Review Article

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Leveled homomorphic encryption schemes for homomorphic encryption standard

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Abstract: Homomorphic encryption allows for computations on encrypted data without exposing the underlying plaintext, enabling secure and private data processing in various applications such as cloud computing and machine learning. This article presents a comprehensive mathematical foundation for three prominent homomorphic encryption schemes: Brakerski–Gentry–Vaikuntanathan (BGV), Brakerski–Fan-Vercauteren (BFV), and Cheon–Kim–Kim–Song (CKKS), all based on the ring learning with errors (RLWE) problem. We align our discussion with the functionalities proposed in the recent homomorphic encryption standard, providing detailed algorithms and correctness proofs for each scheme. In addition, we propose improvements to the current schemes focusing on noise management and optimization of public key encryption and leveled homomorphic computation. Our modifications ensure that the noise bound remains within a fixed function for all levels of computation, guaranteeing correct decryption and maintaining efficiency comparable to existing methods. The proposed enhancements reduce ciphertext expansion and storage requirements, making these schemes more practical for real-world applications.

Keywords: homomorphic encryption, learning with errors, ring learning with errors, noise bounds, lattice attacks

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1 Introduction

Homomorphic encryption describes encryption schemes that allow for addition and multiplication operations to be performed on ciphertexts without needing or leaking any information about the secret key or user messages. Furthermore, the operations in the ciphertext space correspond to performing the same operations on the original messages, which can be performed by any third party with knowledge of only the public information. Homomorphic encryption has several modern applications, such as secure cloud computing and private machine learning. With Craig Gentry's work in 2009 [1], secure homomorphic encryption became viable using ideal lattices. This construction closely relates to the commonly used learning with errors (LWE) problem, with the hardness of LWE being a result proved by Regev in 2005 [2].

Three of the most common modern homomorphic encryption schemes are based on a ring version of LWE problems, known as the ring learning with error (RLWE) problems [3]. These schemes are the Brakerski–Gentry–Vaikuntanathan (BGV) scheme [4,5], the Brakerski–Fan–Vercauteren (BFV) scheme [6], and the Cheon–Kim–Kim–Song (CKKS) scheme [7]. BGV and BFV allow for homomorphic computation for exact arithmetic, while CKKS provides homomorphic computation for numerical computation with certain accuracy. With

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recent efforts to standardize homomorphic encryption schemes and security [8,9], it is desirable to have concrete and mathematically solid discussions on encryption schemes and homomorphic computing protocols that match the functionalities proposed in the standard, including parameter specifications for efficiency and security.

We should mention that Chillotti et al. [10,11] present a fully homomorphic encryption scheme that can perform one bit operation in less than 0.1 s, while Gao [12] and Case et al. [13] present a fully homomorphic encryption scheme with similar running time but a much smaller ciphertext expansion (<20). However, each operation in these schemes is prohibitively expensive at the moment. Leveled schemes have a much larger ciphertext expansion, but each operation is much cheaper. A leveled scheme allows for some predetermined number of operations [5,14,15], which is the style we opt for in this article. Several works study noise bounds for homomorphic encryption [14,16–20] in both the canonical embedding and infinity norms. These analyses include both theory and implementations. Speedups can be implemented via the residue number system (RNS) [19, 21–23], which uses the Chinese remainder theorem to break down computations into smaller rings. The schemes we present in this article can all be implemented using RNS representation.

Our contributions. This article has two main goals. The first goal is to present detailed algorithms for the functionalities proposed in the homomorphic encryption standard [8,9] for each of the BFV, BGV, and CKKS schemes, and to present a detailed correctness proof for all the functionalities. This lays a rigorous mathematical foundation for homomorphic encryption schemes. The second goal is to improve the current schemes for BFV, BGV, and CKKS. We present modified schemes for each of the three schemes, especially in public key encryption and leveled homomorphic computing, and focus on noise control and the worst-case noise bounds, thereby reducing ciphertext expansion and storage expenses. In particular, under the modified schemes, the noise bound for ciphertexts from public key encryption and from homomorphic computing at each level is always bounded by a fixed function ρ, which depends on the underlying ring. The worst-case bound guarantees that ciphertexts can always be decrypted correctly, with no probability of decryption error, which is preferred for many applications. Furthermore, parameter sizes resulting from our worst-case bounds are comparable to parameter sizes from average-case bounds in the literature, thus not degrading the efficiency of the schemes.

Organization of this article. In Section 2, we describe notations and necessary background. We then introduce LWE and RLWE problems. We present and prove two variations of modulus reduction, which is later applied to RLWE-based encryption schemes. In Section 3, we outline three RLWE-based homomorphic encryption schemes: BFV, BGV, and CKKS. For these three schemes, we provide modified encryption to better control noise and conduct a thorough worst-case theoretical noise analysis. In Section 4, we discuss leveled schemes and present techniques in choosing parameters to guarantee homomorphic operations. We also outline operations in RNS here. In Section 5, we give a brief discussion of attack techniques for LWE problems. In Section 6, we provide concluding remarks and further potential research topics. Appendix A contains the proofs of the lemmas on correctness of functionalities for all the algorithms.

