign, we know that ψ will contain a valid PoI for STH_ℓ^p.

A benign RP asks for a new STH once every MMD. Therefore, when it validates ψ at τ_{RP}, the RP should already have requested and received from m all the STHs signed by the monitors until $\tau_{RP} - MMD - 2 \cdot \Delta_{com} - 2 \cdot \Delta_{clk}$; notice that we allow here for the (extremely unlikely) case where the RP's clock was behind by Δ_{clk} when it requested the previous update (including STH) from m, and that it was ahead by Δ_{clk} at τ_{RP}.

From Claim 3, all benign monitors, including m, have a complete threshold signature of the STH which covers Φ, allowing successful validation of ψ, by $\tau_I + MMD + 2 \cdot \Delta_{clk} + (2 + 3 \cdot d_M) \cdot \Delta_{com}$. Since $\tau_{RP} > \tau_I + 2 \cdot MMD + 4 \cdot \Delta_{clk} + (4 + 3 \cdot d_M) \cdot \Delta_{com}$, then, at τ_{RP}, the RP should already have this (signed) STH, and would determine ψ to be valid. □

B. Security and Privacy Goals

Theorem 2. CTng ensures the security and privacy Goals (**G1–G6**).

Proof. Due to the length constraint, we only provide here a high-level argument; for more details, see the full version of the paper [30].

In [58], Wrótniak et al. showed that both PKIX and CT satisfy existential unforgeability (**G1**) and accountability (**G2**) by assuming less restrictive assumptions than in §III-A. Since CTng augments CT and builds upon the core functionality of PKIX/CT used in [58], we prove in [30], by reduction, that CTng also satisfies **G1** and **G2**.

CTng achieves $\Delta-$guaranteed transparency (**G3**) for $\Delta = 3 \cdot d_m \cdot \Delta_{com}$, since a benign relying party considers ψ as valid only if ψ has a valid PoI in the corresponding STH, which must be threshold signed by the monitors, i.e. signed by at least one benign monitor. Once a benign monitor signs the STH, the monitor relays the STH, and then forwards also the corresponding precertificates to benign monitors that request them; so all benign monitors should receive both STH and precertificates within $3 \cdot d_m \cdot \Delta_{com}$, assuming an $f+1$ connected monitor network.

Theorem 1 shows that every revocation will be known to every relying party after at most $\Delta = 2 \cdot MMD + 4 \cdot \Delta_{clk} + (3 + 3 \cdot d_M) \cdot \Delta_{com}$. Based on this, in [30] we show that CTng achieves $\Delta-$guaranteed revocation (**G4**).

CTng ensures unequivocal revocation (**G5**), since any adversary that breaks **G5** in CTng is either using an insecure threshold signature scheme or does not guarantee the models assumptions described in §III-A.

Finally, CTng obviously preserves relying-party privacy (**G6**), because the certificate validation process in CTng does not disclose any information to any third party.

□

C. System Goals

G7: Evolutionary design. CTng does not require any additional entities compared to CT and introduces only modest changes to the roles and processes of existing CT entities. In particular, CT only expands the roles of monitors and requires CA/browser support for the new CTng extension.

G8: Efficient certificate validation. In CTng, the decision to accept or reject a certificate is made entirely by the relying party, based solely on information provided with the certificate or stored locally as shown in §IV-E. As a result, relying parties are not required to initiate any network communication with any entities during the validation process, beyond receiving the certificate itself.

VI. EVALUATION

We begin with a description of our evaluation setup of the experiments (§VI-A). Then, we evaluate the performance impact on each entity individually (§VI-B), followed by experimentally evaluating the performance and scalability of our CTng prototype implementation as a complete system (§VI-C).

