How to Garble Mixed Circuits that Combine Boolean and Arithmetic Computations

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Abstract. The study of garbling arithmetic circuits is initiated by Applebaum, Ishai, and Kushilevitz [FOCS'11], which can be naturally extended to mixed circuits. The basis of mixed circuits includes Boolean operations, arithmetic operations over a large ring and bit-decomposition that converts an arithmetic value to its bit representation. We construct efficient garbling schemes for mixed circuits.

In the random oracle model, we construct two garbling schemes:

- The first scheme targets mixed circuits modulo some $N \approx 2^b$. Addition gates are free. Each multiplication gate costs $O(\lambda \cdot b^{1.5})$ communication. Each bit-decomposition costs $O(\lambda \cdot b^2/\log b)$.
- The second scheme targets mixed circuit modulo some $N \approx 2^b$. Each addition gate and multiplication gate costs $O(\lambda \cdot b \cdot \log b/\log\log b)$. Every bit-decomposition costs $O(\lambda \cdot b^2/\log b)$.

Our schemes improve on the work of Ball, Malkin, and Rosulek [CCS'16] in the same model.

Additionally relying on the DCR assumption, we construct in the programmable random oracle model a more efficient garbling scheme targeting mixed circuits over $\mathbb{Z}_{2^b}$, where addition gates are free, and each multiplication or bit-decomposition gate costs $O(\lambda_{\mathrm{DCR}} \cdot b)$ communication. We improve on the recent work of Ball, Li, Lin, and Liu [Eurocrypt'23] which also relies on the DCR assumption.

1 Introduction

Garbled circuit (GC) is introduced in the seminal work of Yao [1], allowing a *garbler* to efficiently transform any boolean circuit $C : \{0,1\}^{n_{\mathrm{in}}} \to \{0,1\}^{n_{\mathrm{out}}}$ into a *garbled circuit* $\tilde{C}$, along with n_{in} keys $\mathsf{K}_1, \ldots, \mathsf{K}_{n_{\mathrm{in}}}$. Each key is a function $\mathsf{K}_i : \{0,1\} \to \{0,1\}^\lambda$, mapping the i-th input bit to a short string. The output of K_i is referred to as the *label* of the i-th input wire. For any (unknown) input x, the garbled circuit $\tilde{C}$ together with input labels $\mathsf{K}_1(x_1), \ldots, \mathsf{K}_{n_{\mathrm{in}}}(x_{n_{\mathrm{in}}})$ reveal $C(x)$ but nothing else about x.

GC was originally motivated by the 2-party secure computation problem. Since then, GC has found applications to a large variety of problems, and is recognized as one of the most successful and fundamental tools in cryptography.

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For practical applications, people care about the efficiency of GC, especially the communication complexity (i.e., bit length of $\tilde{C}$). A considerable amount of works [2–9] have been dedicated to optimize the *concrete* efficiency of Yao's GC construction. In the most recent construction of Rosulek and Roy [9], XOR and NOT gates involves no communication, every fan-in-2 AND gate requires $1.5\lambda + 5$ bits of communication. Despite making concrete analytic improvement, they still largely follow Yao's construction, binding tightly with boolean circuits. The class of arithmetic operations is a featuring example of computations that are expensive to express as boolean circuits.

The Arithmetic Setting. The beautiful work of Applebaum, Ishai, and Kushilevitz [10] initiated the study of garbling arithmetic circuits.

> *Arithmetic GC* over a ring $\mathcal{R}$ is an efficient algorithm that transforms an arithmetic circuit $C : \mathcal{R}^{n_{\text{in}}} \to \mathcal{R}^{n_{\text{out}}}$ into a garbled circuit $\tilde{C}$, along with n_{in} keys $\mathsf{AK}_1, \ldots, \mathsf{AK}_{n_{\text{in}}}$. Each key is an affine function $\mathsf{AK}_i : \mathcal{R} \to \mathcal{R}^\ell$. For any (unknown) input x, the garbled circuit $\tilde{C}$ together with input labels $\mathsf{AK}_1(x_1), \ldots, \mathsf{AK}_{n_{\text{in}}}(x_{n_{\text{in}}})$ reveal $C(x)$ but nothing else about x.

The construction of AIK is a natural generalization of Yao's boolean GC. For each wire, a key $\mathsf{AK} : \mathcal{R} \to \mathcal{R}^\ell$ is sampled. The output of AK is called the label of that wire, whose length is roughly the security parameter. For any arithmetic gate g, say $\mathsf{AK}_1, \mathsf{AK}_2$ are the keys of the two input wires and AK is the key of the output wire, the garbler generates a table Tab of this gate, such that for any (unknown) $x, y \in \mathcal{R}$, the evaluator can compute $\mathsf{AK}(g(x, y))$ from $\mathsf{AK}_1(x), \mathsf{AK}_2(y), \mathsf{Tab}$, while learning no other information.

As observed by [10], to keep the table size for each gate constant, it suffices to construct the so-called *key-extension*[1] gadget. Such a gadget consists of a garbling algorithm and an evaluation algorithm. The garbling algorithm $\mathsf{KE.Garb}$ takes a key AK and a long key AK^{L} as input, samples a key-extension table Tab such that, $\mathsf{AK}(x), \mathsf{Tab}$ reveal $\mathsf{AK}^{\mathsf{L}}(x)$ but nothing else about $x, \mathsf{AK}^{\mathsf{L}}$.

[10] presents two constructions of key-extension gadgets. One relies on Chinese remainder theorem, enables garbling of mod-$p_1 p_2 \ldots p_k$ (the product of distinct small primes) computation. The other is based on LWE, supports bounded integer computation (computation over the integer ring $\mathbb{Z}$ when all intermediate values are guaranteed to be bounded).

Follow-up research has made improvements within this framework. Similar to FreeXOR, [12] allows free garbling of addition gates. In a different frontier, [11] presents a highly efficient arithmetic GC for bounded integer computation based on Paillier encryption. [11] also presents arithmetic GC for $\mathbb{Z}_p$ based on LWE or Paillier. However, free addition is not supported in [11]. The communication complexity of existing arithmetic GC constructions will be discussed in more detail in Sect. 1.1.

[1] This module is called "key shrinking" in [10]. The name "key extension" comes from [11].

Our research proceeds with this line of study within AIK's framework of arithmetic GC. Our starting point is to understand *how to garble mod-2^b arithmetic circuits,* which is not efficiently supported by previous works. In the search for mod-2^b GC, we realize that it is has a few advantage over GC for mod-p or bounded integer computation.

Match Popular Architectures. In most modern architectures, the only natively supported arithmetic operation is over $\mathbb{Z}_{2^b}$. Most existing tools (programing languages, compilers, processors, etc.) are optimized using/targeting the mod-2^b arithmetic operations. This is our initial motivation to construct the mod-2^b GC.

Mixing Boolean and Arithmetic Computation. Mixed circuits combine boolean and arithmetic computations. The basis include boolean gates, arithmetic operations, together with special gates to convert between boolean and arithmetic values: arithmetic-to-boolean conversion (bit-decomposition) and boolean-to-arithmetic conversion (bit-composition). Previous work [11,12] has considered the garbling of mixed circuits. But in their constructions, the cost of garbling bit-decomposition is expensive.

It turns out that our mod-2^b GC naturally supports *efficient* garbling of bit-decomposition and bit-composition. In fact, in our construction, the key-extension gadget is the combination of bit-decomposition and bit-composition. For example, to double the arithmetic key/label length, first bit-decompose it into boolean labels, then use bit-composition twice to obtain a longer label.

Emulate Arithmetic Computation Modulo Any Modulus N. For any constant N, mod-N computations can be efficiently emulated by mod-2^{4b} mixed circuits if $b = \lceil \log N \rceil$. To prove such a statement, it suffices to show, given $0 \le x < N^2$, how to compute $x \bmod N$ using a mod-2^{4b} mixed circuit. One step further, it is also sufficient to compute integer division $\lfloor x/N \rfloor$ using a mod-2^{4b} mixed circuit. By the rather standard multiply-and-shift trick

$$\lfloor x \cdot \lceil 2^{3b}/N \rceil / 2^{3b} \rfloor = \lfloor x/N \rfloor,$$

the quotient can be computed by first multiplying by constant $\lceil 2^{3b}/N \rceil$ then integer division by 2^{3b}. Both operations are efficient in a mod-2^{4b} mixed circuit.

1.1 Our Results

Mixed GC in the Random Oracle Model. Using only random oracle, the state-of-the-art garbling scheme for arithmetic circuit is that of [12]. They rely on Chinese remainder theorem (CRT) to garble an arithmetic circuit modulo $N = p_1, \ldots p_s \approx 2^b$, by equivalently garbling s copies of the circuit, each modulo a small prime p_i. They allow free addition and each multiplication gate costs $O(\lambda b^2 / \log b)$ bits of communication. However, bit-decomposition operation of this scheme is expensive and not explicitly considered in [12].

Our work improves the state-of-the-art in several directions.

Table 1. Comparison between our GC and previous works

		ADD gate table size	MULT gate table size	bit decomposition	ring modulus	assumption besides RO
boolean	naive	λb	λb^2	free	2^b	
boolean	Karatsuba	λb	$\lambda b^{1.58}$ *	free	2^b	
boolean	FFT-based	λb	$\lambda b \log b$ *	free	2^b	
[12]		free	$\lambda b^2/\log b$	expensive †	$N = p_1 p_2 \dots p_s \approx 2^b$	
Ours (Thm. 1)		free	$\lambda b^2/\log b$	$\lambda b^2/\log b$ ‡	$N = p^k \approx 2^b$	
Ours (Lem. 6)		$\lambda b^2/\log b$	$\lambda b^2/\log b$	$\lambda b^2/\log b$ ‡	any $N \approx 2^b$	
Ours (Thm. 2)		free	$\lambda b^{1.5}$	$\lambda b^2/\log b$ ‡	$N = p_1^{k_1} p_2^{k_2} \dots p_s^{k_s} \approx 2^b$	
Ours (Thm. 3)		$\frac{\lambda b \log b}{\log \log b}$	$\frac{\lambda b \log b}{\log \log b}$	$\lambda b^2/\log b$ ‡	$N = p_1^{k_1} p_2^{k_2} \dots p_s^{k_s} \approx 2^b$	
[11]		$\lambda_{\mathsf{LWE}} b$	$\lambda_{\mathsf{LWE}} b$	unknown	any $N \approx 2^b$	LWE
[11]		$\lambda(\lambda_{\mathsf{DCR}} + b)$	$\lambda(\lambda_{\mathsf{DCR}} + b)$	unknown	any $N \approx 2^b$	strong DCR §
[11]		$\lambda_{\mathsf{LWE}} b$	$\lambda_{\mathsf{LWE}} b$	$\lambda_{\mathsf{LWE}} b^2$	bounded integer	LWE
[11]		$\lambda_{\mathsf{DCR}} + b$	$\lambda_{\mathsf{DCR}} + b$	$\lambda(\lambda_{\mathsf{DCR}} + b)^2$	bounded integer	strong DCR §
Ours (Thm. 4)		free	$\lambda_{\mathsf{DCR}} b$	$\lambda_{\mathsf{DCR}} b$	2^b	DCR
Ours (Cor. 1)		$\lambda_{\mathsf{DCR}} b$	$\lambda_{\mathsf{DCR}} b$	$\lambda_{\mathsf{DCR}} b$	any $N \approx 2^b$	DCR

Constant and $\log(\lambda)$ multiplicative factors are ignored. λ_{LWE} and λ_{DCR} denote the LWE dimension and DCR key length respectively. *Due to large hidden constants, the Karatsuba's method outperforms the naive method only when b is at least a few hundreds, the FFT-base method outperforms Karatsuba's only when b is at least tens of thousands. †The cost is not explicitly stated in [12], but is no less the cost of comparison gate, which is stated to be $O(\lambda b^3/\log b)$. ‡The cost is measured when decomposing to base-p bit representation for some prime p (See Eq. 1). The cost increases to $O(\lambda b^2)$ when decomposing to base-2 bit representation. §Under the standard DCR assumption, "λ" should be replaced by "λ_{DCR}" in its cost expression.

- Our first scheme (Theorem 1) garbles arithmetic gates modulo $N = p^k \approx 2^b$, for some prime p, with the same asymptotic efficiency as [12]: addition is free, each multiplication costs $O(\lambda b^2/\log b)$ bits of communication. Additionally, our scheme supports efficient bit-decomposition gates at a cost of $O(\lambda b^2/\log b)$ communication, enabling the garbling of mixed circuits.
- Our second scheme (Theorem 2) applies CRT in a similar way to [12]. When garbling computations modulo $N = p_1^{k_1} p_2^{k_2} \dots p_s^{k_s} \approx 2^b$, our mixed GC supports free addition and relatively efficient bit-decomposition, and garbles every multiplication gate using $O(\lambda b^{1.5})$ communication.
- Our third scheme (Theorem 3) further improves the multiplication gate cost to $O(\lambda b \log b/\log\log b)$. However, as a trade-off, addition gates are no longer free and have the same cost as multiplication gates.

Mixed GC Based on Computational Assumptions. If allowed to use public key assumptions, the state-of-the-art garbling schemes for arithmetic circuits and mixed circuits are those of [11]. Under the decisional composite residuosity (DCR) assumption, they construct a garbling scheme for bounded integers where

each multiplication gate only costs $O(\lambda_{\mathsf{DCR}} + b)$. In their scheme, the addition gates cost the same as multiplication, the bit-decomposition gates have a more expensive cost of $O(\lambda_{\mathsf{DCR}}^2 \cdot b)$.

Our work improves the state-of-the-art by supporting free addition gates and more efficient bit-decomposition gates. However, as a trade-off, multiplication gates are more expensive, of size $O(\lambda_{\mathsf{DCR}} \cdot b)$.

– Our fourth scheme (Theorem 4) garbles mixed circuits modulo 2^b and allows free addition. Each multiplication gate and bit-decomposition gate costs $O(\lambda_{\mathsf{DCR}} \cdot b)$ communication.

2 Preliminaries

For any positive integer N, let $[N] := \{0, 1, \ldots, N-1\}$, let $\mathbb{Z}_N$ denote the ring of integer modulo N. We assume modulo operation has lower priority than addition. That is, $a + b \bmod p$ should be interpreted as $(a + b) \bmod p$.

Base-p Digit Representation and Bit Representation. For any $x \in [2^b]$, the *bit representation* of x is the unique boolean vector $(x_0, \ldots, x_{b-1}) \in \{0,1\}^n$ such that $x = \sum_i 2^i x_i$. For any $x \in [p^k]$, the *base-p digit representation* of x, is the unique vector $(x_0, \ldots, x_{k-1}) \in [p]^k$ such that $x = \sum_i p^i x_i$.