2 Notations and preliminaries

2.1 Notations

For a positive integer q, define $\mathbb{Z}_q \coloneqq \mathbb{Z} \cap (-q/2, q/2]$ to be the ring of centered representatives modulo q. For an element $v \in \mathbb{Z}$, we denote $[v]_q$ to be the modular reduction of v into the interval $\mathbb{Z}_q$ such that q divides $v - [v]_q$. When v is a vector or a polynomial, $[v]_q$ means reducing each entry or coefficient of v modulo q, respectively. Denote R_n as the ring

$$R_n = \mathbb{Z}[x]/(\phi(x)),$$

where $\phi(x)$ is a polynomial of degree n and $(\phi(x))$ is the ideal generated by $\phi(x)$. Often, we will choose $\phi(x)$ to be a power of two cyclotomic polynomial. That is, a polynomial of the form $\phi(x) = x^n + 1$, where n is a

power of two. For an integer q, we define $R_{n,q}$ as follows:

$$R_{n,q} = \mathbb{Z}_q[x]/(\phi(x)) \cong \mathbb{Z}[x]/(\phi(x), q),$$

where $(\phi(x), q)$ is the ideal generated by $\phi(x)$ and q. When v is a polynomial, $[v]_{\phi(x)}$ denotes modular reduction of the polynomial into R_n. Similarly, when v is a polynomial with integer coefficients, $[v]_{\phi(x),q}$ denotes modular reduction of the polynomial into $R_{n,q}$, where all the coefficients are in $(-q/2, q/2]$.

For a vector or polynomial v, the infinity norm of v, denoted $||v||_\infty$, is the maximum entry or coefficient in absolute value of v. Equivalently, if $v = (a_0, \ldots, a_{n-1})$ or $v = \sum_{i=0}^{n-1} a_i x^i$, then

$$||v||_\infty = \max\{|a_i| : i = 0, \ldots, n-1\}.$$

$||\cdot||_2$ denotes the standard 2-norm. The symbols $\lfloor\cdot\rfloor$ and $\lceil\cdot\rceil$ will denote floor and ceiling respectively, whereas $\lfloor\cdot\rceil$ will denote rounding to the nearest integer, rounding down in the case of a tie. When applying $\lfloor\cdot\rfloor$, $\lceil\cdot\rceil$, or $\lfloor\cdot\rceil$ to a polynomial or vector, we mean the rounding of each entry or coefficient. Define the expansion factor δ_R of R_n as follows:

$$\delta_R = \max\left\{\frac{||uv||_\infty}{||u||_\infty ||v||_\infty} : u, v \in R_n\right\},$$

where uv must be reduced modulo $\phi(x)$ before applying the norm. When $\phi(x) = x^n + 1$, where n is a power of two, it is well known that $\delta_R = n$.

2.2 Noise distributions and learning with errors problems

For a set S, we denote $\chi(S)$ as an arbitrary probability distribution χ on S. We denote U as the uniform distribution. Let $\rho > 0$ be any integer. We denote χ_ρ as any probability distribution on R_n, where each coefficient is random in $[-\rho, \rho]$ and independent. We call χ_ρ an *error distribution* or *noise distribution*. We allow for flexibility in the exact choice of χ_ρ, but most commonly, χ_ρ is chosen as uniform random on $[-\rho, \rho]$, or a truncated discrete Gaussian to maintain security [24,25]. Over $\mathbb{Z}$, the discrete Gaussian distribution $\mathcal{D}_{\mathbb{Z},aq}$ assigns a probability proportional to $\exp(-\pi x^2/(aq)^2)$ for each $x \in \mathbb{Z}$ with standard deviation $\sigma = aq/\sqrt{2\pi}$ [2,24]. For the cyclotomic polynomial $\phi(x) = x^n + 1$, an n-dimensional extension of $\mathcal{D}_{\mathbb{Z},aq}$ to R_n can be constructed by process of sampling each coefficient from $\mathcal{D}_{\mathbb{Z},aq}$. More details on the impact of the error distribution on security can be found in Section 5.

LWE problems. For any secret $s \in \mathbb{Z}_q^n$, we sample $e \leftarrow \chi(\mathbb{Z})$ from some desired distribution χ such that $||e||_\infty \le \rho$, where ρ is a desired parameter, we sample a uniform random $a \leftarrow U(\mathbb{Z}_q^n)$, and calculate b via $b \coloneqq [-\langle a, s\rangle + e]_q$. The ordered pair $(a, b) \in \mathbb{Z}_q^n \times \mathbb{Z}_q$ is called an *LWE sample*. The search-LWE problem is to find s given many LWE samples. The decision-LWE problem is given many samples that are either LWE samples or sampled at uniform random from $\mathbb{Z}_q^n \times \mathbb{Z}_q$, decide, which distributions the samples are drawn from [2,24].

When drawing elements from distributions on R_n and $R_{n,q}$, we can similarly define the RLWE problems. For a secret $s \in R_{n,q}$, sample a polynomial $e \leftarrow \chi_\rho$, sample $a \leftarrow U(R_{n,q})$, and compute $b \in R_{n,q}$ via $b \coloneqq [-as + e]_{\phi(x),q}$. The ordered pair $(a, b) \in R_{n,q}^2$ is called an *RLWE sample*. The RLWE problems can be defined in an analogous way to the LWE problems. Throughout this paper, when given an RLWE sample or similarly structured ordered pair, we will commonly refer to the polynomial e as *the noise term* and $||e||_\infty$ as *the noise size*.

Most leveled homomorphic encryption schemes use RLWE as opposed to LWE. The ciphertexts of homomorphic encryption schemes discussed will all essentially take the form of a modified RLWE sample. Regev originally showed the hardness of the LWE problem [2], which serves as foundation for the security of homomorphic encryption schemes. We discuss more specifics on security in Section 5.