A. Experimental Setup

We implemented a prototype of the PCB protocol [25] in Go 1.23, with and without the erasure-encoding algorithm

Asymptotic Overhead	System	CAs	Loggers	Monitors
Computation per MMD	CT + CRLset	$O(N_{\text{cpm}} + N_{\text{rpm}})$	$O(f \cdot N_{\text{cpm}} \cdot \log(N_{\text{c}}))$	$O(f \cdot N_{\text{cpm}} \cdot \log(N_{\text{c}}))$
	CTng	$O(N_{\text{cpm}} + N_{\text{rpm}})$	$O(N_{\text{cpm}} \cdot \log(N_{\text{cpm}}))$	$O(N_{\text{cpm}})$
Communication per MMD	CT + CRLset	$O(f \cdot N_{\text{cpm}} + N_{\text{rpm}})$	$O(f \cdot N_{\text{m}} \cdot N_{\text{cpm}})$	$O(f \cdot N_{\text{m}} \cdot N_{\text{cpm}})$
	CTng	$O(N_{\text{cpm}} + N_{\text{m}} \cdot N_{\text{rpm}})$	$O(N_{\text{m}} \cdot N_{\text{cpm}})$	$O(N_{\text{m}} \cdot (N_{\text{cpm}} + N_{\text{rpm}}))$
Storage	CT + CRLset	$O(N_{\text{r}})$	$O(f \cdot N_{\text{c}})$	$O(N_{\text{cpm}})$
	CTng	$O(N_{\text{r}})$	$O(N_{\text{c}})$	$O(N_{\text{cpm}} + N_{\text{r}})$

TABLE III: Asymptotic overhead per MMD for CT (as defined in [36]) with CRLset and CTng. Each entry shows the complexity in terms of key system parameters: N_{cpm} (number of all new precertificates issued per MMD), N_{c} (total number of certificates in all logs), N_{rpm} (number of all new revocations per MMD), N_{r} (total number of revoked certificates) N_{ca} (number of CAs), N_{m} (number of monitors), and f (security parameter).

Parameter	Value	Justification
Number of loggers	8	As of January 2025, all publicly usable CT logs are operated by six organizations, with each organization running 1–3 logs at any given time [1].
Precertificate simulation	Random string of size 2000 Bytes	Based on [57], where the average size of precertificate is estimated to be 1570 Bytes.
Precertificate workload	$400K$ certs/hr (uniformly distributed among the loggers)	Cloudflare observed [9] global precertificate throughput of $380K - 390K$ unique precertificates per hour in December 2024.
Number of CAs	100	Number of CAs does not have a noticeable impact [50].
Total number of certificates	100 million	Based on [50]. The daily revocation rate was not explicitly stated in [50], so we confirmed it with the authors.
Daily revocation rate	0.02%	
Total revocation rate	1%	
MMD interval	10 minutes	This is the value data-mined from our own experiments, which leaves enough safety margin across our test suite.
Number of monitors	32	A reasonable upper bound on the number of monitors [2].
Number of monitor faults allowed (f)	3	Slightly larger than CT's minimum fault tolerance (2SCTs [22] $\to f = 1$).
Access link capacity	1000 Mbps	Standard link capacity.
Topology	Star topology	Simulates the internet back bone.

TABLE IV: Baseline experiment setup.

described in §IV-D2. We evaluated our prototype using the Sphere testbed [42], where the test environment consisted of virtual machines running Ubuntu 22.04, each equipped with 8 cores and 16 GB of RAM. The baseline settings for all experiments are in Table IV (with justifications for these choices).

B. Performance per Entity

We analyzed and compared the asymptotic overhead of each type of entity in CTng with respect to the overheads in CT, and summarized the results in Table III. The overhead of different CAs, loggers and monitors may differ widely depending on their usage; to deal with that, we present the overall overhead for all CAs, all loggers and all monitors. The overheads will also depend on the distribution of certificates and revocations between the different CAs and loggers; for simplicity, our computations assume the worst case where all certificates are by a single CA, which is using a fixed set of loggers.

As we detail below, CTng does not introduce undue overhead for system entities, especially when weighted against the added security benefits and their existing responsibilities. Each role remains feasible in terms of performance and operational complexity. Note also that since CT and CTng rely on the same types of entities, the incentives to operate a CTng entity would be similar to the incentives to operate the corresponding CT entity.