For any $x \in [p^k]$, let $(x_0, \ldots, x_{k-1})$ be its base-p digit decomposition, the *base-p bit representation* of x is the unique vector $(x_{i,j})_{i \in [k], j \in [\log p]} \in \{0,1\}^{k \cdot \lceil \log p \rceil}$ such that $x_i = \sum_j p^i 2^j x_{i,j}$ for all $i \in [k]$. As a consequence, $x = \sum_{i,j} p^i 2^j x_{i,j}$. That is, base-$p$ bit representation is the bit representation of the base-p digit decomposition.

2.1 Computation Models

We consider *arithmetic circuits* and its generalization *mixed circuits*, where the computation can switch between arithmetic and boolean. Each wire carries a value x in either the boolean field $\mathbb{F}_2 = \{0,1\}$ or an arithmetic ring $\mathcal{R}$. We mainly consider $\mathcal{R} = \mathbb{Z}_{p^k}$ the ring of integer modulo a prime power, and the special case $\mathcal{R} = \mathbb{Z}_{2^b}$. More specifically, we mostly focus on the following class of circuits.

Mixed Circuit. Let $\mathcal{C}_{\mathsf{mix}}(\mathcal{R})$ denote the class of circuits that mixes boolean gates and arithmetic operations over $\mathcal{R}$. A circuit in this class computes a function $f : \{0,1\}^{n_{\text{in,bool}}} \times \mathcal{R}^{n_{\text{in,arith}}} \to \{0,1\}^{n_{\text{out,bool}}} \times \mathcal{R}^{n_{\text{out,arith}}}$ using the gates as basis:

– $\mathsf{Add}, \mathsf{Mult} : \mathcal{R} \times \mathcal{R} \to \mathcal{R}$ compute addition and multiplication over $\mathcal{R}$.
– Bit-decomposition $\mathsf{BD} : \mathcal{R} \to \{0,1\}^b$ computes the bit representation of an arithmetic value.
When $\mathcal{R} = \mathbb{Z}_{2^b}$, we consider the most natural bit decomposition. That is, $\mathsf{BD}(x) = (x_0, x_1, \ldots, x_{b-1})$ such that $x = \sum_i 2^i x_i$.

When $\mathcal{R} = \mathbb{Z}_{p^k}$, the gate first decomposes the number into digits in base p, then decomposes each digit into bits. That is, $b = k \cdot \lceil \log p \rceil$, and

$$\mathsf{BD}(x) = (x_{i,j})_{i\in[k],j\in\lceil \log p\rceil} \quad \text{s.t. } x = \sum_i p^i \sum_j 2^j x_{i,j}. \tag{1}$$

- Bit-composition $\mathsf{BC} : \{0,1\}^b \to \mathcal{R}$ computes the arithmetic value from its bit representation.
- $g : \{0,1\} \times \{0,1\} \to \{0,1\}$ computes the boolean function g.

Arithmetic Circuit. Let $\mathcal{C}_{\mathsf{arith}}(\mathcal{R})$ denote the class of arithmetic circuits over $\mathcal{R}$. A circuit in this class computes a function $f : \mathcal{R}^{n_{\text{in}}} \to \mathcal{R}^{n_{\text{out}}}$ using the following the gates as basis:

- $\mathsf{Add}, \mathsf{Mult} : \mathcal{R} \times \mathcal{R} \to \mathcal{R}$ compute addition and multiplication over $\mathcal{R}$.

2.2 Garbled Circuits (GC)

The following definition of garbling mixed circuits has been implicitly considered in the previous works. We will not separately define arithmetic GC since it can be viewed as the special case of mixed GC.

Definition 1 (Garbling of Mixed Circuits). *A garbling scheme for $\mathcal{C}_{\mathsf{mix}}(\mathcal{R})$ consists of three efficient algorithms.*

- $\mathsf{KeyGen}(1^\lambda, 1^{n_{in,bool}}, 1^{n_{in,arith}})$ *samples $n_{in,bool}$ boolean wire keys $\mathsf{K}_1, \ldots, \mathsf{K}_{n_{in,bool}}$, $n_{in,arith}$ arithmetic wire keys $\mathsf{AK}_1, \ldots, \mathsf{AK}_{n_{in,arith}}$ and status st. Each boolean key K_i is a function from a bit to a bit string. Each arithmetic key AK_i is an affine function from a ring element to a vector.*
- $\mathsf{Garb}(C, \mathsf{st})$ *takes a mixed circuit $C \in \mathcal{C}_{\mathsf{mix}}(\mathcal{R})$, outputs a garbled circuit $\widetilde{C}$.*
- $\mathsf{Eval}(\widetilde{C}, \{\mathbf{l}_i\}_{i\in[n_{in,bool}]}, \{\mathbf{L}_i\}_{i\in[n_{in,arith}]})$ *takes a garbled circuit $\widetilde{C}$, boolean labels $\mathbf{l}_i$, and arithmetic labels $\mathbf{L}_i$. It outputs the evaluation results $\{y_{\text{bool},i}\}$, $\{y_{\text{arith},i}\}$.*

Correctness. The garbling scheme is correct, if for any circuit $C \in \mathcal{C}_{\mathsf{mix}}(\mathcal{R})$ and any input x, as long as $\widetilde{C}$ and keys $\mathsf{K}_1, \ldots, \mathsf{K}_{n_{\text{in,bool}}}, \mathsf{AK}_1, \ldots, \mathsf{AK}_{n_{\text{in,arith}}}$ are properly generated,

$$\mathsf{Eval}(\widetilde{C}, \mathbf{l}_1, \ldots, \mathbf{l}_{n_{\text{in,bool}}}, \mathbf{L}_1, \ldots, \mathbf{L}_{n_{\text{in,arith}}})$$

always outputs $C(x)$, where $\mathbf{l}_i := \mathsf{K}_i(x_{\text{bool},i})$, $\mathbf{L}_i := \mathsf{AK}_i(x_{\text{arith},i})$ are input labels.

Security. The garbling scheme is *secure* if there exists an efficient simulator Sim such that for any circuit $C \in \mathcal{C}_{\mathsf{mix}}(\mathcal{R})$ and input x, the output of $\mathsf{Sim}(C, C(x))$ is indistinguishable from

$$(\widetilde{C}, \mathbf{l}_1, \ldots, \mathbf{l}_{n_{\text{in,bool}}}, \mathbf{L}_1, \ldots, \mathbf{L}_{n_{\text{in,arith}}})$$

when $\tilde{C}, \mathsf{K}_1, \ldots, \mathsf{K}_{n_{\text{in,bool}}}, \mathsf{AK}_1, \ldots, \mathsf{AK}_{n_{\text{in,arith}}}$ are properly generated from C, and $\mathbf{l}_i := \mathsf{K}_i(x_{\text{bool},i})$, $\mathbf{L}_i := \mathsf{AK}_i(x_{\text{arith},i})$.

Gate Gadgets. The construction is mostly modular. For each gate in the basis, there is a garbling gadget for all the tasks related to this gate. Consider a general gate $g : \mathcal{R}_1^n \to \mathcal{R}_2^m$ where $\mathcal{R}_1, \mathcal{R}_2 \in \{\mathbb{Z}_2, \mathcal{R}\}$. The garbling gadget for g consists of three efficient algorithms $g.\mathsf{Garb}, g.\mathsf{Eval}, g.\mathsf{Sim}$. The garbling algorithm $g.\mathsf{Garb}$ takes input wire keys $\mathsf{K}_1, \ldots, \mathsf{K}_n$ (which are boolean keys if $\mathcal{R}_1 = \mathbb{F}_2$, arithmetic keys if $\mathcal{R}_1 = \mathcal{R}$) and output wire keys $\mathsf{K}'_1, \ldots, \mathsf{K}'_m$, generates a table Tab, such that:

- *Correctness.* For any $x_1, \ldots, x_n \in \mathcal{R}_1$ and $(y_1, \ldots, y_m) = g(x_1, \ldots, x_n)$, the evaluation algorithm $g.\mathsf{Eval}(\mathsf{K}_1(x_1), \ldots, \mathsf{K}_n(x_n), \mathsf{Tab})$ will always output $(\mathsf{K}'_1(y_1), \ldots, \mathsf{K}'_m(y_m))$.
- *Handwavy Security.* For any $x_1, \ldots, x_n \in \mathcal{R}_1$, the distribution of Tab is indistinguishable from $g.\mathsf{Sim}(\mathsf{K}_1(x_1), \ldots, \mathsf{K}_n(x_n), \mathsf{K}'_1(y_1), \ldots, \mathsf{K}'_m(y_m))$ when $\mathsf{K}_1(x_1), \ldots, \mathsf{K}_n(x_n)$ are also given to the distinguisher.

As the name suggested, this security definition is imprecise. The issue is mainly caused by the global key. It can be formalized by a global simulator. The global simulator first samples a label for each wire, then samples the garbling table of each gate using the simulation algorithm of the corresponding gadget. In short, the simulation is modular, but the actual security definition is global. For simplicity, we will work in the random oracle model.[2]

There is also a modular approach [10,11] that allows the precise security definition of each gate garbling gadget, but it is incompatible with the existence of the global key. The modular approach requires the simulation algorithm of the gate gadget to sample labels on the input wires. This causes another issue that a label can not be reused by multiple gates. Thus extra work is required when a gate has fan-out greater than 1.

3 Technical Overview

This section briefly discusses AIK's framework of arithmetic GC (Sect. 3.1) and a technically less interesting extension (Sect. 3.2) discussing the sufficiency of bit-decomposition and bit-composition. The takeaway is: Mixed circuits can be efficiently garbled, as long as there are efficient garbling gadgets for bit-decomposition and bit-composition.

In Sect. 3.3, we presents a naive construction of the two garbling gadgets. The resulting GC does not have superior efficiency, but it is simple enough and will be optimized in later sections.

3.1 Background: Key-Extension Implies Arithmetic GC

We recap the framework of AIK [10] for arithmetic GC over some ring $\mathcal{R}$, with the modification that there is a global key $\boldsymbol{\Delta}$ for all arithmetic wires. As observed

[2] In the boolean GC setting, [13] shows how random oracle can be replaced with symmetric encryption resisting a combined related-key and key-dependent message attack. Their technique are likely to work in the arithmetic GC setting as well.

by FreeXOR [4] and "FreeADD" [12], the garbling of addition gates will cost no communication if a global key is sampled.

In more detail, an arithmetic key is sampled for each wire as follows (where λ denotes the security parameter):

- A global key $\mathbf{\Delta} \in \mathcal{R}^\ell$ is sampled for all arithmetic wires, where ℓ is the label length. If $\mathcal{R} = \mathbb{Z}_{2^b}$, we will set $\ell = \lambda$. If $\mathcal{R} = \mathbb{Z}_{p^k}$, we will set $\ell = \lceil \lambda / \log p \rceil$. For each arithmetic wire, the key is an affine function $\mathsf{AK} : \mathcal{R} \to \mathcal{R}^{\ell+1}$. The output $\mathsf{AK}(x)$ consists of ℓ-dimension label and a *color number*. That is, AK can be represented by $\mathsf{AK} = (\mathbf{A} \in \mathcal{R}^\ell, \alpha \in \mathcal{R})$ such that

$$\mathsf{AK}(x) = (\mathbf{\Delta} x + \mathbf{A}, x + \alpha) \quad (\text{in } \mathcal{R}).$$

  α is called the mask number of this wire. Set $\alpha = 0$ for every output wire.

The circuit is garbled gate-by-gate. The garbling gadget for arithmetic gate g consists of a garbling algorithm $g.\mathsf{Garb}$, an evaluation algorithm $g.\mathsf{Eval}$ and a simulation algorithm $g.\mathsf{Sim}$. The garbling algorithm $g.\mathsf{Garb}$ takes the keys of input wires $\mathsf{AK}_1, \mathsf{AK}_2$ and a key of output wire AK, outputs a table Tab such that:

- *Correctness.* For any $x, y \in \mathcal{R}$, $g.\mathsf{Eval}(\mathsf{AK}_1(x), \mathsf{AK}_2(y), \mathsf{Tab}) = \mathsf{AK}(g(x, y))$.
- *Handwavy Security.* For any $x, y \in \mathcal{R}$, the distribution of Tab is indistinguishable from $g.\mathsf{Sim}(\mathsf{AK}_1(x), \mathsf{AK}_2(y), \mathsf{AK}(g(x, y)))$ when $\mathsf{AK}_1(x), \mathsf{AK}_2(y)$ are also given to the distinguisher but the global arithmetic key $\mathbf{\Delta}$ is hidden.

If g is addition, note that

$$\begin{aligned}&\mathsf{AK}_1(x) + \mathsf{AK}_2(y) - \mathsf{AK}(x + y)\\ &= (\mathbf{\Delta} x + \mathbf{A}_1, x + \alpha_1) + (\mathbf{\Delta} y + \mathbf{A}_2, y + \alpha_2) - (\mathbf{\Delta}(x + y) + \mathbf{A}, x + y + \alpha)\\ &= (\mathbf{A}_1 + \mathbf{A}_2 - \mathbf{A}, \alpha_1 + \alpha_2 - \alpha) \quad (\text{in } \mathbb{Z}_{2^d})\end{aligned}$$

can be determined by the input/output labels. Setting it as the table will not violate security and is sufficient for correctness. A smarter solution, as suggested by [12], is to set the table Tab to be *empty*, and to change how the output wire key AK is generated. Instead of sampling AK at random, set $\mathbf{A} = \mathbf{A}_1 + \mathbf{A}_2$ and $\alpha = \alpha_1 + \alpha_2$, thus $\mathsf{AK}_1(x) + \mathsf{AK}_2(y) \bmod 2^d = \mathsf{AK}(x + y)$.

If g is multiplication, first use *randomized encoding* [14,15] to sample two affine functions (long keys) $\mathsf{AK}^{\mathsf{L}}_1, \mathsf{AK}^{\mathsf{L}}_2$ such that $\mathsf{AK}^{\mathsf{L}}_1(x), \mathsf{AK}^{\mathsf{L}}_2(y)$ reveals $\mathsf{AK}(xy)$ but nothing else about x, y, AK. This is formalized as a so-called *affinization gadget* in [10] (called "arithmetic operation gadgets" in [11]).

The affinization gadget for multiplication can be formalized by a garbling algorithm $\mathsf{Aff}_\times.\mathsf{Garb}$, an evaluation algorithm $\mathsf{Aff}_\times.\mathsf{Eval}$ and a simulation algorithm $\mathsf{Aff}_\times.\mathsf{Sim}$.