We remark that, in our schemes, we use noise size $\rho = n$, and the noise distribution can be uniform on integers bounded by ρ or a discrete Gaussian distribution but truncated at ρ. When using a bounded uniform

distribution, our choice of noise has standard deviation σ of about $n/\sqrt{3}$. In comparison, most implementations in practice (including the homomorphic encryption standard) use a discrete Gaussian distribution with $\sigma \approx 3.2$ [8,9,19,26]. Our larger noise bound increases the security, which will be discussed later.

2.3 Modulus reductions

Let Q and q be any positive integers. Given an RLWE sample $(a_0, b_0) \in R_{n,Q}^2$, where $b_0 \equiv -a_0 s + e_0 \bmod (\phi(x), Q)$, we can compute a new RLWE sample $(a_0', b_0') \in R_{n,q}^2$ satisfying $b_0' \equiv -a_0' s + e_0' \bmod (\phi(x), q)$ for some new noise term e_0'. Although this is a new RLWE sample with a new integer modulus q, the key observation is that the polynomial s remains the same. Furthermore, if given an initial bound on e_0, we can guarantee a bound on e_0' dependent on Q and q. Algorithm 1 gives the procedure for modulus reduction, while Lemma 2.1 shows correctness and the resulting bound on e_0'. Although Q and q are any positive integers, we typically choose $Q > q$ and refer to this procedure as a *modulus reduction* rather than a modulus switch as many other papers do. Note here that we use a subscript of 0 in our RLWE sample, as it will provide more consistency with our later applications of this algorithm to ciphertexts. For this reason, we also label (a_0, b_0) as ct_0 in Algorithm 1.

Algorithm 1. BFV modulus reduction

	BFV.Modreduce (ct_0, Q, q)
Input:	$Q \in \mathbb{N}$ an integer, $q \in \mathbb{N}$ an integer, $\mathrm{ct}_0 = (a_0, b_0) \in R_{n,Q}^2$.
Output:	$\mathrm{ct}_0' = (a_0', b_0') \in R_{n,q}^2$.
Step 1.	Compute $a_0' \coloneqq \left\lfloor \frac{q a_0}{Q} \right\rceil$ and $b_0' \coloneqq \left\lfloor \frac{q b_0}{Q} \right\rceil$.
Step 2.	Return $\mathrm{ct}_0' = (a_0', b_0') \in R_{n,q}^2$.

Lemma 2.1. *Suppose the input* ct_0 *of Algorithm 1 is an RLWE sample such that* $\|e_0\|_\infty \le E$. *Let* ct_0' *be the output of Algorithm 1. Then*,

$$b_0' \equiv -a_0' s + e_0' \bmod (\phi(x), q)$$

and $\|e_0'\|_\infty \le \frac{q}{Q} E + \frac{\delta_R \|s\|_\infty + 1}{2}$. *Furthermore, if* $Q/q > \frac{2E}{\delta_R \|s\|_\infty - 1}$, *then* $\|e_0'\|_\infty < \delta_R \|s\|_\infty$.

What will be more useful than a modulus reduction for a generic RLWE sample will be a modulus reduction for a "modified" RLWE sample that takes the form of a standard BFV ciphertext, hence the algorithm name BFV.Modreduce. By a BFV ciphertext, we mean that our input for Algorithm 1 $\mathrm{ct}_0 = (a_0, b_0) \in R_{n,Q}^2$ satisfies

$$b_0 + a_0 s \equiv D_Q m_0 + e_0 \bmod (\phi(x), Q)$$

for some noise term $e_0 \in R_n$, where $m_0 \in R_{n,t}$ and D_Q is a positive integer. This format is further clarified in Section 3.1. Classic BFV [6] uses $D_Q = \lfloor Q/t \rfloor$. In this article, we will always assume $t|(Q-1)$ for any given ciphertext modulus Q, meaning that $D_Q = (Q-1)/t$ when using the same floor definition as in classic BFV. The resulting output $\mathrm{ct}_0' = (a_0', b_0') \in R_{n,q}^2$ of Algorithm 1 satisfies $b_0' + a_0' s \equiv D_q m_0 + e_0' \bmod (\phi(x), q)$ for some e_0', where $D_q = (q-1)/t$ when $t|(q-1)$. Algorithm 1 and the proof of Lemma 2.2 have a similar style to the proof of Lemma 2.3 in [12].

Lemma 2.2. *Suppose the input of Algorithm 1 is a BFV ciphertext such that* $||e_0||_\infty \le E$. *Let* ct_0' *be the output of Algorithm 1. If* $t|(Q-1)$ *and* $t|(q-1)$, *then*

$$b_0' + a_0's \equiv D_q m_0 + e_0' \bmod (\phi(x), q)$$

and $||e_0'||_\infty \le \frac{q}{Q}E + 1 + \frac{\delta_R||s||_\infty}{2}$. *Furthermore, if* $Q/q > \frac{2E}{\delta_R||s||_\infty - 2}$, *then* $||e_0'||_\infty < \delta_R||s||_\infty$.

The final reformulation essentially states that if Q/q meets a specific bound depending on E, modulus reduction always produces a new noise term e_0' bounded by $\delta_R||s||_\infty$. A similar algorithm and lemma can also be constructed for a standard BGV ciphertext [4,5], which is an ordered pair $\mathrm{ct}_0 = (a_0, b_0) \in R_{n,Q}^2$ satisfying

$$b_0 + a_0 s \equiv m_0 + te_0 \bmod (\phi(x), Q)$$

for some noise term e_0 and given message $m_0 \in R_{n,t}$, which is the message space for some integer $t > 1$. The procedure for computing the new ciphertext differs from the previous two modulus reduction algorithms. In particular, we use the procedure outlined in Lemma 4.3.1 of [27], which is given here as Algorithm 2.