CAs: In both CT and CTng, the computation overhead per MMD is $O(N_{\text{cpm}} + N_{\text{rpm}})$, where N_{cpm} is the number of all new precertificates issued per MMD, and N_{rpm} is the number of all new revocations per MMD. However, since CT uses log redundancy, its communication overhead per MMD is $O(f \cdot N_{\text{cpm}} + N_{\text{rpm}})$, as certificates need to be logged over multiple loggers, and the CRLsets mechanism needs to learn about all newly revoked certificates. In comparison, CTng's communication overhead is $O(N_{\text{cpm}} + N_{\text{m}} \cdot N_{\text{rpm}})$, where N_{m} is the number of monitors, since certificates are logged with only a single logger, which then sends the SRH and ΔCRV to all monitors.

The storage requirements in both CT and CTng are $O(N_{\text{r}})$, where N_{r} is the total number of revoked certificates. This is because CAs do not need to store the certificates they have issued, but must retain those they have revoked. Although the storage overhead is asymptotically the same, the CRV approach used in CTng is significantly more space-efficient, requiring approximately 1.8 bits per revocation compared to CRLite's 6.6 bits, CRLset's 110 bits, and OneCRL's 1928 bits per revocation [34], [50].

Loggers: The computational overhead for CT loggers is dominated by the need to update the Merkle tree by adding, in each MMD, $f \cdot N_{cpm}$ new precertificates, where the factor of $O(f)$ comes from the logger redundancy required. The computational cost for each precertificate is proportional to the height of the tree, i.e., $\log(N_c)$. The $O(f)$ factor similarly impacts the communication and storage complexities of CT. In contrast, in CTng, we avoid the $O(f)$ factor; the computational cost is mainly due to the need to generate the PoI for each precertificate. Also, the height of the tree is only $O(\log(N_{cpm}))$, since CTng uses separate trees for each MMD period.

Monitors: The computational and communication costs for CT monitors (if implemented as the RFC [35], [36] prescribe) are similar to these of CT loggers. In case of CT, these are dominated by the need to receive f copies of each precertificate and to add $f \cdot N_{cpm}$ new precertificates to the Merkle tree of certificates[16]. For CTng, the costs are significantly reduced since we avoid the need to receive and handle f redundant copies of each precertificate, and since the Merkle trees are much smaller (per MMD). CTng has an additional overhead of also handling revocations; in practice, the overhead due to revocations is negligible compared to the overhead of certificates.

The storage requirements of CT and CTng monitors are the same for issued certificates[17]. CTng also stores the ΔCRV identifying newly revoked certificates. Note that Table III does not explicitly reflect the additional overhead of the activities that CTng monitors do and CT monitors do not: the PCB protocol and client prefetching. The reason is that the additional PCB protocol overhead is dominated by the aforementioned asymptotic overhead of handling the certificates. Also, clients can fetch the (signed) data from intermediaries such as CDNs, ISPs and browser vendors; they do not have to obtain it directly from monitors.

Subjects (Websites): The changes CTng introduce with respect to subjects are insignificant in terms of overhead. Notice, however, that issuing certificates in CTng takes more time than in CT (see Theorem 1).

Relying Parties: We measured the prefetching method (§IV-E) for an MMD of 1 day and a standard revocation rate of 1%, and found that, for CTng, the daily communication overhead is 249 KB and the storage overhead is 2.33 MB; this is for the entire set of revocations. In comparison, Chrome's implementation of CT with CRLsets incurs a daily bandwidth cost of 250 KB [34], for only 2% of the certificates.

Next, we measured the per connection communication and computation overheads. The communication overhead is similar; a CT extension with the minimum 2 SCTs is 222 B when using ECSDA and 594 B when using RSA-2048, while a CTng extension is 780 B [18] .However, CT requires significantly more computations to validate a certificate: approximately 25 ms using RSA 2048 and 50 ms using ECDSA, compared to only about 0.28 μs in CTng. The reason for this significant difference is the fact that CTng requires only hash computations and no (real-time[19]) signature verification, unlike CT, which uses log redundancy, and each SCT requires a public key signature verification.