- Given an affine function, the garbling algorithm $\mathsf{Aff}_\times.\mathsf{Garb}(\mathsf{AK})$ samples two affine functions $\mathsf{AK}^{\mathsf{L}}_1, \mathsf{AK}^{\mathsf{L}}_2$ such that the output dimension of $\mathsf{AK}^{\mathsf{L}}_i$ is at most twice the output dimension of AK. (The multiplicative factors of $\mathsf{AK}^{\mathsf{L}}_1, \mathsf{AK}^{\mathsf{L}}_2$ are not necessarily the global $\mathbf{\Delta}$. We represent a "long key" as $\mathsf{AK}^{\mathsf{L}} = (\mathbf{A}, \mathbf{B})$ such that $\mathsf{AK}^{\mathsf{L}}(x) = \mathbf{A}x + \mathbf{B}$.)

- *Correctness.* For any x, y in the ring, given "long labels", the evaluation algorithm $\mathsf{Aff}_\times.\mathsf{Eval}(\mathsf{AK}_1^\mathsf{L}(x), \mathsf{AK}_2^\mathsf{L}(y))$ always outputs $\mathsf{AK}(xy)$.
- *Security.* For any AK, x, y, the distribution of $(\mathsf{AK}_1^\mathsf{L}(x), \mathsf{AK}_2^\mathsf{L}(y))$ is perfectly indistinguishable from $\mathsf{Aff}_\times.\mathsf{Sim}(\mathsf{AK}(xy))$. The randomness of the former comes from the randomness tape of $\mathsf{Aff}_\times.\mathsf{Garb}$.

The construction of GC is complete by the *key-extension* gadget, which allows the evaluator to compute $\mathsf{AK}_1^\mathsf{L}(x), \mathsf{AK}_2^\mathsf{L}(y)$ from $\mathsf{AK}_1(x), \mathsf{AK}_2(y)$.

The key-extension gadget can be formalized by three efficient algorithms $\mathsf{KE.Garb}$, $\mathsf{KE.Eval}$, $\mathsf{KE.Sim}$.

- Given a key AK and an affine function AK^L, the garbling algorithm $\mathsf{KE.Garb}(\mathsf{AK}, \mathsf{AK}^\mathsf{L})$ samples a table Tab.
- *Correctness.* For any x in the ring, $\mathsf{KE.Eval}(\mathsf{AK}(x), \mathsf{Tab}) = \mathsf{AK}^\mathsf{L}(x)$.
- *Handwavy Security.* For any x, the distribution of Tab is indistinguishable from $\mathsf{KE.Sim}(\mathsf{AK}(x), \mathsf{AK}^\mathsf{L}(x))$ when $\mathsf{AK}(x)$ are also given to the distinguisher but $\boldsymbol{\Delta}$ is hidden.

The garbling gadget for multiplication gates can be constructed as follows.

- Garbling algorithm $\mathsf{Mult.Garb}(\mathsf{AK}_1, \mathsf{AK}_2, \mathsf{AK})$:
  $\mathsf{Aff}_\times.\mathsf{Garb}(\mathsf{AK}) \to (\mathsf{AK}_1^\mathsf{L}, \mathsf{AK}_2^\mathsf{L})$.
  $\mathsf{KE.Garb}(\mathsf{AK}_i, \mathsf{AK}_i^\mathsf{L}) \to \mathsf{Tab}_i$ for $i \in \{1, 2\}$.
  Output $\mathsf{Tab} = (\mathsf{Tab}_1, \mathsf{Tab}_2)$.
- Evaluation algorithm $\mathsf{Mult.Eval}(\mathbf{L}_1, \mathbf{L}_2, \mathsf{Tab})$:
  $\mathsf{KE.Eval}(\mathbf{L}_i, \mathsf{Tab}_i) \to \mathbf{L}_i^\mathsf{L}$ for $i \in \{1, 2\}$.
  $\mathsf{Aff}_\times.\mathsf{Eval}(\mathbf{L}_1^\mathsf{L}, \mathbf{L}_2^\mathsf{L}) \to \mathbf{L}$.
  Output $\mathbf{L}$.
- Simulation algorithm $\mathsf{Mult.Sim}(\mathbf{L}_1, \mathbf{L}_2, \mathbf{L})$:
  $\mathsf{Aff}_\times.\mathsf{Sim}(\mathbf{L}) \to \mathbf{L}_1^\mathsf{L}, \mathbf{L}_2^\mathsf{L}$.
  $\mathsf{KE.Sim}(\mathbf{L}_i, \mathbf{L}_i^\mathsf{L}) \to \mathsf{Tab}_i$ for $i \in \{1, 2\}$.
  Output $\mathsf{Tab} = (\mathsf{Tab}_1, \mathsf{Tab}_2)$.

This arithmetic GC framework [10,12] reduces the problem to constructing a key-extension gadget. As long as there is a secure key-extension gadget that doubles the key length (i.e., the output of AK^L can be twice as long as AK), the framework will yield an arithmetic GC of the same complexity.

Lemma 1 (informal). *If there is a secure key-extension gadget that doubles the key length whose table size is c_{KE}, there is an arithmetic GC for the same ring such that each addition gate costs no communication, and each multiplication gate costs $2 \cdot c_{\mathrm{KE}}$ communication.*

3.2 Bit-Decomposition and Bit-Composition Imply Mixed GC

We extend the AIK framework to support mixed circuit, which consists of arithmetic operation gates as described before, boolean gates such as AND, XOR, and NOT, and two conversion gates, bit-decomposition and bit-compositions.

A wire in the circuit is either an arithmetic wires as described before, or a boolean wire. The keys for arithmetic wires stay unchanged. The keys for boolean wires are sampled as follows:

- A global key $\Delta \in \{0,1\}^\lambda$ is sampled for all boolean wires.
  For each boolean wire, the key is an affine function $\mathsf{K} : \{0,1\} \to \{0,1\}^{\lambda+1}$. The output $\mathsf{K}(x)$ consists of a λ-bit label and a color bit. That is, K can be represented by $\mathsf{K} = (\mathbf{b} \in \{0,1\}^\lambda, \alpha \in \{0,1\})$ such that

$$\mathsf{K}(x) = (\Delta x \oplus \mathbf{b}, x \oplus \alpha).$$

  α is called the mask bit of this wire. Set $\alpha = 0$ for every output wire.

The arithmetic operation gates are garbled as before, and we skip the rather standard boolean gate garbling gadgets. We describe gadgets for garbling bit-decomposition and bit-composition gates in more detail below.

The bit-decomposition gadget consists of $\mathsf{BD.Garb}, \mathsf{BD.Eval}, \mathsf{BD.Sim}$. The garbling algorithm $\mathsf{BD.Garb}$ takes an arithmetic key AK and b boolean keys $\mathsf{K}_0, \ldots, \mathsf{K}_{b-1}$ as inputs, outputs a table Tab, such that

- *Correctness.* For any $x \in \mathcal{R}$, $\mathsf{BD.Eval}(\mathsf{AK}(x), \mathsf{Tab}) = (\mathsf{K}_0(x_0), \ldots, \mathsf{K}_{b-1}(x_{b-1}))$.
- *Handwavy Security.* For any $x \in \mathcal{R}$, the distribution of $\mathsf{AK}(x), \mathsf{Tab}$ is indistinguishable from $\mathsf{AK}(x), \mathsf{BD.Sim}(\mathsf{AK}(x), \mathsf{K}_0(x_0), \ldots, \mathsf{K}_{b-1}(x_{b-1}))$ when the global arithmetic key $\boldsymbol{\Delta}$ is hidden.

The bit-composition gadget consists of $\mathsf{BC.Garb}, \mathsf{BC.Eval}, \mathsf{BC.Sim}$. The garbling algorithm $\mathsf{BC.Garb}$ takes b boolean keys $\mathsf{K}_0, \ldots, \mathsf{K}_{b-1}$ and an arithmetic affine function AK^L as inputs, outputs a table Tab, such that

- *Correctness.* For any $x \in \mathcal{R}$, $\mathsf{BC.Eval}(\mathsf{K}_0(x_0), \ldots, \mathsf{K}_{b-1}(x_{b-1}), \mathsf{Tab}) = \mathsf{AK}^\mathsf{L}(x)$.
- *Handwavy Security.* For any $x \in \mathcal{R}$, the distribution of Tab is indistinguishable from $\mathsf{BC.Sim}(\mathsf{K}_0(x_0), \ldots, \mathsf{K}_{b-1}(x_{b-1}), \mathsf{AK}^\mathsf{L}(x))$ when $\mathsf{K}_0(x_0), \ldots, \mathsf{K}_{b-1}(x_{b-1})$ is also given to the adversary but the global key Δ is hidden.

We stress that AK^L can be an arbitrary affine function: its multiplicative factor does not have to be the global key; and its output dimension can be larger. Although for simplicity, we assume the output dimension of AK^L equals the dimension of a label. In case we need longer AK^L, we can always divide it into a few pieces and use the bit-composition gadget multiple times.

It is obvious that bit-decomposition gadget and bit-composition gadget imply key-extension gadget, and thus imply mixed GC. Previous work did not construct the key-extension gadget through this approach because bit-decomposition is expensive in their constructions.

Lemma 2 (informal). *If there are a secure bit-decomposition gadget whose table size is c_{BD} and a secure bit-composition gadget whose table size is c_{BC}, then there is a mixed GC for the same ring such that each addition gate costs no communication, and each multiplication/bit-decomposition/bit-composition gate costs $O(c_{\mathrm{BD}} + c_{\mathrm{BC}})$ communication.*

3.3 The Naive Construction

This section presents garbling gadgets for bit-decomposition and bit-composition when the ring is $\mathbb{Z}_{2^b}$. For each $x \in \mathbb{Z}_{2^b}$, let x_i denote the i-th lowest bit of x, so that $x = \sum_i 2^i x_i$. Let $x_{a:b}$ denote $\sum_{a \le i < b} 2^{i-a} x_i$, so that the bit representation of $x_{a:b}$ is a substring of the bit representation of x.

BC. The bit-composition gadget is straight-forward. Given boolean input labels $\mathsf{K}_0(x_0), \ldots, \mathsf{K}_{b-1}(x_{b-1})$, the evaluator need to compute the output label $\mathsf{AK}^{\mathsf{L}}(x) = \mathbf{A}x + \mathbf{B}$ (recall that in bit-composition gadget, the output key can be any affine function). The garbling algorithm $\mathsf{BC.Garb}$ samples additive sharing $\mathbf{B}_0, \ldots, \mathbf{B}_{b-1}$ such that $\sum_i \mathbf{B}_i = \mathbf{B}$, then generates table that allows the evaluator to compute $\mathbf{A}2^i x_i + \mathbf{B}_i$ from $\mathsf{K}_i(x_i)$. The most direct solution is to let the table contain ciphertexts

$$\mathsf{Enc}(\mathsf{K}_i(\beta), \mathbf{A}2^i\beta + \mathbf{B}_i) \text{ for all } \beta \in \{0,1\}.$$

The order of the two ciphertexts are permuted according to the mask bit in K_i, so that the evaluator can pick the right ciphertext using the color bit.

BD. The bit decomposition gadget is inspired by the following two observations.

- Let $\mathbf{L} = \mathsf{AK}(x) = \boldsymbol{\Delta}x + \mathbf{A}$ denote the given arithmetic label. Then

  $$\mathbf{L} \bmod 2 = \boldsymbol{\Delta}x + \mathbf{A} \bmod 2 = \boldsymbol{\Delta}x_0 + \mathbf{A} \bmod 2.$$

  If the table contains $\mathsf{Enc}(\boldsymbol{\Delta}\beta + \mathbf{A} \bmod 2, \mathsf{K}_0(\beta))$ for $\beta \in \{0,1\}$, the evaluator can properly decrypt the boolean label $\mathsf{K}_0(x_0)$ of x_0 with $\mathbf{L} \bmod 2$.
- To continue, the evaluator should be able to compute a mod-2^{b-1} arithmetic label for all but the least significant bit of x

  $$\mathbf{L}^{(1)} = \boldsymbol{\Delta}x_{1:b} + \mathbf{A}^{(1)} \bmod 2^{b-1}.$$

  Then the evaluator can iteratively compute all the boolean labels. Note that,

  $$\mathbf{L} - 2\mathbf{L}^{(1)} \bmod 2^b = \boldsymbol{\Delta}x_0 + \mathbf{A} - 2\mathbf{A}^{(1)} \bmod 2^b. \tag{2}$$

  If the table also contains ciphertexts

  $$\mathsf{Enc}(\boldsymbol{\Delta}\beta + \mathbf{A} \bmod 2,\ \boldsymbol{\Delta}\beta + \mathbf{A} - 2\mathbf{A}^{(1)} \bmod 2^b) \text{ for } \beta \in \{0,1\},$$

  the evaluator can decrypt the ciphertext to get (2) and compute $\mathbf{L}^{(1)}$.

These observations lead us to the bit-decomposition gadget in Fig. 1. For simplicity, the encryption is implemented by a secure function H which is modeled as a random oracle

$$\mathsf{Enc}(key, m) = \mathsf{H}(key, \mathsf{aux}) \oplus m, \quad \mathsf{Dec}(key, c) = \mathsf{H}(key, \mathsf{aux}) \oplus c,$$

where aux contains auxiliary information such as the id of current gate. The H queries under some auxiliary information is bounded: For each aux, the construction only queries $\mathsf{H}(key, \mathsf{aux})$ for up to two distinct key.

Garbling algorithm $\mathsf{BD.Garb}$ takes an arithmetic key $\mathsf{AK} = (\mathbf{A}, \alpha)$ and b boolean keys $\mathsf{K}_0, \ldots, \mathsf{K}_{b-1}$ as inputs.

- Let $\mathbf{A}^{(0)} = \mathbf{A}$. For each $1 \le i < b$, samples $\mathbf{A}^{(i)} \leftarrow (\mathbb{Z}_{2^{b-i}})^\lambda$.
- Let $\alpha^{(0)} = \alpha$. For each $1 \le i < b$, samples $\alpha^{(i)} \leftarrow \mathbb{Z}_{2^{b-i}}$.
- For each $0 \le i < b$, for each $\beta \in \{0,1\}$, compute

$$\mathsf{C}_{i,\beta+\alpha^{(i)} \bmod 2} \leftarrow \mathsf{H}(\boldsymbol{\Delta}\beta + \mathbf{A}^{(i)} \bmod 2,\ (\mathsf{id}, i)) \oplus \begin{cases} (\mathsf{K}_i(\beta), \boldsymbol{\Delta}\beta + \mathbf{A}^{(i)} - 2\mathbf{A}^{(i+1)} \bmod 2^{b-i}, \\ \qquad \beta + \alpha^{(i)} - 2\alpha^{(i+1)} \bmod 2^{b-i}) & \text{if } i < b-1 \\ \mathsf{K}_i(\beta), & \text{if } i = b-1 \end{cases}$$

- Output table $\mathsf{Tab} = (\mathsf{C}_{i,\beta})_{i\in[b],\beta\in\{0,1\}}$

Evaluation algorithm $\mathsf{BD.Eval}$ takes input label $(\mathbf{L}, \bar{x})$ and a table Tab as inputs.