Algorithm 2. BGV modulus reduction

	BGV.Modreduce(ct_0, Q, q)
Input:	$Q \in \mathbb{N}$, an integer, $q \in \mathbb{N}$, an integer, $\mathrm{ct}_0 = (a_0, b_0) \in R_{n,Q}^2$, BGV ciphertext.
Output:	$\mathrm{ct}_0' = (a_0', b_0') \in R_{n,q}^2$, BGV ciphertext.
Step 1.	Compute $\omega_a \coloneqq [-a_0 q t^{-1}]_Q$ and $\omega_b \coloneqq [-b_0 q t^{-1}]_Q$.
Step 2.	Compute $a_0' \coloneqq \left[\frac{q\ a_0 + t\ \omega_a}{Q}\right]_q$ and $b_0' \coloneqq \left[\frac{q\ b_0 + t\ \omega_b}{Q}\right]_q$.
Step 3.	Return $\mathrm{ct}_0' = (a_0', b_0') \in R_{n,q}^2$

Lemma 2.3. *Suppose the input of Algorithm 2 is a BGV ciphertext such that* $||e_0||_\infty \le E$. *Let* ct_0' *be the output of Algorithm 2. If* $t|(Q-1)$ *and* $t|(q-1)$, *then*

$$b_0' + a_0's \equiv m_0 + te_0' \bmod (\phi(x), q)$$

and $||e_0'||_\infty \le \frac{q}{Q}E + 1 + \frac{\delta_R||s||_\infty}{2}$. *Furthermore, if* $Q/q > \frac{2E}{\delta_R||s||_\infty - 2}$, *then* $||e_0'||_\infty < \delta_R||s||_\infty$.

3 Homomorphic encryption schemes and noise bounds

Most homomorphic encryption schemes in the literature use a modified version of RLWE to hide messages. In this section, we will cover three main schemes: BFV [6], BGV [4,5], and CKKS [7]. For these three schemes, we present modified versions where the noise sizes are improved and always controlled by a fixed bound, namely $\rho = \delta_R||s||_\infty$.

Overview of specifications. Before outlining our specific algorithms, we first provide an overview of specifications for parameters and spaces. Although there are variations between the schemes, the parameter choices outlined below work for all of BFV, BGV, and CKKS (when applicable). These parameter conditions ensure proper functionality regarding homomorphic computation for each scheme. Further caution must be taken when choosing parameters in practice to ensure security, which is discussed in Section 5.

Specifications for homomorphic encryption schemes

Public Parameters :	$q \in \mathbb{N}$
	$p_0 \in \mathbb{N}$ with $p_0 \geq 5\delta_R + 3$
	$p_1 \in \mathbb{N}$ with $p_1 \geq 6q$
	$t \in \mathbb{N}$ with $t\|(q-1), t\|(p_0-1)$, and $t\|(p_1-1)$
Plaintext :	$m \in R_{n,t}$ for BFV and BGV
	$m \in \mathbb{C}^{n/2}$ for CKKS
Secret key :	$\mathrm{sk} \in R_{n,3}$
Public key :	$\mathrm{pk} \in R^2_{n,p_0q}$
Evaluation key :	$\mathrm{ek} \in R^2_{n,p_1q}$
Ciphertext :	$\mathrm{ct} \in R^2_{n,q}$
Noise bound :	$\rho = \delta_R\|\|s\|\|_\infty$
Noise distribution :	χ_ρ, a probabilistic distribution on R_n
	with each coefficient random in $[-\rho, \rho]$.

We want to emphasize that each coefficient of the distribution χ_ρ can be uniform random on $[-\rho, \rho]$, or any sub-Gaussian truncated by the bound ρ, or any other distribution that is bounded by ρ. All the noise bounds in this article will be valid, since they depend only on the worst-case bound ρ. We will simply say "Sample $e \leftarrow \chi_\rho$" in all the algorithms for this generic distribution.

The bound ρ appears prominently in our algorithms. It is not just a bound on the noise distribution, but also a worst-case bound for both fresh ciphertexts from public key encryption and new ciphertexts after modulus reduction when going from one level to the next level, as indicated by the lemmas in the previous section. In practical implementations, we often choose n to be a power of 2 and $\phi(x) = x^n + 1$, hence $\delta_R = n$. When $\|s\|_\infty = 1$, we have $\rho = \delta_R\|s\|_\infty = n$.

Remark on message encoding and choice of t. In the BFV scheme, we choose to encode a message polynomial m as $D_q m$, where $D_q = (q-1)/t$, which requires that $q-1$ is divisible by t. Kim et al. [16] proposes to encode m as $\lfloor qm/t \rceil$, which works for any t and q, hence no restriction that $t|(q-1)$. An extra small amount of noise is introduced from their encryption style, but has minimal impact and gives about the same bounds in our lemmas in the case of t dividing $q-1$ that we consider. This alternate encryption style is slightly more expensive from a computation standpoint, as additional rounding operations must be performed as opposed to just integer multiplication. We refer the reader the study by Kim et al. [16] for more details on plaintext modulus choice and SIMD.