C. Performance and Scalability of the System as a Whole

To evaluate the performance of our prototype, we measured the impact of several key system parameters on the *maximum convergence time*, defined as the duration required for the last benign monitor in the network to reach the "done" state described in Figure 2. This state indicates that all transparency information is fully prepared to be served to a relying party. The maximum convergence time effectively represents the smallest MMD that our system can support. In each experiment, we use the base settings from Table IV, except for the specific parameter being evaluated to assess its impact.

The effect of increasing the number of monitors. In Figure 4a, we plot the impact of increasing the number of monitors from 8 to 32. Our results show that CTng can easily support even a large number of monitors, as the maximum convergence time for 32 monitors is only 32.77 seconds. Furthermore, we observe that CTng scales efficiently in the number of monitors: quadrupling the number of monitors from 8 to 32 does not result in any noticeable increase in convergence time; the results are all within our (quite tight) error margins.

The effect of increasing number of monitor faults allowed (Figure 4b). Using the baseline setting of 32 monitors, we experimented with increasing the monitor fault tolerance (f) from 1 to 8, both with and without the erasure encoding algorithm. As shown in Figure 4b, CTng performs well under both configurations, with a convergence time increasing up to 67.11 seconds without erasure encoding, and increasing very slightly, from 32.51 to 34.51 seconds, with erasure encoding. We observe that when the number of faulty monitors is small, the overhead introduced by erasure encoding is not justified; in fact, the version without erasure encoding outperforms the encoded version. However, when the number of faulty monitors exceeds 4, erasure encoding becomes more effective.

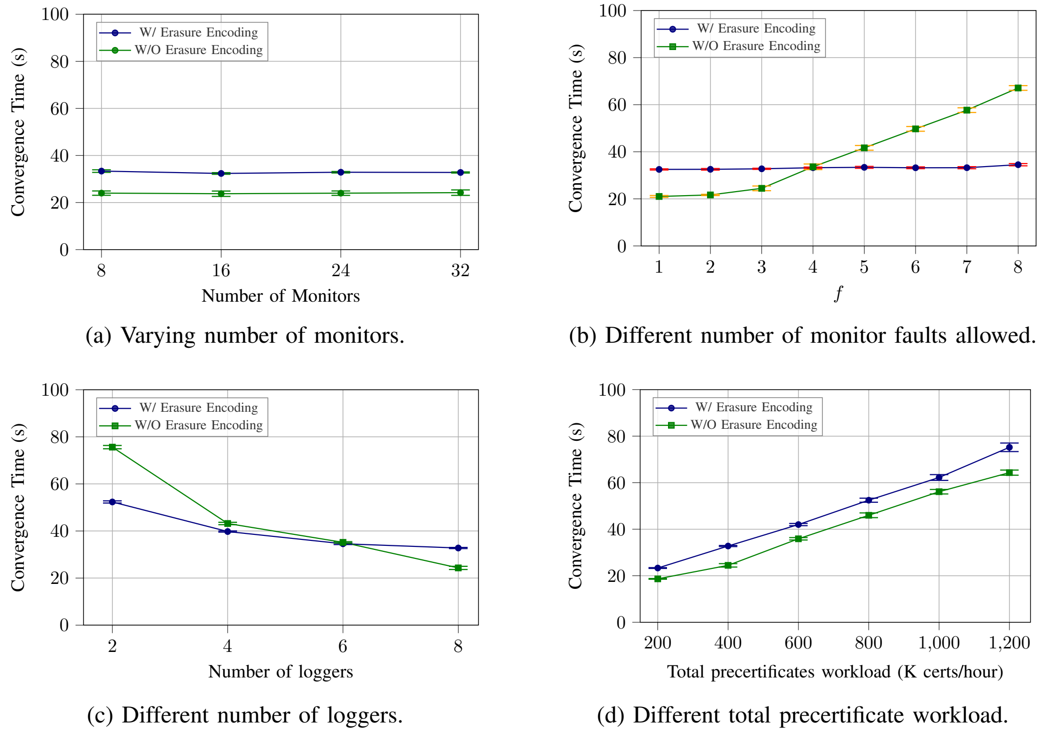
The impact of increased workload on loggers (Figure 4c). Our workload baseline setting uniformly distributes a workload of 400 K precertificates per hour across 8 loggers. To evaluate the impact of increased workload per logger, we

[16] In practice, loggers may maintain multiple Merkle trees.