- Let $\mathbf{L}^{(0)} := \mathbf{L}$, $\bar{x}^{(0)} = \bar{x}$.
- For $i = 0, 1, 2, \ldots, b-1$:Compute $(\mathbf{l}_i, \mathbf{D}^{(i)}, d^{(i)}) \leftarrow \mathsf{H}(\mathbf{L}^{(i)} \bmod 2,\ (\mathsf{id}, i)) \oplus \mathsf{C}_{i,\bar{x}^{(i)} \bmod 2}$. If $i < b-1$, compute

$$\mathbf{L}^{(i+1)} = (\mathbf{L}^{(i)} - \mathbf{D}^{(i)} \bmod 2^{b-i})/2\,, \qquad \bar{x}^{(i+1)} = (\bar{x}^{(i)} - d^{(i)} \bmod 2^{b-i})/2\,.$$

- Output boolean labels $\mathbf{l}_0, \mathbf{l}_1, \ldots, \mathbf{l}_{b-1}$.

Simulation algorithm $\mathsf{BD.Sim}$ takes arithmetic label $(\mathbf{L}, \bar{x})$ and boolean labels $\mathbf{l}_0, \mathbf{l}_1, \ldots, \mathbf{l}_{b-1}$ as inputs.

- Let $(\mathbf{L}^{(0)}, \bar{x}^{(0)}) = (\mathbf{L}, \bar{x})$.
- Sample random $\mathbf{L}^{(i)} \leftarrow (\mathbb{Z}_{2^{b-i}})^\lambda, \bar{x}^{(i)} \leftarrow \mathbb{Z}_{2^{b-i}}$ for each $1 \le i < b$.
- The active ciphertexts in the table Tab are set as

$$\mathsf{C}_{i,\bar{x}^{(i)} \bmod 2} = \mathsf{H}(\mathbf{L}^{(i)} \bmod 2,\ (\mathsf{id}, i)) \oplus \begin{cases} (\mathbf{l}_i, \mathbf{L}^{(i)} - 2\mathbf{L}^{(i+1)} \bmod 2^{b-i},\ \bar{x}^{(i)} - 2\bar{x}^{(i+1)} \bmod 2^{b-i}) & \text{if } i < b-1 \\ \mathbf{l}_i & \text{if } i = b-1 \end{cases}$$

  The rest are inactive ciphertexts, and are simulated as random strings.

Fig. 1. The Naive Bit-Decomposition Gadget

Lemma 3. *There are statistically secure bit-decomposition gadget (Fig. 1) and bit-composition gadget (a specialization of Fig. 2) for ring $\mathbb{Z}_{2^b}$, whose table size is $O(b^2\lambda)$. They yield statistically secure mixed GC for $\mathbb{Z}_{2^b}$ in the random oracle model, where each addition gate costs no communication, and each multiplication/bit-decomposition/bit-composition gate costs $O(b^2\lambda)$ communication.*

The proof of Lemma 3 is deferred to the full version[3].

[3] https://eprint.iacr.org/2023/1584.

4 Mixed GC for $\mathbb{Z}_{p^k}$

This section presents a mix GC for $\mathbb{Z}_{p^k}$. Recall how the arithmetic key, label, color number are defined for each arithmetic wire (where λ is the security parameter):

- A global key $\mathbf{\Delta} \in \mathbb{Z}_{p^k}^{\ell}$ is sampled for all arithmetic wires, where $\ell = \lceil \lambda / \log p \rceil$ is the label length.
  For each arithmetic wire, the key is an affine function $\mathsf{AK} : \mathbb{Z}_{p^k} \to \mathbb{Z}_{p^k}^{\ell+1}$. The output $\mathsf{AK}(x)$ consists of ℓ-dimension label and a *color number*. That is, AK can be represented by $\mathsf{AK} = (\mathbf{A} \in \mathbb{Z}_{p^k}^{\ell}, \alpha \in \mathbb{Z}_{p^k})$ such that

$$\mathsf{AK}(x) = (\mathbf{\Delta} x + \mathbf{A}, x + \alpha) \mod p^k.$$

  α is called the mask number of this wire. Set $\alpha = 0$ for every output wire.

As discussed in Sect. 3.2, it suffices to construct efficient garbling gadgets for bit-decomposition and bit-composition over ring $\mathbb{Z}_{p^k}$. The construction of the two gadgets for $\mathbb{Z}_{p^k}$ generalizes the constructions for $\mathbb{Z}_{2^b}$ in Sect. 3.3.

For each $x \in \mathbb{Z}_{p^k}$, let x_i denote the i-th lowest digit of x, so that $x = \sum_i p^i x_i$. Let $x_{a:b}$ denote $\sum_{a \le i < b} p^{i-a} x_i$, so that the base-$p$ digit representation of $x_{a:b}$ is a substring of the base-p digit representation of x. Let $x_{i,j}$ denote the j-th lowest bit of x_i, so that $x_i = \sum_j 2^j x_{i,j}$.

For each $\beta \in \mathbb{Z}_p$, let β_i denote the i-th lowest bit of β, so that $\beta = \sum_i 2^i \beta_i$. Let $\beta_{a:b}$ denote $\sum_{a \le i < b} 2^{i-a} \beta_i$, so that the bit representation of $\beta_{a:b}$ is a substring of the bit representation of β.

BC. The bit-composition gadget is straight-forward. Given boolean input labels $\mathsf{K}_{i,j}(x_{i,j})$ for $i \in [k], j \in [\log p]$, the evaluator needs to compute the output label $\mathsf{AK}^{\mathsf{L}}(x) = \mathbf{A}x + \mathbf{B}$ (recall that in the bit-composition gadget, the output key can be any affine function). The garbling algorithm $\mathsf{BC.Garb}$ samples additive sharing $\mathbf{B}_{i,j}$ such that $\sum_{i,j} \mathbf{B}_{i,j} = \mathbf{B}$, then generates a table that allows the evaluator to compute $\mathbf{A}p^i 2^j x_{i,j} + \mathbf{B}_{i,j}$ from $\mathsf{K}_{i,j}(x_{i,j})$. The most direct solution is to let the table contain ciphertexts

$$\mathsf{Enc}(\mathsf{K}_{i,j}(\beta), \mathbf{A}p^i 2^j \beta + \mathbf{B}_{i,j}) \text{ for all } \beta \in \{0, 1\}.$$

The order of the two ciphertexts are permuted according to the mask bit in $\mathsf{K}_{i,j}$, so that the evaluator can pick the right ciphertext according to the color bit.

The construction is formalized in Fig. 2. The table consists of $O(k \log p)$ ciphertexts, each ciphertext is $k\lambda$-bit long, thus the table size is $O(\lambda k^2 \log p)$ bit.

BD. The bit-decomposition gadget starts with the same observations as the one in Sect. 3.3. Let $\mathbf{L} = \mathsf{AK}(x) = \mathbf{\Delta} x + \mathbf{A} \mod p^k$ denote the given arithmetic label. Define

$$\mathbf{L}^{(i)} = \mathbf{\Delta} x_{i:k} + \mathbf{A}^{(i)} \mod p^{k-i}.$$

Garbling algorithm $\mathsf{BC.Garb}$ takes boolean keys $\mathsf{K}_{i,j}$ for $i \in [k], j \in [\log p]$, and an arithmetic key $\mathsf{AK}^{\mathsf{L}} = (\mathbf{A}, \mathbf{B})$ as inputs. Let $\alpha_{i,j}$ denote the mask bit of $\mathsf{K}_{i,j}$.

- Sample random $\mathbf{B}_{i,j}$ for $i \in [k], j \in [\log p]$, satisfying $\sum_{i,j} \mathbf{B}_{i,j} \bmod p^k = \mathbf{B}$.
- For each $i \in [k], j \in [\log p]$, for each $\beta \in \{0, 1\}$, compute

$$\mathsf{C}_{i,j,\beta+\alpha_{i,j} \bmod 2} \leftarrow \mathsf{H}(\mathsf{K}_{i,j}(\beta),\ (\mathsf{id}, i, j)) \oplus (\mathbf{A} p^i 2^j \beta + \mathbf{B}_{i,j} \bmod p^k)$$

- Output table $\mathsf{Tab} = (\mathsf{C}_{i,j,\beta})_{i \in [k], j \in [\log p], \beta \in \{0,1\}}$

Evaluation algorithm $\mathsf{BC.Eval}$ takes input labels $(\mathbf{l}_{i,j}, \bar{x}_{i,j})$ for $i \in [k], j \in [\log p]$ and a table Tab as inputs.

- For $i \in [k], j \in [\log p]$, compute $\mathbf{L}_{i,j} \leftarrow \mathsf{H}(\mathbf{l}_{i,j},\ (\mathsf{id}, i, j)) \oplus \mathsf{C}_{i,j,\bar{x}_{i,j}}$.
- Output arithmetic label $\mathbf{L} = \sum_{i,j} \mathbf{L}_{i,j} \bmod p^k$.

Simulation algorithm $\mathsf{BC.Sim}$ takes input labels $(\mathbf{l}_{i,j}, \bar{x}_{i,j})$ for $i \in [k], j \in [\log p]$ and arithmetic label $\mathbf{L}$ as inputs.

- Sample random $\mathbf{L}_{i,j}$ for $i \in [k], j \in [\log p]$, satisfying $\sum_{i,j} \mathbf{L}_{i,j} \bmod p^k = \mathbf{L}$.
- The active ciphertexts in the table Tab are set as

$$\mathsf{C}_{i,j,\bar{x}_{i,j}} = \mathsf{H}(\mathbf{l}_{i,j},\ (\mathsf{id}, i, j)) \oplus \mathbf{L}_{i,j}$$

  The rest are inactive ciphertexts, and are simulated by random strings.

Fig. 2. The Naive Bit-Composition Gadget

where $\mathbf{A}^{(0)} := \mathbf{A}$ and $\mathbf{A}^{(i)}$ are randomly sampled. Thus, $\mathbf{L}^{(0)} = \mathbf{L}$. Note that,

$$\mathbf{L}^{(i)} \bmod p = \boldsymbol{\Delta} x_{i:k} + \mathbf{A}^{(i)} \bmod p = \boldsymbol{\Delta} x_i + \mathbf{A}^{(i)} \bmod p.$$

If the table contains ciphertext

$$\mathsf{Enc}(\boldsymbol{\Delta} x_i + \mathbf{A}^{(i)} \bmod p,\ (\text{boolean labels of } x_i, \boldsymbol{\Delta} x_i + \mathbf{A}^{(i)} - p\mathbf{A}^{(i+1)} \bmod p^{k-i}))$$

the evaluator can, given $\mathbf{L}^{(i)}$, computes all the boolean labels of x_i and the next label $\mathbf{L}^{(i+1)}$. This observation can be formalized as a secure bit-decomposition gadget, who has poor efficiency. The table consists of pk ciphertexts, each ciphertext is $(\lambda \log p + \lambda k)$-bit long, the total length is no less than $\lambda p k^2$. Under constraint $p^k \approx 2^b$, the table size is minimized when $p = O(1)$, which is asymptotically equivalent to the naive construction in Sect. 3.3.

The bottleneck is the encryption of $\boldsymbol{\Delta} x_i + \mathbf{A}^{(i)} - p\mathbf{A}^{(i+1)} \bmod p^{k-i}$. To optimize the efficiency, we replace the long ciphertexts by shorter ciphertexts

$$\mathsf{Enc}(\boldsymbol{\Delta} x_i + \mathbf{A}^{(i)} \bmod p,\ \text{boolean labels of } x_i)$$

that only encrypts the boolean labels. Since the evaluator can computes the boolean labels of x_i, it uses a mini bit-composition gadget (Fig. 3) to compute $\boldsymbol{\Delta} x_i + \mathbf{A}^{(i)} - p\mathbf{A}^{(i+1)} \bmod p^{k-i}$.

The optimized construction is formalized in Fig. 4. After optimization, the table consists of $O(kp)$ ciphertexts, each of which is $O(\lambda \log p)$ bit long, and k mini-tables for the mini bit-composition, each of which is $O(\lambda k \log p)$ bit long. The total table size is $O(\lambda k(k+p)\log p)$.

Garbling algorithm $\mathsf{miniBC}_k.\mathsf{Garb}$ takes boolean keys K_j for $j \in [\log p]$, and an arithmetic key $\mathsf{AK}^{\mathsf{L}} = (\mathbf{A}, \mathbf{B})$ as inputs. Let α_j denote the mask bit of K_j.

- Sample random $\mathbf{B}_j$ for $j \in [\log p]$, satisfying $\sum_j \mathbf{B}_j \bmod p^k = \mathbf{B}$.
- For each $j \in [\log p]$, for each $\beta \in \{0,1\}$, compute

$$\mathsf{C}_{j,\beta+\alpha_j \bmod 2} \leftarrow \mathsf{H}(\mathsf{K}_j(\beta),\ (\mathsf{id}, j)) \oplus (\mathbf{A}2^j\beta + \mathbf{B}_j \bmod p^k)$$

- Output table $\mathsf{Tab} = (\mathsf{C}_{j,\beta})_{j\in[\log p],\beta\in\{0,1\}}$

Evaluation algorithm $\mathsf{miniBC}_k.\mathsf{Eval}$ takes input labels $(\mathbf{l}_j, \bar{x}_j)$ for $j \in [\log p]$ and a table Tab as inputs.

- For $j \in [\log p]$, compute $\mathbf{L}_j \leftarrow \mathsf{H}(\mathbf{l}_j,\ (\mathsf{id}, j)) \oplus \mathsf{C}_{j,\bar{x}_j}$.
- Output arithmetic label $\mathbf{L} = \sum_j \mathbf{L}_j \bmod p^k$.

Simulation algorithm $\mathsf{miniBC}_k.\mathsf{Sim}$ takes input labels $(\mathbf{l}_j, \bar{x}_j)$ for $j \in [\log p]$ and arithmetic label $\mathbf{L}$ as inputs.