3.1 Modified BFV scheme

BFV key generation. The key generation process we use is slightly different from the standard BFV scheme [6] in that the public key and evaluation key are generated in a larger modulus [6,14,16,19], which will be useful for reducing the noise size in ciphertexts. Algorithm 3 gives the key generation for the BFV keys needed, which is the secret key sk, the public key pk, and the evaluation key ek. Here, sk is kept secret, while pk and ek are published. The public key $\mathrm{pk} = (\mathrm{k}_0, \mathrm{k}_1) \in R^2_{n,p_0q}$ satisfies

$$\mathrm{k}_1 + \mathrm{k}_0 s \equiv e \bmod (\phi(x), p_0 q)$$

for noise term $e \leftarrow \chi_\rho$. The evaluation key $\mathrm{ek} = (\mathrm{k}'_0, \mathrm{k}'_1) \in R^2_{n,p_1q}$ satisfies

$$\mathrm{k}'_0 + \mathrm{k}'_1 s \equiv p_1 s^2 + e'_1 \bmod (\phi(x), p_1 q)$$

for noise term $e'_1 \leftarrow \chi_\rho$. We remark that in Algorithm 3 we choose s randomly rather than specifying the sampling distribution. This is intentional, as s may be desired to be chosen to satisfy certain properties in practice. For instance, s is often chosen randomly with a predetermined Hamming weight in practice.

Algorithm 3. BFV key generation.

	BFV.Keygen(q, p_0, p_1)
Input:	$q \in \mathbb{N}$,
	$p_0 \in \mathbb{N}$ with $p_0 \geq 5\delta_R + 3$,
	$p_1 \in \mathbb{N}$ with $p_1 \geq 6q$.
Output:	$\mathrm{sk} = s \in R_{n,3}$ secret key,
	$\mathrm{pk} = (\mathrm{k}_0, \mathrm{k}_1) \in R^2_{n,p_0q}$ public key,
	$\mathrm{ek} = (\mathrm{k}'_0, \mathrm{k}'_1) \in R^2_{n,p_1q}$ evaluation key.
Step 1.	Choose randomly $s \in R_{n,3}$.
Step 2.	Sample $\mathrm{k}_0 \leftarrow U(R_{n,p_0q})$ and $e \leftarrow \chi_\rho$.
	Compute $\mathrm{k}_1 \coloneqq [-(\mathrm{k}_0 s + e)]_{\phi(x),p_0q}$.
Step 3.	Sample $\mathrm{k}'_1 \leftarrow U(R_{n,p_1q})$ and $e'_1 \leftarrow \chi_\rho$.
	Compute $\mathrm{k}'_0 \coloneqq [-\mathrm{k}'_1 s + p_1 s^2 + e'_1]_{\phi(x),p_1q}$.
Step 4.	Return $\mathrm{sk} = s$, $\mathrm{pk} = (\mathrm{k}_0, \mathrm{k}_1)$, and $\mathrm{ek} = (\mathrm{k}'_0, \mathrm{k}'_1)$.

BFV encryption and decryption. We encrypt a message $m_0 \in R_{n,t}$ using a modified version of the standard public key procedure for BFV. Note that we choose the plaintext modulus t so that t divides $p_0 - 1$ and $q - 1$, and therefore divides $p_0q - 1$. We immediately reduce the ciphertext modulus from p_0q to q before adding the message bits, then return the ciphertext. The purpose of this is to ensure the noise term on the returned ciphertext ct'_0 is within the constant bound of ρ. Given our description of the secret key selection, it is obvious that $||s||_\infty = 1$. Most results we provide can be easily modified for the case of general $||s||_\infty$ however, allowing for some flexibility in key generation if desired. The exact choices of p_0 and q will of course depend on several factors, such as the desired number of homomorphic computations. We will further discuss the choices of these in Section 4. Assuming these parameters, Algorithm 4 describes the encryption procedure for BFV. In many algorithms, we will refer to inputs as "BFV ciphertexts." By this, we mean some ordered pair $\mathrm{ct} = (a, b) \in R^2_{n,q}$ satisfying

$$b + as \equiv D_q m + e \bmod (\phi(x), q)$$

for some message $m \in R_{n,t}$, some noise term $e \in R_n$, ciphertext modulus q, and constant $D_q = \lfloor q/t \rfloor = (q - 1)/t$. This encryption style is the classic form of BFV encryption [6]. If $||e||_\infty \leq E$, we will say $\mathrm{ct} = (a, b)$ is a BFV ciphertext with noise bounded by E.

Algorithm 4. Modified BFV encryption

	BFV.Encrypt(m_0, D_q, pk)
Input:	$m_0 \in R_{n,t}$ message,
	$D_q \in \mathbb{N}$ constant,
	$\mathrm{pk} = (\mathrm{k}_0, \mathrm{k}_1) \in R^2_{n,p_0q}$ public key.
Output:	$\mathrm{ct}'_0 = (a'_0, b'_0) \in R^2_{n,q}$ BFV ciphertext.
Step 1.	Sample $u \leftarrow U(R_{n,3})$ and sample $e_1, e_2 \leftarrow \chi_\rho$.
Step 2.	Compute $(a_0, b_0) \in R^2_{n,p_0q}$ where
	$a_0 \coloneqq [\mathrm{k}_0 u + e_1]_{\phi(x),p_0q}$,
	$b_0 \coloneqq [\mathrm{k}_1 u + e_2]_{\phi(x),p_0q}$.

Step 3. Compute
$(a_0', b_0^*) \coloneqq \text{BFV.Modreduce}((a_0, b_0), p_0 q, q)$,
$b_0' \coloneqq [b_0^* + D_q m_0]_q$.
Step 4. Return $\text{ct}_0' = (a_0', b_0') \in R_{n,q}^2$.