[17] Notice that in Table III the storage requirement considers N_{cpm} and not N_{c}, because monitors do not have to store all certificates; of course, some monitors might store certificates.

[18] Assuming an MMD of one day and a precertificate generation rate of 400 K certs/hr [9], the logger's Merkle tree at the end of the MMD period would contain approximately 9.6 M leaves (certificates), i.e., each certificate's PoI consists of 24 hashes. Assuming the hash function used is SHA-256 and adding the RN (4 B), the logger ID ℓ (4 B), and the period number p (4 B), the total size of a CTng extension is 780 B ($= 32\,\text{B} \times 24 + 12\,\text{B}$)

[19] Relying parties would require a few additional signature verifications when prefetching that data, but these operations are negligible, since they are performed once per MMD; also, these operations are done in advance, rather than during the connection.



(a) Varying number of monitors.

(b) Different number of monitor faults allowed.

(c) Different number of loggers.

(d) Different total precertificate workload.

Fig. 4: System convergence time measured under four distinct experimental conditions.

performed experiments in which the same total workload was distributed among fewer loggers. As shown in Figure 4c, our findings indicate that although fewer loggers introduce a bottleneck, CTng still converges within 75.62 seconds with only two loggers each handling a workload of 200 K precertificates per hour.

The impact of increased workload on the system (Figure 4d). In Figure 4d, we plot the measured convergence time of our prototype under workloads ranging from 200 K to 1,200 K precertificates per hour. Our results show that the convergence time of CTng scales linearly with the overall workload processed by the monitor network. Under the baseline workload of 400 K precertificates per hour, the system converges within 32.77 seconds. Even at a workload of 1,200 K precertificates per hour, CTng converges within 77.04 seconds. In other words, CTng maintains a reasonable convergence time even with a safety factor of three, which is important not only for accommodating future growth but also for handling additional overheads that may not be reflected in our experimental setup.

In summary, we have demonstrated that: (1) CTng scales linearly with the input workload; (2) CTng can support MMDs in the order of minutes, with a good safety margin; (3) The convergence time of CTng does not significantly increase due to an increase in the number of monitors; (4) Most importantly, with erasure encoding, convergence time increases only mildly even as the number of tolerated faulty monitors grows. In contrast, CT incurs substantial overhead when tolerating additional faulty loggers.

VII. Related Work

Several works focus on extending CT with an auditing mechanism; some of these also address the privacy risks associated with auditing. Examples include CT gossip [44], [43] and CTor [13], which enhance CT by integrating SCT auditing, STH gossiping, and conflict reporting to detect logger misbehavior in a timely manner. However, these mechanisms face scalability challenges as the number of participating entities grows. Other designs, such as [46], [18], [51], [26], emphasize privacy-preserving solutions for SCT auditing that account for the risk of a client's activity being revealed through either querying or reporting during the audit process.

Vendor-assisted revocation mechanisms such as CRLsets [21] and OneCRL [20] attempt to reduce revocation overhead by collecting and distributing a curated subset of certificates selected through proprietary policies, thus eliminating the need for relying parties to fetch large CRLs from a third party (e.g., CA). However, both adopt a soft-fail mechanism, under which a certificate is accepted if its revocation status is unknown [34], and they cover only a small subset of certificates. Subsequent work, such as CRV [50] (used in our system) and CRLite [34], improves encoding efficiency, enabling them to cover all certificates with reasonable overhead.