- Sample random $\mathbf{L}_j$ for $j \in [\log p]$, satisfying $\sum_j \mathbf{L}_j \bmod p^k = \mathbf{L}$.
- The active ciphertexts in the table Tab are set as

$$\mathsf{C}_{j,\bar{x}_j} = \mathsf{H}(\mathbf{l}_j,\ (\mathsf{id}, j)) \oplus \mathbf{L}_j$$

  The rest are simulated by random strings.

Fig. 3. The Mini Bit-Composition Gadget

Theorem 1. *There are statistically secure bit-composition gadget (Fig. 2) for ring $\mathbb{Z}_{p^k}$ whose table size is $O(\lambda k^2 \log p)$ and bit-decomposition gadget (Fig. 4) for ring $\mathbb{Z}_{p^k}$, whose table size is $O(\lambda k(k+p)\log p)$. They yield a statistically secure mixed GC for $\mathbb{Z}_{p^k}$ in the random oracle model, such that each addition gate costs no communication, and each multiplication/bit-decomposition/bit-composition gate costs $O(\lambda k(k+p)\log p)$ communication.*

The bit-composition gadget (Fig. 2) and the mini bit-composition gadget (Fig. 3) are special cases of the linear bit-composition gadget (Fig. 5), whose correctness and security will be analyzed in Sect. 4.1. The proof of the bit-decomposition gadget is similar to that of Lemma 3 in Sect. 3.3.

Under the constraint that $p^k \approx 2^b$, the asymptotic cost per gate is minimized when $p \approx b/\log^c b$ for any constant $c \geq 1$. The minimal cost is $O(\lambda b^2/\log b)$.

Garbling algorithm $\mathsf{BD.Garb}$ takes an arithmetic key $\mathsf{AK} = (\mathbf{A}, \alpha)$ and $k \cdot \lceil \log p \rceil$ boolean keys $\mathsf{K}_{i,j}$ for $i \in [p], j \in [\log p]$ as inputs.

- Let $\mathbf{A}^{(0)} = \mathbf{A}$. For each $1 \le i < k$, samples $\mathbf{A}^{(i)} \leftarrow (\mathbb{Z}_{p^{k-i}})^{\lambda}$.
- Let $\alpha^{(0)} = \alpha$. For each $1 \le i < k$, samples $\alpha^{(i)} \leftarrow \mathbb{Z}_{p^{k-i}}$.
- For each $i \in [k], j \in [\log p]$, for each $\beta \in [p]$, compute

$$\mathsf{C}_{i,\beta+\alpha^{(i)} \bmod p} \leftarrow \mathsf{H}(\boldsymbol{\Delta}\beta + \mathbf{A}^{(i)} \bmod p,\ (\mathsf{id}, i)) \oplus (\mathsf{K}_{i,j}(\beta_j) \text{ for } j \in [\log p])$$

- For each $0 \le i < k-1$, define affine function $\mathsf{DK}^{(i)}$

$$\mathsf{DK}^{(i)}(\beta) = (\boldsymbol{\Delta}\beta + \mathbf{A}^{(i)} - p\mathbf{A}^{(i+1)},\ \beta + \alpha^{(i)} - p\alpha^{(i+1)}) \bmod p^{k-i},$$

  compute table $\mathsf{tb}_i \leftarrow \mathsf{miniBC}_{k-i}.\mathsf{Garb}(\mathsf{K}_{i,j} \text{ for } j \in [\log p], \mathsf{DK}^{(i)})$.
- Output table Tab consisting of $(\mathsf{C}_{i,\beta})_{i\in[k],\beta\in[p]}$ and $(\mathsf{tb}_i)_{i\in[k-1]}$

Evaluation algorithm $\mathsf{BD.Eval}$ takes input label $(\mathbf{L}, \bar{x})$ and a table Tab as inputs.

- Let $\mathbf{L}^{(0)} := \mathbf{L}$, $\bar{x}^{(0)} = \bar{x}$.
- For $i = 0, 1, 2, \ldots, k-1$:
  Compute $(\mathbf{l}_{i,j} \text{ for } j \in [\log p]) \leftarrow \mathsf{H}(\mathbf{L}^{(i)} \bmod p,\ (\mathsf{id}, i)) \oplus \mathsf{C}_{i,\bar{x}^{(i)} \bmod p}$.
  If $i < k-1$, compute $(\mathbf{D}^{(i)}, d^{(i)}) \leftarrow \mathsf{miniBC}_{k-i}.\mathsf{Eval}(\mathbf{l}_{i,j} \text{ for } j \in [\log p], \mathsf{tb}_i)$

$$\mathbf{L}^{(i+1)} = (\mathbf{L}^{(i)} - \mathbf{D}^{(i)} \bmod p^{k-i})/p\,, \qquad \bar{x}^{(i+1)} = (\bar{x}^{(i)} - d^{(i)} \bmod p^{k-i})/p\,.$$

- Output boolean labels $\mathbf{l}_{i,j}$ for $i \in [p], j \in [\log p]$.

Simulation algorithm $\mathsf{BD.Sim}$ takes arithmetic label $(\mathbf{L}, \bar{x})$ and boolean labels $\mathbf{l}_{i,j}$ for $i \in [p], j \in [\log p]$ as inputs.

- Let $(\mathbf{L}^{(0)}, \bar{x}^{(0)}) = (\mathbf{L}, \bar{x})$.
- Sample random $\mathbf{L}^{(i)} \leftarrow (\mathbb{Z}_{p^{k-i}})^{\lambda}, \bar{x}^{(i)} \leftarrow \mathbb{Z}_{p^{k-i}}$ for each $1 \le i < k$.
- The active ciphertexts in the table Tab are set as

$$\mathsf{C}_{i,\bar{x}^{(i)} \bmod p} = \mathsf{H}(\mathbf{L}^{(i)} \bmod p,\ (\mathsf{id}, i)) \oplus (\mathbf{l}_{i,j} \text{ for } j \in [\log p])$$

  The rest are inactive ciphertexts, and are simulated by random strings.
- For each $0 \le i < k-1$, compute

$$(\mathbf{D}^{(i)}, d^{(i)}) \leftarrow (\mathbf{L}^{(i)} - p\mathbf{L}^{(i+1)},\ \bar{x}^{(i)} - p\bar{x}^{(i+1)}) \bmod p^{k-i}$$

  and simulate tb_i by $\mathsf{tb}_i \leftarrow \mathsf{miniBC}_{k-i}.\mathsf{Sim}(\mathbf{l}_{i,j} \text{ for } j \in [\log p], (\mathbf{D}^{(i)}, d^{(i)}))$.

Fig. 4. The Bit-Decomposition Gadget in Ring $\mathbb{Z}_{p^k}$

Further Optimization. The bit-decomposition gadget in Fig. 4 can be further optimized. Currently, for each $i \in [k]$ the table contains ciphertexts

$$\mathsf{C}_{i,\beta+\alpha^{(i)} \bmod p} \leftarrow \mathsf{H}(\boldsymbol{\Delta}\beta + \mathbf{A}^{(i)} \bmod p,\ (\mathsf{id}, i)) \oplus (\mathsf{K}_{i,j}(\beta_j) \text{ for } j \in [\log p])$$

for each $j \in [\log p], \beta \in [p]$. Notice that, every potential boolean label, such as $\mathsf{K}_{i,j}(0)$, is encrypted in $O(p)$ ciphertexts. This is rather wasteful.

For better efficiency, $\mathsf{C}_{i,\beta+\alpha^{(i)} \bmod p}$ only encrypts a key $K_{0,\beta}$

$$\mathsf{C}_{i,\beta+\alpha^{(i)} \bmod p} \leftarrow \mathsf{H}(\boldsymbol{\Delta}\beta + \mathbf{A}^{(i)} \bmod p,\ (\mathsf{id}, i)) \oplus K_{0,\beta}.$$

The key $K_{0,\beta}$ is sampled by the garbler, and can decrypt the ciphertext

$$\mathsf{Enc}(K_{0,\beta}, (\mathsf{K}_{i,0}(\beta_0), K_{1,\beta_{1:\log p}})),$$

which reveals the next boolean label and the next key $K_{1,\beta_{1:\log p}}$. That is, the garbler samples keys $K_{j,\beta_{j:\log p}}$ for every $j \in [\log p], \beta \in [p]$, and the table additionally includes ciphertexts

$$\mathsf{Enc}(K_{j,\beta_{j:\log p}}, (\mathsf{K}_{i,j}(\beta_j), K_{j+1,\beta_{j+1:\log p}}))$$

for every $j \in [\log p], \beta \in [p]$. The ciphertexts should be properly shuffled, and some color bits/digits should be introduced to help the evaluation.

After optimization, the table consists of $O(kp)$ ciphertexts, each of which is $O(\lambda)$ bit long, and k mini-tables for the mini bit-composition, each of which is $O(\lambda k \log p)$ bit long. The total table size is $O(\lambda k(k \log p + p))$. It produces a statistically secure mixed GC in the random oracle model that has a marginal efficiency improvement compared to Theorem 1. But we will not explicitly state the further optimized gadget construction. The improvement is not significant enough to change the results in Table 1.

4.1 Extension: Linear BC and General BD

Our mixed GC for $\mathbb{Z}_{p^k}$ (Theorem 1) allows conversion between an arithmetic label and boolean labels of its base-p bit representation using bit-decomposition and bit-composition gadgets.

The base-p bit representation is quite useful, for example, it allows comparison between arithmetic numbers. But in many cases, we may need or may want to use the base-p' bit representation for a different base p'. The most naive solution is to use an expensive boolean circuit for base conversion. In this section, we presents an alternative solution.

BC. Let x be an arithmetic value. Given boolean labels of the base-p' bit representation of x, how to compute the $\mathbb{Z}_{p^k}$-arithmetic label of x? We ask a more general question:

> Given boolean labels of $(z_0, \ldots, z_{m-1})$, how to compute the $\mathbb{Z}_{p^k}$-arithmetic label of $\sum_i c_i z_m$, where $c_0, \ldots, c_{m-1}$ are fixed constants?

Essentially, we are asking how to garble gate $f : \{0,1\}^m \to \mathbb{Z}_{p^k}$, which is defined as $f(z_0, \ldots, z_{m-1}) = \sum_i c_i z_m \bmod p^k$.

The construction is rather straightforward. Let $\mathsf{K}_0, \ldots, \mathsf{K}_{m-1}$ be the input wire keys, let $\mathsf{AK}^{\mathsf{L}}(x) = \mathbf{A}x + \mathbf{B} \bmod p^k$ be the output wire key. Let

The gadget is parameterized by coefficients $c_0, \ldots, c_{m-1} \in \mathbb{Z}_{p^k}$.

Garbling algorithm $\mathsf{linBC.Garb}$ takes boolean keys $\mathsf{K}_0, \ldots, \mathsf{K}_{m-1}$, and an arithmetic key $\mathsf{AK}^{\mathsf{L}} = (\mathbf{A}, \mathbf{B})$ as inputs. Let α_i denote the mask bit of K_i.

- Sample random $\mathbf{B}_i$ for $i \in [m]$, satisfying $\sum_i \mathbf{B}_i \bmod p^k = \mathbf{B}$.
- For each $i \in [m]$, for each $\beta \in \{0,1\}$, compute

$$\mathsf{C}_{i,\beta+\alpha_i \bmod 2} \leftarrow \mathsf{H}(\mathsf{K}_i(\beta),\ (\mathsf{id}, i)) \oplus (\mathbf{A}c_i\beta + \mathbf{B}_i \bmod p^k)$$

- Output table $\mathsf{Tab} = (\mathsf{C}_{i,\beta})_{i\in[m],\beta\in\{0,1\}}$.

Evaluation algorithm $\mathsf{linBC.Eval}$ takes input labels $(\mathbf{l}_i, \bar{x}_i)$ for $i \in [m]$ and a table Tab as inputs.

- For $i \in [m]$, compute $\mathbf{L}_i \leftarrow \mathsf{H}(\mathbf{l}_i,\ (\mathsf{id}, i)) \oplus \mathsf{C}_{i,\bar{x}_i}$.
- Output arithmetic label $\mathbf{L} = \sum_i \mathbf{L}_i \bmod p^k$.

Simulation algorithm $\mathsf{linBC.Sim}$ takes input labels $(\mathbf{l}_i, \bar{x}_i)$ for $i \in [m]$ and arithmetic label $\mathbf{L}$ as inputs.

- Sample random $\mathbf{L}_i$ for $i \in [m]$, satisfying $\sum_i \mathbf{L}_i \bmod p^k = \mathbf{L}$.
- The active ciphertexts in the table Tab are set as

$$\mathsf{C}_{i,\bar{x}_i} = \mathsf{H}(\mathbf{l}_i,\ (\mathsf{id}, i)) \oplus \mathbf{L}_i$$

  The rest are inactive ciphertexts, and are simulated by random strings.

Fig. 5. The Linear Bit-Composition Gadget over Ring $\mathbb{Z}_{p^k}$

$\mathbf{B}_0, \ldots, \mathbf{B}_{m-1}$ be an additive sharing of $\mathbf{B}$ that are sampled by the garbler. Given $\mathsf{K}_i(z_i)$, the evaluator can compute $\mathbf{L}_i = \mathbf{A}c_i z_i + \mathbf{B}_i \bmod p^k$ because the table contains

$$\mathsf{Enc}(\mathsf{K}_i(\beta), \mathbf{A}c_i\beta + \mathbf{B}_i)$$

for all $i \in [m], \beta \in \{0,1\}$. The evaluator outputs

$$\begin{aligned} \mathbf{L} := \sum_i \mathbf{L}_i \bmod p^k &= \sum_i (\mathbf{A}c_i z_i + \mathbf{B}_i) \bmod p^k \\ &= \mathbf{A}f(z_0, \ldots, z_{m-1}) + \mathbf{B} \bmod p^k. \end{aligned} \tag{3}$$

This is formalized in Fig. 5.

Lemma 4. *For any $f(z_0, \ldots, z_{m-1}) = \sum_i c_i z_m \bmod p^k$, there is a secure garbling gadget for general linear bit-composition function f (Fig. 5), called* linear bit-composition gadget*, in the random oracle model. The table size is $O(\lambda mk)$, assume the output label dimension is $\lambda/\log p$.*

Proof. For any input $z_0, \ldots, z_{m-1}$, the evaluator computes $\mathbf{L}_i \leftarrow \mathsf{H}(\mathbf{l}_i,\ (\mathsf{id}, i)) \oplus \mathsf{C}_{i,z_i\oplus\alpha_i}$, then $\mathbf{L}_i = \mathbf{A}c_i\beta + \mathbf{B}_i \bmod p^k$. The correctness of the output is guaranteed by (3).