Lemma 3.1 provides correctness and the corresponding noise bound resulting from encryption. The bounds in Lemma 3.1 are assuming that $||s||_\infty = ||u||_\infty = 1$. However, the bounds can be discussed in terms of more generic u and s, in which case the bound condition on p_0 is $p_0 > \frac{2\delta_R^2(||u||_\infty + ||s||_\infty) + 2\delta_R}{\delta_R||s||_\infty - 1}$. In this case, the resulting noise term from encryption still satisfies $||e_0'||_\infty < \rho$.

Lemma 3.1. *Let* ct_0' *be the output of Algorithm 4. Suppose that* $||s||_\infty = 1$, $t|(p_0 - 1)$, $t|(q - 1)$, *and* $p_0 > \frac{4\delta_R^2 + 2\delta_R}{\delta_R - 1}$. *Then* ct_0' *is a BFV ciphertext with noise bounded by* ρ.

We argue that when $\delta_R \geq 16$, the condition on p_0 in Lemma 3.1 is satisfied when p_0 is chosen so that $p_0 \geq 5\delta_R + 3$ as per our parameter specifications, since

$$5\delta_R + 3 > \frac{32}{7}\delta_R + \frac{16}{7} = \frac{16}{7}(2\delta_R + 1) = \frac{2\delta_R(2\delta_R + 1)}{\frac{7}{8}\delta_R} \geq \frac{4\delta_R^2 + 2\delta_R}{\delta_R - 1}.$$

This technique of encryption with a built-in modulus reduction in Step 3 was first mentioned in [19], but is overall not especially well outlined in the literature. Implementations do often reduce the modulus immediately after encryption to reduce noise. For instance, Microsoft SEAL [28] chooses p_0 as a "special prime," then generates all keys with an integer modulus of $p_0 q$ (for q a product of some primes) before reducing down to integer modulus q to house any ciphertext data. SEAL documentation recommends choosing this special prime p_0 to be at least as big as any prime divisor of q, though it is not a strict requirement. In our modification, we propose computing the modulus reduction locally during encryption and doing so before adding the message bits. The advantage to this approach is we can choose p_0 to be much smaller.

Algorithm 5 provides for the decryption of a BFV ciphertext, which is the standard BFV decryption.

Algorithm 5. BFV decryption

$\text{BFV.Decrypt}(\text{ct}_0, \text{sk})$

Input: $\text{ct}_0 = (a_0, b_0) \in R_{n,q}^2$ BFV ciphertext,
$\text{sk} = s \in R_{n,3}$ secret key.
Output: $m_0 \in R_{n,t}$ message.
Step 1. Compute $c \coloneqq [b_0 + a_0 s]_{\phi(x),q}$.
Step 2. Compute $m_0 \coloneqq \left[\left\lfloor \frac{tc}{q} \right\rceil\right]_t$.
Step 3. Return m_0.

Lemma 3.2. *If the input* ct_0 *of Algorithm 5 is a BFV ciphertext with noise bounded by* $(D_q - 1)/2$ *and* $t|(q - 1)$, *then the decryption in Algorithm 5 is correct.*

BFV additions and linear combinations. We allow for linear combinations of ciphertexts with scalars from $\mathbb{Z}$. Based on the sum of absolute values of these scalars, we can guarantee a bound on the resulting ciphertext noise. In particular, we discuss the case of linear combinations with scalars $\alpha_0, \ldots, \alpha_{k-1} \in \mathbb{Z}$ such that $\sum_{i=0}^{k-1} |\alpha_i| \leq M$. Assuming each input ciphertext has noise bounded by E, a linear combination of k ciphertexts using these scalars results in noise bounded by $M(E + 1)$. Algorithm 6 gives the algorithm for linear combinations, while Lemma 3.3 gives the resulting noise bound for BFV ciphertexts.

Algorithm 6. Linear combinations

	Linearcombo($\mathrm{ct}_0, \ldots, \mathrm{ct}_{k-1}, a_0, \ldots, a_{k-1}$)
Input:	$\mathrm{ct}_0, \ldots, \mathrm{ct}_{k-1} \in R_{n,q}^2$ (or $\mathrm{ct}_0, \ldots, \mathrm{ct}_{k-1} \in R_{n,q}^3$), $a_0, \ldots, a_{k-1} \in \mathbb{Z}$ scalars.
Output:	$\mathrm{ct}_0' \in R_{n,q}^2$ (or $\mathrm{ct}_0' \in R_{n,q}^3$).
Step 1.	Set $\mathrm{ct}_0' \coloneqq [0, 0]$ (or $[0, 0, 0]$). For i from 0 to $k-1$ do $\mathrm{ct}_0' \coloneqq [\mathrm{ct}_0' + a_i \mathrm{ct}_i]_q$.
Step 2.	Return ct_0'.

Lemma 3.3. *Suppose the inputs of Algorithm 6 are BFV ciphertexts each with noise bounded by E and suppose* $\sum_{i=0}^{k-1} |a_i| \leq M$. *Let* ct_0' *be the output of Algorithm 6. If* $t|(q-1)$, *then* ct_0' *is a BFV ciphertext with noise bounded by* $M(E+1)$.