Other works focus on ensuring security with robustness to up to f rogue entities. In COCA [63], certificates must be co-signed by multiple CAs. ARPKI [5] extends this model by requiring additional synchronization across servers. A consensus model for CT logs, presented in [57], prevents log split-world attacks but introduces significant overhead. Furthermore, works such as [28], [3] leverage blockchain technology to achieve certificate and/or revocation transparency by completely decoupling these processes from traditional PKI entities like CAs. However, this approach requires a complete overhaul of the existing PKI system. Later work, such as [59], [17], takes a more incremental approach. However, these works' approach still inherits several blockchain-related drawbacks.

The PCB protocol builds on the many works on reliable broadcast and consensus, e.g., [32], [7], [33], [23], particularly recent efficient designs such as [60], [61], [14]. A custom design was necessary to meet CTng's goals, specifically, to support bounded clock drift and communication delay, and to provide proofs of misbehavior as shown in §IV-D.

VIII. Conclusions and Future Work

We presented CTng, an evolutionary extension of CT and the current Web-PKI. CTng is a secure and efficient system that supports certificate transparency and ensures guaranteed, unequivocal revocation. It achieves the NTTP goal that originally motivated CT but has yet to be fully realized. We experimentally validated that a prototype of CTng has low overhead and modest requirements for all participating entities. We analyzed CTng and showed that CTng meets its NTTP-security and systems goals.

We hope this work contributes to the much-needed improvements in the Web-PKI ecosystem, which forms the foundation for applied cryptographic protocols such as TLS, and that it encourages further research in this area. In particular, while we provided intuitive security arguments in §V and in the full version [30], a complete formal proof of security for CTng—similar to the proof presented in [58] for CT/PKIX—remains an important direction for future work. Furthermore, although our protocol does not require real-time signature verification, relying parties must still validate one signature per entity in each MMD. This requirement may pose challenges for resource-constrained clients, especially when post-quantum signature schemes are used, highlighting the importance of exploring alternative designs that reduce relying-party overhead. Finally, our evaluation relies on synthetic data and virtual machines; assessing the performance of CTng in geographically distributed deployments with real network traffic would provide valuable insights into its practicality.

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APPENDIX A
ARTIFACT APPENDIX

A. Abstract

We implemented and evaluated a prototype of the Periodic Consistent Broadcast (PCB) protocol in Go 1.23 [55], performing experiments both with and without erasure encoding. Experiments were conducted on the Sphere Research Infrastructure Testbed [42], using Ubuntu 22.04 virtual machines, each equipped with 8 cores and 16 GB RAM.

B. Description & Requirements

The artifact consists of two components:

1) a minimal example that can be run locally on commodity hardware to verify the program's functionality, and
2) the full experimental setup, which includes all required files and automation scripts for deployment on the Sphere Research Infrastructure testbed [42]. The Sphere setup runs on 50 VMs with a total of 396 CPU cores and 792 GB of memory.

C. How to Access

- The source repository with instructions is available at: https://github.com/jik18001/CTngV3
- A DOI for this release is available at: https://doi.org/10.5281/zenodo.16999030
- Replicating the experiments requires some familiarity with the Sphere testbed. A tutorial is provided at: https://launch.sphere-testbed.net/tutorials

D. Major Claims

Using the code, we have experimentally validated the following claims:

- (C_1) Increasing the number of monitors has a negligible impact on the PCB with the baseline settings (see §A-F).
- (C_2) Applying an erasure encoding algorithm improves overall system performance with higher values of f (see §A-G).
- (C_3) PCB can easily support the current Unique Precertificate Generation rate ($\approx$ 400K precertificates/hr as of Dec 2024[9], see §A-H and §A-I).
- (C_4) PCB scales linearly with increased precertificate workload (see §A-I).

E. Experiment Setup

Before starting each of the large-scale experiments, reset the system to the baseline configuration using the following settings, and then adjust the specific variables required for each experiment:

The experiment is performed on the control node, to access the control node, switch to a non-root user on the XDC and simply run `ssh control`.