To prove security, is suffices to notice that $\mathbf{B}_0, \ldots, \mathbf{B}_{m-1}$ is an additive sharing implies $\mathbf{L}_0, \ldots, \mathbf{L}_{m-1}$ is an additive sharing. In other words, we know $\mathbf{L}_0, \ldots, \mathbf{L}_{m-2}$ is i.i.d. uniform in the real world because they are one-time padded by i.i.d. uniform $\mathbf{B}_0, \ldots, \mathbf{B}_{m-2}$. And $\mathbf{L}_{m-1}$ is determined by $\mathbf{L}_0, \ldots, \mathbf{L}_{m-2}$ and $\mathbf{L}$ from $\mathbf{L} := \sum_i \mathbf{L}_i \bmod p^k$. □

BD. Given the $\mathbb{Z}_{p^k}$-arithmetic label of x, if we want to compute the boolean labels of the base-p' bit representation of x:

- First compute the boolean labels of the base-p bit representation of x, using bit-decomposition gadget.
- Compute the $\mathbb{Z}_{p'^{k'}}$-arithmetic label of x, using linear bit-composition gadget.
- Compute the boolean labels of the base-p' bit representation of x, using bit-decomposition gadget.

In particular, the cost of conversion from base-p bit representation to base-2 representation is $O(\lambda b^2)$ where $2^b \approx p^k$. This is much cheaper than using the boolean circuit for base conversion.

4.2 Extension: Emulating Computations for $\mathbb{Z}_N$

Our mixed GC for $\mathbb{Z}_{p^k}$ can emulate arithmetic mod-N operations if $p^k > N^2$ and there is an efficient garbling gadget for the modulo gate $\mathsf{mod}_N : \mathbb{Z}_{p^k} \to \mathbb{Z}_{p^k}$, which is defined as $\mathsf{mod}_N(x) = x \bmod N$. The emulation is rather straightforward:

- Every number in $\mathbb{Z}_N$ is emulated by the same number in $\mathbb{Z}_{p^k}$
- Every mod-N arithmetic operation (ADD or MULT) is emulated the by the same operation over $\mathbb{Z}_{p^k}$, followed by mod_N.

  Remark: The cost of emulating addition gates can be dramatically optimized. Instead of appending mod_N after every addition gate, append mod_N only if the accumulated magnitude is close to $p^k/2$ or when the fan-out includes a multiplication gate.

Garbling the modulo gate mod_N is mostly equivalent to garbling the integer division gate $\mathsf{div}_N : \mathbb{Z}_{p^k} \to \mathbb{Z}_{p^k}$, which is defined as $\mathsf{div}_N(x) = \lfloor x/N \rfloor$, since $\mathsf{mod}_N(x) = x - N \cdot \mathsf{div}_N(x)$.

Unfortunately, the garbling gadget for div_N is hard to construct.[4] We will define a similar gate div^*_N whose garbling gadget is efficient and also suffices for emulating mod-N computations. The definition of $\mathsf{div}^*_N(x)$ is inspired by a well-known optimization that reduce division by constant to multiplication and shifting.

[4] An efficient garbling gadget of div_N can be constructed based on the garbling gadget of div^*_N.

Lemma 5 (Generalization of [16]). *For any positive integers $N, p, k_{\mathsf{I}}, k_{\mathsf{E}}, m$ satisfying $p^{k_{\mathsf{I}}+k_{\mathsf{E}}} \le mN < p^{k_{\mathsf{I}}+k_{\mathsf{E}}} + p^{k_{\mathsf{E}}}$,*

$$\left\lfloor \frac{x}{N} \right\rfloor = \left\lfloor \frac{mx}{p^{k_{\mathsf{I}}+k_{\mathsf{E}}}} \right\rfloor \qquad \text{for all } 0 \le x < p^{k_{\mathsf{I}}}.$$

Proof. $p^{k_{\mathsf{I}}+k_{\mathsf{E}}} \le mN < p^{k_{\mathsf{I}}+k_{\mathsf{E}}} + p^{k_{\mathsf{E}}}$ implies, by multiplying $\frac{x}{p^{k_{\mathsf{I}}+k_{\mathsf{E}}}N}$,

$$\frac{x}{N} \le \frac{mx}{p^{k_{\mathsf{I}}+k_{\mathsf{E}}}} < \frac{x}{N} + \frac{x}{Np^{k_{\mathsf{I}}}} < \frac{x+1}{N}.$$

□

Now we are ready to define the gate $\mathsf{div}^*_N : \mathbb{Z}_{p^{2k+1}} \to \mathbb{Z}_{p^{2k+1}}$. Let $k_{\mathsf{E}} := \lceil \log_p(N) \rceil$ be the minimum integer satisfying $p^{k_{\mathsf{E}}} \ge N$. Let $m = \lceil \frac{p^{k_{\mathsf{I}}+k_{\mathsf{E}}}}{N} \rceil$, thus $p^{k_{\mathsf{I}}+k_{\mathsf{E}}} \le mN < p^{k_{\mathsf{I}}+k_{\mathsf{E}}} + N \le p^{k_{\mathsf{I}}+k_{\mathsf{E}}} + p^{k_{\mathsf{E}}}$. By Lemma 5,

$$\left\lfloor \frac{x}{N} \right\rfloor = \left\lfloor \frac{mx}{p^{k+k_{\mathsf{E}}}} \right\rfloor$$

for any $0 \le x < p^k$. Therefore we define $\mathsf{div}^*_N : \mathbb{Z}_{p^{2k+1}} \to \mathbb{Z}_{p^{2k+1}}$ as

$$\mathsf{div}^*_N(x) = \left\lfloor \frac{mx \bmod p^{2k+1}}{p^{k+k_{\mathsf{E}}}} \right\rfloor.$$

It satisfies $\mathsf{div}^*_N(x) = \lfloor x/N \rfloor$ for all $x < p^k$. Since div^*_N is the composition of multiplication in $\mathbb{Z}_{p^{2k+1}}$ and digit shifting, it can be efficiently garbled by our mixed GC for $\mathbb{Z}_{p^{2k+1}}$.

Define gate $\mathsf{mod}^*_N : \mathbb{Z}_{p^{2k+1}} \to \mathbb{Z}_{p^{2k+1}}$ as $\mathsf{mod}^*_N(x) = x - N \cdot \mathsf{div}^*_N(x)$. Then mod^*_N can be efficiently garbled by our mixed GC for $\mathbb{Z}_{p^{2k+1}}$, and $\mathsf{mod}^*_N(x) = x \bmod N$ for all $x < p^k$.

Lemma 6. *For any $N \le 2^b$, there is a statistically secure mixed GC for $\mathbb{Z}_N$ in the random oracle model, such that each addition/multiplication/bit-decomposition/bit-composition gate costs $O(\lambda b^2 / \log b)$ communication. The bit-decomposition is over a prime base $p = \Theta(b/\log b)$.*

Proof. Mod-N computations can be emulated in a $\mathbb{Z}_{p^{2k+1}}$-mixed circuits. Combing with Theorem 1, the cost per gate is $O(\lambda k(k+p)\log p)$. The cost is minimized by letting $p = \Theta(b/\log b)$. □

Remarks. Although Lemma 6 does not claim free addition, we observe from its construction that addition is free up to a certain extent.

In this mixed GC for $\mathbb{Z}_N$, the bit decomposition gate outputs base-p bit representations. In case a (base-2) bit representation is needed, it can be computed from the base-p bit representation by a cost of $O(\lambda b^2)$, using the trick stated in Sect. 4.1.

5 Mixed GC Based on Chinese Remainder Theorem

Chinese remainder theorem (CRT) is used in [12] to solve the following natural task: Given b, find an efficient arithmetic GC over ring $\mathbb{Z}_N$ for some $N \approx 2^b$.

Since there is no more specific constraints on N, [12] sets $N = p_1 p_2 \dots p_s$ being the product of the first s primes. Then $s = \Theta(b/\log b)$ and $p_s = \Theta(b)$. Consider an arithmetic circuit over $\mathbb{Z}_{p_i}$, denoted by "$C \bmod p_i$", that is identical to C except the ring is replaced by $\mathbb{Z}_{p_i}$. Then

$$C(x) \bmod p_i = (C \bmod p_i)(x \bmod p_i).$$

Therefore, by CRT, the task of evaluating $C(x)$ is reduced to evaluating mod-p_i arithmetic circuit $(C \bmod p_i)(x \bmod p_i)$ for all $1 \le i \le s$. In [12], the reduction is combined with mixed GC for every ring $\mathbb{Z}_{p_i}$, resulting in an arithmetic GC for $\mathbb{Z}_N$ where each multiplication gate costs about $O(\lambda b^2/\log b)$ bits.

In this section, we will strengthen the result in two dimensions.

Based on Mod-p^k Mixed GC. [12] sets $N = p_1 p_2 \dots p_s$ because their basic GC only supports computation modulo a prime number. In Sect. 4, we have already construct relatively efficient mixed GC for prime power rings. Therefore, we will set

$$N = p_1^{k_1} p_2^{k_2} \dots p_s^{k_s} \approx 2^b$$

and reduce the problem of garbling mod-N computation to garbling mod-$p_i^{k_i}$ computations for each $1 \le i \le s$.

Efficient BD. In the CRT framework, if the actual value of a $\mathbb{Z}_N$-wise is x, it is not hard to get the boolean labels of the bit representation of $x \bmod p_i^{k_i}$ (via Theorem 1), for each $1 \le i \le s$. To compute the bit representation of x, the naive idea is garble the CRT algorithm.

For more efficient bit-decomposition, we make the following observation. There are constants $c_1, \dots, c_s \in \mathbb{Z}_N$ such that, for any $x \in \mathbb{Z}_N$

$$x = \sum_i c_i x^{(i)} \bmod N,$$

where $x^{(i)} := x \bmod p_i^{k_i}$ denotes the mod-$p_i^{k_i}$ component of x. $(x^{(1)}, \dots, x^{(s)})$ is usually called the *CRT representation* of x. The fact that x is a linear function (modulo N) on its CRT representation suggests a more efficient bit-decomposition construction in the "CRT framework".

Our new bit-decomposition construction is essentially a mixed circuit over the ring $\mathbb{Z}_{p^{2k+1}}$, where p, k satisfy $p^k > N^2 > \sum_i c_i x^{(i)}$. The input of the mixed circuits consists of the bit representation of $x^{(i)}$ for all $1 \le i \le s$. All the input wires can be merged into $\sum_i c_i x^{(i)}$ through the generalized linear BC gate (Fig. 5). Then next step is mod_N^*, whose output $\sum_i c_i x^{(i)} \bmod N$ always equals x. The last gate is the standard bit-decomposition of $\mathcal{C}_{\mathsf{mix}}(\mathbb{Z}_{p^k})$, producing the base-$p$ bit representation of x.

The linear BC costs λmk bits, where $m = \sum_i k_i \log p_i = O(b)$. The modulo gate mod_N^* and bit-decomposition gate cost $O(\lambda b(k+p))$. The overall cost is $O(\lambda b(k+p))$, which can be minimized as $O(\lambda b^2/\log b)$ by setting $p = \Theta(b/\log b)$.

If (base-2) bit representation of x is required, the overall cost of BD is $O(\lambda b^2)$.

By combining the "CRT framework" with Theorem 1 and Lemma 6 respectively, we have two more efficient mixed GC for $\mathbb{Z}_N$.

Theorem 2. *For any b, there exist $N > 2^b$ and a statistically secure mixed GC for $\mathbb{Z}_N$ in the random oracle model, such that each addition gate costs no communication, and each multiplication gate costs $O(\lambda b^{1.5})$ communication, and each bit-decomposition/bit-composition gate costs $O(\lambda b^2/\log b)$ communication.*

Proof. Set $N = p_1^{k_1} p_2^{k_2} \dots p_s^{k_s} \approx 2^b$. The task of garbling mod-N mixed circuits is reduced to garbling mod-$p_i^{k_i}$ mixed circuits for all $1 \le i \le s$. Each mod-$p_i^{k_i}$ mixed circuit will be garbled the mixed GC in Theorem 1.

Thus each mod-N addition gate will cost nothing.

Each mod-N multiplication gate costs

$$\sum_i O(\lambda k_i (k_i + p_i) \log p_i).$$

We want to minimize the cost, under the constraint that $p_1^{k_1} p_2^{k_2} \dots p_s^{k_s} \approx 2^b$.

For any i, if k_i increases by 1, then $\log N$ will increase by $\log p_i$, the total cost will increase by $O(\lambda(k_i + p_i)\log p_i)$. The "marginal cost increase per bit of N by changing k_i" is

$$\frac{\partial \mathrm{cost}(k_1, \dots, k_s)}{\partial k_i} \Big/ \frac{\partial \log N(k_1, \dots, k_s)}{\partial k_i} = O(\lambda(k_i + p_i)).$$

To minimize the cost, this ratio should be roughly the same for all i.

Following this intuitive argument, we choose a constant c and let $p_i + k_i = c$ for all i. The value of c is determined by the constraint $N = p_1^{k_1} p_2^{k_2} \dots p_s^{k_s} \approx 2^b$.

$$b \le \log \prod_{i \le s} p_i^{k_i} = \sum_{i \le s} k_i \log p_i = \sum_{i \le s} (c - p_i) \log p_i \approx \sum_{p=2}^{c} (c - p) = \Theta(c^2).$$

Thus we set $c = \Theta(\sqrt{b})$.

The cost per multiplication gate is

$$\sum_i \lambda k_i (k_i + p_i) \log p_i = \sum_i \lambda (c - p_i) c \log p_i \approx \sum_{p=2}^{c} \lambda (c - p) c = O(\lambda c^3) = O(\lambda b^{1.5}).$$

The total cost of having one BD gate in the mod-$p_i^{k_i}$ part for all $1 \le i \le s$ is also $O(\lambda b^{1.5})$. But these parallel BD gates only compute (the bit representation of) the CRT representation. To compute the bit representation, an additional cost of $O(\lambda b^2)$ (or $O(\lambda b^2/\log b)$, if the representation can use any base) is needed.

For BC, say the boolean representation of the number has at most $O(b)$ bits. Applying linear BC (Fig. 5) for all $1 \le i \le s$ will cost $O(\sum_i \lambda b k_i)$ bits.

$$\sum_i \lambda b k_i = \lambda b \sum_i (c - p_i) \le \lambda b c s = O(\lambda b^2 / \log b)$$

□

Theorem 3. *For any b, there exist $N > 2^b$ and a statistically secure mixed GC for $\mathbb{Z}_N$ in the random oracle model, such that each addition/multiplication gate costs $O(\lambda b \log b / \log\log b)$ communication, each bit-decomposition costs $O(\lambda b^2 / \log b)$ communication, each bit-composition gate costs $O(\lambda b^2 / \log\log b)$ communication.*

Proof. Set $N = p_1^{k_1} p_2^{k_2} \dots p_s^{k_s} \approx 2^b$. The task of garbling mod-N mixed circuits is reduced to garbling mod-$p_i^{k_i}$ mixed circuits for all $1 \le i \le s$. Each mod-$p_i^{k_i}$ mixed circuit will be garbled with the mixed GC in Lemma 6.