We remark that we allow for inputs of Algorithm 6 to also be in $R_{n,q}^3$. For the input of Algorithm 6, when using elements of the form $(c_0, c_1, c_2) \in R_{n,q}^3$ satisfying

$$c_0 + c_1 s + c_2 s^2 \equiv D_q m + e \bmod (\phi(x), q)$$

for some $m \in R_{n,t}$ and $e \in R_n$, we still refer to e as a noise term (and refer to a bound on $||e||_\infty$ as a noise bound). If each input has noise bounded by E, we can slightly adjust the proof of Lemma 3.3 to obtain an element in $R_{n,q}^3$ with noise bounded by $M(E+1)$ from the output of Algorithm 6. This alternate choice of inputs will be utilized when discussing budgeted operations in Section 4.1.

BFV multiplication. As expected, multiplication incurs much bigger increase in ciphertext noise and more tedious noise analysis. The procedure is again standard for the BFV scheme as in the study by Fan and Vercauteren [6]. The proof is similar to that presented in the study by Fan and Vercauteren [6], but we give a simpler worst-case noise bound with Lemma 3.4.

Algorithm 7. BFV multiplication

	BFV.Multiply($\mathrm{ct}_0, \mathrm{ct}_1$)
Input:	$\mathrm{ct}_0 = (a_0, b_0), \mathrm{ct}_1 = (a_1, b_1) \in R_{n,q}^2$ BFV ciphertexts.
Output:	$(c_0', c_1', c_2') \in R_{n,q}^3$.
Step 1.	Compute $c_0 \coloneqq [b_0 b_1]_{\phi(x)}$, $c_1 \coloneqq [b_1 a_0 + b_0 a_1]_{\phi(x)}$, $c_2 \coloneqq [a_0 a_1]_{\phi(x)}$.
Step 2.	Compute $c_0' \coloneqq \lfloor t c_0 / q \rceil$, $c_1' \coloneqq \lfloor t c_1 / q \rceil$, $c_2' \coloneqq \lfloor t c_2 / q \rceil$.
Step 3.	Return (c_0', c_1', c_2').

Lemma 3.4. *Suppose the inputs of Algorithm 7 are BFV ciphertexts for messages* m_0 *and* m_1 *respectively, both with noise bounded by E. Let* (c_0', c_1', c_2') *be the output of Algorithm 7. If* $t|(q-1)$ *and* $\delta_R \geq 16$, *then*

$$c_0' + c_1' s + c_2' s^2 \equiv D_q [m_0 m_1]_{\phi(x),t} + e' \bmod (\phi(x), q), \tag{1}$$

with $||e'||_\infty \leq 3.5 E t \rho^2$.

The simple bound provided in Lemma 3.4 will allow us to choose parameters easily and stack moduli as we will do in Section 4, while having minimal influence on functionality.

Comparison to current bounds. As mentioned earlier, we can see that our bound is on the order of $Et\delta_R^2$ when choosing $||s||_\infty$=1, where E is the bound on the noise term of each ciphertext before multiplication. Comparing to more current works, [16] achieves a similar multiplication noise bound. Rather than restricting choices of t and q, [16] achieves this bound by an alternative encryption method, namely, by computing $b_0 + a_0 s \equiv \lfloor qm/t \rceil + e \bmod (\phi(x), q)$ rather than standard BFV, which computes $b_0 + a_0 s \equiv \lfloor q/t \rfloor m + e \bmod (\phi(x), q)$. We provide a short comparison of multiplication noise bounds, with e' being the noise term resulting from multiplication. Most notably, we assume two ciphertext noise terms are both bounded by E rather than having separate input bounds. Note this comparison does not include relinearization noise. We discuss the additional relinearization noise accumulated for our modified BFV scheme in Lemma 3.5.

Classic BFV [6]:

$$||e'||_\infty \le 2\delta_R t E(1 + \delta_R||s||_\infty) + 2\delta_R^2 t^2(||s||_\infty + 1)^2.$$

Improved BFV [16]:

$$||e'||_\infty \le \frac{\delta_R t}{2}\left(\frac{2E}{q} + (4 + \delta_R||s||_\infty)2E\right) + \frac{1 + \delta_R||s||_\infty + \delta_R^2||s||_\infty^2}{2}.$$

Our BFV variant:

$$||e'||_\infty \le 3.5Et\delta_R^2||s||_\infty^2.$$

In the study by Kim et al. [16], the dominant noise term is $\frac{\delta_R t}{2}(\delta_R||s||_\infty)2E = Et\delta_R^2||s||_\infty$. We note that, by going through their proof for the worst-case bound, the factor $\delta_R/2$ in their bound should be $\delta_R||s||_\infty$, which is what we used in our proof. Hence, the dominant term should be $2Et\delta_R^2||s||_\infty$. In comparison, our bound for all the noise terms is $3.5Et\delta_R^2||s||_\infty^2$, which is slightly bigger than their bound. Our goal was to provide a simple bound that is easier to use in practice. We will expand upon how we use this simple bound further in Section 4.

BFV relinearization. To convert a returned ciphertext from Algorithm 7 back to the proper form of a BFV ciphertext, we can employ a relinearization (or keyswitch) algorithm [6]. The algorithm converts a linear form in s and s^2 to a linear form in only s, while introducing a small additional noise term. Note that to accomplish this, we must use the published *evaluation key*, from Algorithm 3, denoted as ek. Algorithm 8 gives the relinearization algorithm for BFV. Lemma 3.5 provides for the resulting noise bound.