- Modify `CTngV3/deter/gen_test.go`:
  - Set `MUD` (= MMD) to 600 (seconds).
  - Set `dmode` to either `def.DEFAULT` or `def.EEA`.
  - Set `num_monitors` to 32.
  - Set `mal` to 3.
  - Set `bmode` to `def.MIN_BC`.
  - Run `go test` under `CTngV3/deter`.
- Modify `CTngV3/deter/CTngexp/inv.ini` to restore the monitor and logger host list back to 32 and 8 respectively.
- Modify `CTngV3/deter/CTngexp/ctngv3.yml`:
  - Set all `async` values to 620 (seconds).
  - Set `remote_user` to your Sphere username
- Modify `CTngV3/deter/CTngexp/redis.yml`:
  - Set `owner` to your Sphere username
  - Set `group` to your Sphere project name
- In `CTngV3/deter/CTngexp`, apply changes across nodes by running:
  - ansible-playbook -i inv.ini redis.yml
- To verify that the changes have been applied to the target hosts (monitors), SSH into any monitor of your choice, navigate to the `CTngV3/deter` folder, and inspect the contents of `detersettings.json` to confirm that the modifications are present.

F. Experiment 1: Varying Number of Monitors

- Repeat the setup as specified in §A-E
- In `CTngV3/deter/gen_test.go`, set `num_monitors` to 8, 16, 24, or 32.
- In `CTngexp/inv.ini`, adjust the monitor host list to match.
- Set `dmode` to the other version and repeat the process above.

Despite increasing the number of monitors, the convergence times remain roughly the same in all trials, proving (C_1).

G. Experiment 2: Different Number of Monitor Faults Allowed

- Repeat the setup as specified in §A-E
- In `CTngV3/deter/gen_test.go`, set `mal` from 1 to 8.
- Set `dmode` to the other version and repeat the process above.

To verify (C_2), we only need to set `mal` to a relatively large value (e.g., `mal` = 7 or 8) and compare the results of `dmode = def.DEFAULT` with those of `dmode = def.EEA`.

H. Experiment 3: Different number of Loggers

- Repeat the setup as specified in §A-E
- In `CTngV3/deter/gen_test.go`, set `num_loggers` to 2, 4, 6, or 8.
- In the same file, set `Certificate_Per_logger` to 33332 (2 loggers), 16666 (4 loggers), 11111 (6 loggers), or 8333 (8 loggers).

- In CTngexp/inv.ini, update the logger host list accordingly.
- Set dmode to the other version and repeat the process above.

In this experiment, we vary the number of loggers while keeping the total workload fixed at 400K precertificates/hr (putting more pressure on fewer loggers). Regardless of the number of loggers used, the convergence time of CTng remains much smaller than the MMD, supporting (C_3).

I. Experiment 4: Different total precertifiate workload.

- Repeat the setup as specified in §A-E
- In CTngV3/deter/gen_test.go, set Certificate_Per_logger from to 4167, 8333, 12500,16666, 20834 and 25000 respectively. Corresponding total workload (in certs/hr) can be computed as:

$$\text{MMD/hr} \times \text{num\_logger} \times \text{cert/logger/MMD}$$

- Example:

$$6 \text{ MMD/hr} \times 8 \text{ loggers} \times 25000 \text{ cert/logger/MMD}$$
$$= 1200K \text{ certs/hr}$$

- Set dmode to the other version and repeat the process above.

Here, we keep 8 loggers but vary the per-logger workload. For all Certificate_Per_logger greater than 8333, the total workload will be greater than 400K precertificates/hr. Despite the greater than average workload, CTng still converges well before the MMD elapses, further evidencing (C_3). By plotting the results for different Certificate_Per_logger, we can easily verify the linearity as specified in (C_4).

J. Result

The *maximum convergence time* is the longest convergence time of all M_i.json files, where i is the monitor identifier.

- **Sample File:** https://github.com/jik18001/CTngV3/blob/main/M1.json
- **Complete Results:** Available in the paper §4 and at: https://github.com/jik18001/CTngV3?tab=readme-ov-file#results