Each mod-N addition/multiplication gate costs

$$\sum_i O(\lambda d_i^2 / \log d_i), \text{ where } 2^{d_i} > p_i^{k_i}.$$

We want to minimize the cost, under the constraint that $p_1^{k_1} p_2^{k_2} \dots p_s^{k_s} \approx 2^b$.

We choose a constant d such that $d = d_1 = d_2 = \dots = d_s$, and let $k_i = \lfloor d_i / \log p_i \rfloor$. So all primes are smaller than 2^d and $s = \Theta(2^d/d)$. The value of d is determined by the constraint $N = p_1^{k_1} p_2^{k_2} \dots p_s^{k_s} \approx 2^b$.

$$b \le \log \prod_{i \le s} p_i^{k_i} = \sum_{i \le s} k_i \log p_i \le \frac{1}{2} \sum_{i \le s} d = \Theta(sd) = \Theta(2^d).$$

Thus we set $d = \log b + O(1)$. Then $s = O(b / \log b)$.

The cost of each mod-N addition/multiplication gate is

$$\sum_i O\Big(\frac{\lambda d_i^2}{\log d_i}\Big) = O\Big(\frac{s \lambda d^2}{\log d}\Big) = O\Big(\frac{b \lambda \log b}{\log\log b}\Big).$$

The cost of BD, by the same analysis as in the proof of Theorem 2, is $O(\lambda b^2)$ if the outcome is base-2 bit representation, $O(\lambda b^2 / \log b)$ if the representation can use any base.

The cost of BC is trickier to trace. For each i, the mod-$p_i^{k_i}$ computations are emulated, according to the construction of Lemma 6, by a mod-p^k mixed circuit. Such that $k = O(d / \log d) = O(\log b / \log\log b)$. For each i, using linear BC to compute the arithmetic value costs $\lambda b k$. The total cost is $s \lambda b k = \lambda b^2 / \log\log b$. But linear BC computes a linear function modulo p^k, rather than the desired modulus $p_i^{k_i}$. This issue is resolve by slightly enlarge p^k to some $\text{poly}(p_i^{k_i}, b) = b^{\Theta(1)}$ so that linear BC computes the linear function over $\mathbb{Z}$. This modification over enlarge k by a constant factor, thus will not asymptotically increase the cost of any operations. □

6 Mixed GC Based on DCR

In this section, we show how to improve the efficiency of our mixed GC construction by relying on computational assumption. The new construction is most similar to the naive mixed construction (Lemma 3 in Sect. 3.3) over ring $\mathbb{Z}_{2^b}$.

The construction is built upon the (public-key) encryption schemes described in [17,18] based on the decisional composite residuosity (DCR) assumption [17, 19]. We consider two private-key variants described below. (We provide a brief overview of the DCR assumption and more details of the private-key variants in the full version.)

The first variant $\mathsf{Paillier}$ consists of four algorithms. (1) $\mathsf{Paillier.Setup}(1^\lambda, 1^\zeta)$ samples public parameters pp, which defines a key space $\mathbb{Z}$, a ciphertext space $\mathbb{Z}^*_{M^{\zeta+1}}$, and a message space $\mathbb{Z}_{M^\zeta}$. (2) $\mathsf{Paillier.Gen}(\mathsf{pp})$ outputs a secret key $\mathsf{sk} \in [\lfloor M/4 \rfloor]$. (3) $\mathsf{Paillier.Enc}(\mathsf{sk}, m)$ outputs a ciphertext $c \in \mathbb{Z}^*_{M^{\zeta+1}}$. (4) $\mathsf{Paillier.Dec}(\mathsf{sk}, c)$ recovers the message $m \in \mathbb{Z}_{M^\zeta}$.

The second variant DamJur is similar to the first, except in two aspects. First, the key space is $\mathbb{Z}^*_M$ instead of $\mathbb{Z}$. Second, the setup algorithm $\mathsf{DamJur.Setup}$ outputs a trapdoor tp in addition to the public parameters pp. The trapdoor is only used by a special inversion algorithm $\mathsf{DamJur.Inv}(\mathsf{tp}, c)$, which takes any ciphertext $c \in \mathbb{Z}^*_{M^{\zeta+1}}$ and outputs a unique secret key $\mathsf{sk} = g \in \mathbb{Z}^*_M$, and a message $m \in \mathbb{Z}_{M^\zeta}$ such that $\mathsf{DamJur.Dec}(\mathsf{sk} = g, c) = m$.

Both constructions have some kind of homomorphism. For any message m_1, m_2 and keys $\mathsf{sk}_1, \mathsf{sk}_2, g_1, g_2$.

$$\begin{aligned}
&\mathsf{Paillier.Enc}(\mathsf{sk}_1, m_1) \cdot \mathsf{Paillier.Enc}(\mathsf{sk}_2, m_2) \bmod M^{\zeta+1} \\
&\quad = \mathsf{Paillier.Enc}(\mathsf{sk}_1 + \mathsf{sk}_2 \text{ over } \mathbb{Z},\ m_1 + m_2 \bmod M^\zeta) \\
&\mathsf{DamJur.Enc}(g_1, m_1) \cdot \mathsf{DamJur.Enc}(g_2, m_2) \bmod M^{\zeta+1} \\
&\quad = \mathsf{DamJur.Enc}(g_1 \cdot g_2 \bmod M,\ m_1 + m_2 \bmod M^\zeta)
\end{aligned}$$

6.1 Bit-Composition Based on Paillier Encryption

As observed in Sect. 4.1, the more general bit-composition function $(x_0, \ldots, x_{m-1}) \to \sum_i c_i x_i \bmod 2^b$ is not harder to garble. Thus we will directly construct this more general bit-composition.

Let $\mathsf{K}_0, \ldots, \mathsf{K}_{m-1}$ be the boolean keys. Let $\mathsf{AK}^\mathsf{L} = (\mathbf{A} \in \mathbb{Z}^\ell_{2^b}, \mathbf{B} \in \mathbb{Z}^\ell_{2^b})$ be the arithmetic key. In the analysis of the complexity, we will assume $m = O(b)$ and $\ell = O(\lambda)$. For any $x_0, \ldots, x_{m-1} \in \{0,1\}$, given $\mathsf{K}_0(x_0), \ldots, \mathsf{K}_{m-1}(x_{m-1})$ and the table, the evaluator of the bit-composition gadget should output the arithmetic label $\mathbf{L} = \mathsf{AK}^\mathsf{L}(x) = x\mathbf{A} + \mathbf{B} \bmod 2^b$ where $x = \sum_i c_i x_i \bmod 2^b$.

The construction is based on the following intuition (informally): Allow the evaluator to decrypts $x + r$ and $(x + r)\mathsf{sk}^A + \mathsf{sk}^B$. Let the table contain

$$\mathsf{ct}^A = \mathsf{Enc}(\mathsf{sk}^A, \mathbf{A}), \quad \mathsf{ct}^B = \mathsf{Enc}(-r\mathsf{sk}^B, -r\mathbf{A} + \mathbf{B})$$

using some homomorphic encryption. Then the evaluator can compute

$$(\mathsf{ct}^A)^{x+r}\mathsf{ct}^B = \mathsf{Enc}((x + r)\mathsf{sk}^A + \mathsf{sk}^B, x\mathbf{A} + \mathbf{B})$$

The gadget is parameterized by coefficients $c_0, \ldots, c_{m-1} \in \mathbb{Z}_{2^b}$.

Garbling algorithm $\mathsf{BC.Garb}$ takes boolean keys $\mathsf{K}_0, \ldots, \mathsf{K}_{m-1}$, and an arithmetic key $\mathsf{AK}^{\mathsf{L}} = (\mathbf{A}, \mathbf{B})$ as inputs. Let α_i denote the mask bit of K_i.

- (global step) Generate $M, \zeta, g, p'q'$ using $\mathsf{vPai.Setup}$, while setting ζ such that $M^\zeta \geq 2^{\ell(2b+\lambda+1)}$. Add (M, ζ, g) to the beginning of the garbled circuit.
- Sample keys $\mathsf{sk}^A, \mathsf{sk}_0^B, \ldots, \mathsf{sk}_{m-1}^B \leftarrow [p'q']$. Let $\mathsf{sk}^B := \sum_i \mathsf{sk}_i^B$.
  Sample masks $r_0, \ldots, r_{m-1} \leftarrow [2^\lambda]$, $\mathbf{R} \leftarrow [2^{b+2\lambda}]^\ell$. Let $r = \sum_i c_i r_i$. Compute

$$\mathsf{ct}^A = \mathsf{vPai.Enc}(\mathsf{sk}^A, \mathbf{A}), \qquad \mathsf{ct}^B = \mathsf{vPai.Enc}(\mathsf{sk}^B, (2^{2b+\lambda} - r\mathbf{A}) + \mathbf{B} + 2^b\mathbf{R}).$$

- For each $i \in [m]$, for each $\beta \in \{0, 1\}$, compute

$$\mathsf{C}_{i,\beta+\alpha_i \bmod 2} \leftarrow \mathsf{H}(\mathsf{K}_i(\beta),\ (\mathsf{id}, i)) \oplus (x_i + r_i, c_i(x_i + r_i)\mathsf{sk}^A + \mathsf{sk}_i^B \bmod p'q')$$

- Output table $\mathsf{Tab} = ((\mathsf{C}_{i,\beta})_{i\in[m],\beta\in\{0,1\}}, \mathsf{ct}^A, \mathsf{ct}^B)$.

Evaluation algorithm $\mathsf{BC.Eval}$ takes input labels $(\mathbf{l}_i, \bar{x}_i)$ for $i \in [m]$ and a table Tab as inputs.

- For $i \in [m]$, compute $(\hat{x}_i, \mathsf{sk}_i) \leftarrow \mathsf{H}(\mathbf{l}_i,\ (\mathsf{id}, i)) \oplus \mathsf{C}_{i,\bar{x}_i}$.
- Compute $\mathsf{sk} = \sum_i \mathsf{sk}_i$, $\hat{x} = \sum_i c_i \hat{x}_i$.
- Output label $\mathbf{L} = \hat{\mathbf{L}} \bmod 2^b$, where $\hat{\mathbf{L}} = \mathsf{vPai.Dec}(\mathsf{sk}, (\mathsf{ct}^A)^{\hat{x}}\mathsf{ct}^B)$.

Simulation algorithm $\mathsf{BC.Sim}$ takes input labels $(\mathbf{l}_i, \bar{x}_i)$ for $i \in [m]$ and arithmetic label $\mathbf{L}$ as inputs.

- (global step) Sample (M, ζ, g) using the $\mathsf{vPai.Setup}$.
- Sample random $\hat{x}_0, \ldots, \hat{x}_{m-1} \leftarrow [2^\lambda]$. Let $\hat{x} = \sum_i c_i \hat{x}_i$.
  Sample random $\mathsf{sk}_0, \ldots, \mathsf{sk}_{m-1} \leftarrow [M/4]$. Let $\mathsf{sk} = \sum_i \mathsf{sk}_i$.
- Sample masks $\mathbf{R} \leftarrow [2^{b+\lambda}]^\ell$, let $\hat{\mathbf{L}} = \mathbf{L} + 2^b\mathbf{R}$.
  Simulate ct^A by randomly sample $\mathsf{ct}^A \leftarrow \mathsf{QR}_{M^{\zeta+1}}$. Simulate ct^B as

$$\mathsf{ct}^B = \mathsf{vPai.Enc}(\mathsf{sk}, \hat{\mathbf{L}})/(\mathsf{ct}^A)^{\hat{x}} \quad (\text{over } \mathbb{Z}^*_{M^{\zeta+1}}).$$

- The active ciphertexts in the table Tab are set as

$$\mathsf{C}_{i,\bar{x}_i} = \mathsf{H}(\mathbf{l}_i,\ (\mathsf{id}, i)) \oplus (\hat{x}_i, \mathsf{sk}_i)$$

  The rest are inactive ciphertexts, and are simulated by random strings.

The modifications with respect to Fig. 5 are highlighted.

Fig. 6. The Bit-Composition Gadget based on Paillier

which can be decrypted into $x\mathbf{A} + \mathbf{B}$.

To formalize the intuition: i) We will add large random noise $\mathbf{R}$, and let the evaluator get $x\mathbf{A} + \mathbf{B} + 2^b\mathbf{R}$ instead. ii) We need to construct an encryption scheme that has the required homomorphism.

As the section name suggested, the encryption scheme is (almost) Paillier. Except that we want the scheme to encrypt a vector rather than a number. We consider the following natural encoding $\mathsf{encode} : \mathbb{Z}^\ell \to \mathbb{Z}$, parameterized by ℓ and B,

$$\mathsf{encode}(v_0, \ldots, v_{\ell-1}) = \sum_{i \in [\ell]} B^i v_i,$$

together with an efficient decoder $\mathsf{decode} : [B^\ell] \to [B]^\ell$, satisfying

- For any $\mathbf{A}, \mathbf{B} \in \mathbb{Z}^\ell$, $\mathsf{encode}(\mathbf{A} + \mathbf{B}) = \mathsf{encode}(\mathbf{A}) + \mathsf{encode}(\mathbf{B})$.
- For any $\mathbf{A} \in [B]^\ell$, $\mathsf{encode}(\mathbf{A}) \in [B^\ell]$ and $\mathsf{decode}(\mathsf{encode}(\mathbf{A})) = \mathbf{A}$.

Set the parameter of the encoder by $B = 2^{2b+2\lambda+1}$. Define the following encryption scheme vPai,

- $\mathsf{vPai.Setup}(1^\lambda)$ is $\mathsf{Paillier.Setup}(1^\lambda, 1^\zeta)$, by choosing smallest ζ s.t. $M^\zeta \geq B^\ell$.
- $\mathsf{vPai.Gen}$ is $\mathsf{Paillier.Gen}$.
- $\mathsf{vPai.Enc}(\mathsf{sk}, \mathbf{V}) = \mathsf{Paillier.Enc}(\mathsf{sk}, \mathsf{encode}(\mathbf{V}))$.
- $\mathsf{vPai.Dec}(\mathsf{sk}, c) = \mathsf{decode}(\mathsf{Paillier.Dec}(\mathsf{sk}, c))$.

Using vPai, our intuition can be formalized as a bit-composition gadget.

Lemma 7. *For any linear bit-composition function* $f(z_0, \ldots, z_{m-1}) = \sum_i c_i z_m \bmod 2^b$ *satisfying* $\sum_i c_i \leq 2^b$ *(otherwise the construction should be slightly modified), there is a secure garbling gadget for* f *(Fig. 6), under DCR assumption in the random oracle model. The table size is* $O(m\lambda_{\mathsf{DCR}} + \ell(b + \lambda))$, *which is* $O(\lambda_{\mathsf{DCR}} b + \lambda^2)$ *when* $\ell = O(\lambda)$ *and* $m = O(b)$.

The proof of Lemma 7 is deferred to the full version.

6.2 Bit-Decomposition Based on Damgård-Jurik Encryption

In the bit-decomposition gadget, the evaluator is given an arithmetic label $\mathbf{L} = \mathsf{AK}(x) = x\boldsymbol{\Delta} + \mathbf{A} \bmod 2^b$, and its color number $\bar{x} = x + \alpha \bmod 2^b$ together with a table generated by the garbler from $\mathsf{AK}, \mathsf{K}_0, \ldots, \mathsf{K}_{b-1}$, and should output $\mathsf{K}_0(x_0), \ldots, \mathsf{K}_{b-1}(x_b)$.

Recall our intuition behind the naive BD (Fig. 1): In each inductive step, the evaluator gets $\mathbf{L}^{(i)} = x_{i:b}\boldsymbol{\Delta} + \mathbf{A}^{(i)}$ and computes

$$\mathbf{L}^{(i)} \bmod 2 = x_i \boldsymbol{\Delta} + \mathbf{A}^{(i)} \bmod 2.$$

Using $\mathbf{L}^{(i)} \bmod 2$ as the key, the evaluator decrypts a ciphertext

$$\mathsf{H}(x_i\boldsymbol{\Delta} + \mathbf{A}^{(i)} \bmod 2) \oplus (\mathsf{K}(x_i), x_i\boldsymbol{\Delta} + \mathbf{S})$$

in the table, gets $\mathsf{K}(x_i)$ and $x_i\boldsymbol{\Delta} + \mathbf{S}$. The latter allows the evaluator to compute $\mathbf{L}^{(i+1)}$ and proceed to the next step.

The bottleneck is the ciphertext size. Let us replace the ciphertext by

$$\mathsf{H}(x_i\boldsymbol{\Delta} + \mathbf{A}^{(i)} \bmod 2) \oplus (\mathsf{K}(x_i), x_i + r, (x_i + r)\mathsf{sk}^\Delta + \mathsf{sk}^S).$$

And let the table additionally contains two ciphertexts

$$\mathsf{ct}^{\Delta} = \mathsf{Enc}(\mathsf{sk}^{\Delta}, \mathbf{\Delta}), \quad \mathsf{ct}^{S} = \mathsf{Enc}(\mathsf{sk}^{S}, -r\mathbf{\Delta} + \mathbf{S} + 2^b\mathbf{R}),$$

using a homomorphic encryption scheme. Then the evaluator can instead compute $x_i\mathbf{\Delta} + \mathbf{S} + 2^b\mathbf{R}$ from

$$\mathsf{Dec}((x_i + r)\mathsf{sk}^{\Delta} + \mathsf{sk}^{S}, (\mathsf{ct}^{\Delta})^{x_i+r}/\mathsf{ct}^{S}).$$

Such modification does not improves the complexity yet, because $\mathsf{ct}^{\Delta}, \mathsf{ct}^{S}$ become the new dominating part. Notice that, all tables may share a global ct^{Δ} as it only depends on the global key.

For the last bottleneck ct^{S}, we require its distribution to be *"dense"*, in the sense that, the distribution of ct^{S} is statistically close to the uniform distribution over a samplable domain. This requires i) a "dense" encryption scheme, and ii) the distribution of the message $-r\mathbf{\Delta} + \mathbf{S} + 2^b\mathbf{R}$ is statistically close to uniform over the message space.

If our requirement is satisfied, the garbler can instead sample a random seed, and let $\mathsf{ct}^{S} = \mathsf{H}(\mathsf{seed})$. The ciphertext ct^{S} in the table can be replaced by seed. For correctness, the garbler need to reversely compute the key and message behind the ciphertext ct^{S}.

As discussed in [18], all of our requirements are satisfied by Damgård-Jurik encryption [17].

- *Density:* For random $g \leftarrow \mathbb{Z}_M^*$ and random $m \leftarrow [M^{\zeta}]$, the distribution of ciphertext $\mathsf{DamJur.Enc}(g, m)$ is uniform in $\mathbb{Z}_{M^{\zeta+1}}^*$.
- *Invertibility:* There is an efficient algorithm Inv, which takes a ciphertext $\mathsf{ct} \in \mathbb{Z}_{M^{\zeta+1}}^*$ and the trapdoor tp, computes g, m such that $\mathsf{DamJur.Enc}(g, m) = \mathsf{ct}$.

Damgård-Jurik encrypts a number rather a vector. Similar to Sect. 6.1, we need a encoder-decoder pair between vectors and numbers. The encoder has to be dense in the sense that almost all encodings in the codomain are valid. Again, consider the natural encoding $\mathsf{encode} : \mathbb{Z}^{\lambda+1} \to \mathbb{Z}$, parameterized by B,

$$\mathsf{encode}(v_0, \ldots, v_\lambda) = \sum_{i \in [\lambda+1]} B^i v_i,$$

together with an efficient decoder $\mathsf{decode} : [B^{\lambda}] \to [B]^{\lambda}$.

- For security, set $B \geq 2^{b+2\lambda}$.
- For density, ensure $M^{\zeta} \geq B^{\lambda+1} \geq M^{\zeta}(1 - 2^{-\lambda})$.

Define the following encryption scheme vDJ,

- $\mathsf{vDJ.Setup}(1^{\lambda})$ is $\mathsf{DamJur.Setup}(1^{\lambda}, 1^{\zeta})$, by choosing smallest ζ s.t. $M^{\zeta} \geq (2^{2b+\lambda+1})^{\lambda+1}$. Also let B be the largest multiple of 2^b satisfying $M^{\zeta} \geq B^{\lambda+1}$. Then all the three requirements on B can be satisfied.
- $\mathsf{vDJ.Gen}$ is $\mathsf{DamJur.Gen}$.
- $\mathsf{vDJ.Enc}(\mathsf{sk}, \mathbf{V}) = \mathsf{DamJur.Enc}(\mathsf{sk}, \mathsf{encode}(\mathbf{V}))$.

Garbling algorithm $\mathsf{BD.Garb}$ takes an arithmetic key $\mathsf{AK} = (\mathbf{A}, \alpha)$ and b boolean keys $\mathsf{K}_0, \ldots, \mathsf{K}_{b-1}$ as inputs.

- (global step) Generate $M, \zeta, p'q'$ using $\mathsf{vDJ.Setup}$, while setting ζ, B properly. Sample key $g^{\Delta} \leftarrow \mathbb{Z}_M^*$ and compute $\mathsf{ct}^{\Delta} = \mathsf{vDJ.Enc}(g^{\Delta}, (\boldsymbol{\Delta}, 1))$. Add $M, \zeta, \mathsf{ct}^{\Delta}$ to the beginning of the garbled circuit.
- Let $\mathbf{A}^{(0)} = \mathbf{A}$, $\alpha^{(0)} = \alpha$.
- For each $0 \le i < b$, sample $r_i \leftarrow [2^{\lambda}]$, $\mathsf{seed}^{(i)} \leftarrow \{0,1\}^{\lambda}$. Compute $\mathsf{ct}^{(i)} = \mathsf{H}(\mathsf{seed}^{(i)}, (\mathsf{id}, i)) \in \mathbb{Z}_{M^{\zeta+1}}$. Find $g^{(i)}, \mathbf{S}^{(i)}, s^{(i)}$ satisfying

$$(g^{(i)}, (2^{b+\lambda} - r_i(\boldsymbol{\Delta}, 1)) + (\mathbf{S}^{(i)}, s^{(i)}) = \mathsf{vDJ.Inv}(\mathsf{tp}, \mathsf{ct}^{(i)})\,.$$

  Resample $\mathsf{seed}^{(i)}$ if $(\mathbf{S}^{(i)}, s^{(i)}) \notin [B - 2^{b+\lambda}]^{\lambda+1}$ to prevent overflow. Set

$$\mathbf{A}^{(i+1)} = \lfloor \frac{\mathbf{A}^{(i)} - \mathbf{S}^{(i)}}{2} \rfloor \bmod 2^{b-i-1} \qquad \alpha^{(i+1)} = \lfloor \frac{\alpha^{(i)} - s^{(i)}}{2} \rfloor \bmod 2^{b-i-1}\,.$$

- For each $0 \le i < b$, for each $\beta \in \{0,1\}$, compute

$$\mathsf{C}_{i,\beta+\alpha^{(i)} \bmod 2} \leftarrow \mathsf{H}(\boldsymbol{\Delta}\beta + \mathbf{A}^{(i)} \bmod 2,\ (\mathsf{id}, i)) \oplus (\mathsf{K}_i(\beta), \beta + r_i, (g^{\Delta})^{\beta + r_i} g^{(i)})$$

- Output table $\mathsf{Tab} = ((\mathsf{C}_{i,\beta})_{i \in [b], \beta \in \{0,1\}}, (\mathsf{seed}^{(i)})_{i \in [b-1]})$

Evaluation algorithm $\mathsf{BD.Eval}$ takes input label $(\mathbf{L}, \bar{x})$ and a table Tab as inputs.

- Let $\mathbf{L}^{(0)} := \mathbf{L}$, $\bar{x}^{(0)} = \bar{x}$.
- For $i = 0, 1, 2, \ldots, b-1$: Compute $(\mathbf{l}_i, \hat{x}_i, h^{(i)}) \leftarrow \mathsf{H}(\mathbf{L}^{(i)} \bmod 2,\ (\mathsf{id}, i)) \oplus \mathsf{C}_{i, \bar{x}^{(i)} \bmod 2}$. If $i < b-1$, compute $\mathsf{ct}^{(i)} = \mathsf{H}(\mathsf{seed}^{(i)}, \mathsf{id}, i)$, $(\mathbf{D}^{(i)}, d^{(i)}) \leftarrow \mathsf{vDJ.Dec}(h^{(i)}, (\mathsf{ct}^{\Delta})^{\hat{x}_i} \mathsf{ct}^{(i)})$

$$(\mathbf{L}^{(i+1)}, \bar{x}^{(i+1)}) = \lfloor ((\mathbf{L}^{(i)}, \bar{x}^{(i)}) - (\mathbf{D}^{(i)}, d^{(i)}) \bmod 2^{b-i})/2 \rfloor\,,$$

- Output boolean labels $\mathbf{l}_0, \mathbf{l}_1, \ldots, \mathbf{l}_{b-1}$.

Simulation algorithm $\mathsf{BD.Sim}$ takes arithmetic label $(\mathbf{L}, \bar{x})$ and boolean labels $\mathbf{l}_0, \mathbf{l}_1, \ldots, \mathbf{l}_{b-1}$ as inputs.

- (global step) Generate $M, \zeta, p'q'$ using $\mathsf{vDJ.Setup}$, while setting ζ, B properly. Simulate ct^{Δ} as a random ciphertext.
- Let $(\mathbf{L}^{(0)}, \bar{x}^{(0)}) = (\mathbf{L}, \bar{x})$.
- Sample random $\hat{x}_i \leftarrow [2^{\lambda}]$, $\mathsf{seed}^{(i)} \leftarrow \{0,1\}^{\lambda}$, $\mathbf{D}^{(i)} \leftarrow [B]^{\lambda}$, $d^{(i)} \leftarrow [B]$ for each $i \in [b-1]$. Program H so that

$$\mathsf{vDJ.Enc}(h^{(i)}, (\mathbf{D}^{(i)}, d^{(i)})) = (\mathsf{ct}^{\Delta})^{\hat{x}_i}\, \mathsf{H}(\mathsf{seed}^{(i)}, \mathsf{id}, i) \quad (\text{in } \mathbb{Z}_{M^{\zeta+1}})$$

- The active ciphertexts in the table Tab are set as

$$\mathsf{C}_{i, \bar{x}^{(i)} \bmod 2} = \mathsf{H}(\mathbf{L}^{(i)} \bmod 2,\ (\mathsf{id}, i)) \oplus (\mathbf{l}_i, \hat{x}_i, h^{(i)})$$

  The rest are inactive ciphertexts, and are simulated by random strings.

The modifications with respect to Fig. 1 are highlighted.

Fig. 7. The Bit-Decomposition Gadget based on Damgård-Jurik

– $\mathsf{vDJ.Dec}(\mathsf{sk}, c) = \mathsf{decode}(\mathsf{vDJ.Dec}(\mathsf{sk}, c))$.
– $\mathsf{vDJ.Inv}(\mathsf{tp}, c) = (g, \mathsf{decode}(v))$ for $(g, v) = \mathsf{DamJur.Inv}(\mathsf{tp}, c)$.

Now we are ready to present the bit-decomposition gadget in Fig. 7.

Lemma 8. *There is a secure bit-decomposition gadget (Fig. 7) over ring $\mathbb{Z}_{2^b}$, under DCR assumption in the programmable random oracle model. The table size is $O(b\lambda_{\mathsf{DCR}})$.*

The proof of Lemma 8 is deferred to the full version.

Combining the bit-composition gadget in Lemma 7 and the bit-decomposition gadget in Lemma 8 produces a mix GC scheme, as stated by the following theorem.

Theorem 4. *There is a secure mixed GC for $\mathbb{Z}_{2^b}$ under DCR assumption in the programmable random oracle model, such that each addition gate costs no communication, each multiplication/bit-decomposition gate costs $O(\lambda_{\mathsf{DCR}} b)$ communication, and each bit-composition gate costs $O(\lambda_{\mathsf{DCR}} b + \lambda^2)$ communication.*

Our mixed GC for $\mathbb{Z}_{2^b}$ implies a mixed GC for any $\mathbb{Z}_N$ for any $N \approx 2^b$, using the emulation technique discussed in Sect. 4.2.

Corollary 1. *For any $N \leq 2^b$, there is a secure mixed GC for $\mathbb{Z}_N$ under DCR assumption in the programmable random oracle model, such that each addition/multiplication/bit-decomposition gate costs $O(\lambda_{\mathsf{DCR}} b)$ communication, and each bit-composition gate costs $O(\lambda_{\mathsf{DCR}} b + \lambda^2)$ communication.*

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