Algorithm 8. BFV Relinearization

	BFV.Relinearize((c_0', c_1', c_2'), ek)
Input:	$(c_0', c_1', c_2') \in R_{n,q}^3$ polynomial ordered triple,
	ek = $(\mathrm{k}_0', \mathrm{k}_1') \in R_{n,p_1q}^2$ evaluation key.
Output:	$(c_0, c_1) \in R_{n,q}^2$.
Step 1.	Compute $\beta_0 \coloneqq [c_2'\mathrm{k}_0']_{\phi(x),p_1q}$ and $\beta_1 \coloneqq [c_2'\mathrm{k}_1']_{\phi(x),p_1q}$.
Step 2.	Compute $d_0' \coloneqq \left\lfloor \frac{\beta_0}{p_1} \right\rceil$ and $d_1' \coloneqq \left\lfloor \frac{\beta_1}{p_1} \right\rceil$.
Step 3.	Compute $c_0 \coloneqq [c_0' + d_0']_q$ and $c_1 \coloneqq [c_1' + d_1']_q$.
Step 4.	Return (c_0, c_1).

Lemma 3.5. *Let* (c_0, c_1) *be the output of Algorithm 8 and suppose the input* (c_0', c_1', c_2') *satisfies* (1) *in* Lemma 3.4. *If* $p_1 \ge 6q$ *and* $\delta_R \ge 16$, *then* (c_0, c_1) *is a BFV ciphertext with noise bounded by* $3.6Etp^2$.

Alternate relinearization technique. Algorithm 8 is not the only option for relinearizing a ciphertext. Another technique [6,14,27,29] involves generating the evaluation key differently, by expanding c_2' with respect to some integer base. In this relinearization process, the evaluation key is a vector pair $\mathrm{ek} = (\mathbf{u}, \mathbf{v}) \in (R_{n,q}^{\gamma})^2$ where each entry of $\mathbf{u}$ is sampled from $U(R_{n,q})$, and γ and $\mathbf{v}$ are computed in the following way. For chosen public base $B \in \mathbb{N}$, find the smallest $\gamma \in \mathbb{N}$ such that $B^{\gamma} > q$ and define $\mathbf{g} \in R_{n,q}^{\gamma}$ as follows:

$$\mathbf{g}^T = (1, B, B^2, \ldots, B^{\gamma-1}).$$

Let $\mathbf{w} \in R_{n,q}^{\gamma}$ a vector with each entry sampled from χ_{ρ}. Compute $\mathbf{v}$ as

$$\mathbf{v} \coloneqq s^2\mathbf{g} - \mathbf{u}s + \mathbf{w} \pmod{\phi(x), q}.$$

To obtain a new relinearized ciphertext from (c_0', c_1', c_2'), one can first write

$$c_2' = \sum_{j=0}^{\gamma-1} h_j B^j,$$

where $h_j \in R_{n,q}$ such that $\|h_j\|_{\infty} \leq B/2$ and define $\mathbf{h}^T \in R_{n,q}^{\gamma}$ as $h = (h_0, h_1, \ldots, h_{\gamma-1})$. Then

$$c_2's^2 = \mathbf{h}(s^2\mathbf{g}).$$

Using $\mathrm{ek} = (\mathbf{u}, \mathbf{v})$, the new ciphertext can be computed as $([c_1' + \mathbf{hu}]_{\phi(x),q}, [c_0' + \mathbf{hv}]_{\phi(x),q})$ since $c_2's^2 + \mathbf{hw} \equiv \mathbf{hv} + \mathbf{hu}s \bmod (\phi(x), q)$. Here, $\mathbf{hw}$ is the noise introduced during relinearization and satisfies $\|\mathbf{hw}\|_{\infty} \leq (\gamma B \delta_R^2 \|s\|_{\infty})/2$. However, this technique is less used since the evaluation key $(\mathbf{u}, \mathbf{v}) \in (R_{n,q}^{\gamma})^2$ is much larger than the evaluation key generated in Algorithm 3. Specifically, $\mathrm{ek} = (\mathbf{u}, \mathbf{v})$ is of size $2\gamma \log_2 q$. The noise incurred by relinearization grows linearly with B; hence, B must be relatively small, which means γ will likely be much bigger than 4. On the other hand, $\mathrm{ek} = (\mathrm{k}_0', \mathrm{k}_1')$ from Algorithm 3 is of size $4\log_2 q$. Although Algorithm 3 gives a smaller key size, a larger ring dimension n must be used to maintain security. We refer the reader to the references mentioned earlier for details. Some implementations of BFV such as Microsoft SEAL [28] do not realinearize their ciphertexts after each multiplication and allow the degree of the linear form s to grow larger than 2 [30].

3.2 Modified BGV scheme

BGV key generation. As we did with BFV, we use a slightly different key generation process from the standard BGV scheme [4,5] by generating the public key and evaluation key in a larger modulus to reduce noise sizes in ciphertexts. Algorithm 9 gives the key generation for the BGV keys. Just like BFV, sk is kept secret, while pk and ek are published.

Algorithm 9. BGV key generation

	BGV.Keygen(q, p_0, p_1)
Input:	$q \in \mathbb{N}$,
	$p_0 \in \mathbb{N}$ with $p_0 \geq 5\delta_R + 3$,
	$p_1 \in \mathbb{N}$ with $p_1 \geq 6q$.
Output:	$\mathrm{sk} = s \in R_{n,3}$ secret key,
	$\mathrm{pk} = (\mathrm{k}_0, \mathrm{k}_1) \in R_{n,p_0q}^2$ public key,
	$\mathrm{ek} = (\mathrm{k}_0', \mathrm{k}_1') \in R_{n,p_1q}^2$ evaluation key.
Step 1.	Choose randomly $s \in R_{n,3}$.
Step 2.	Sample $\mathrm{k}_0 \leftarrow U(R_{n,p_0q})$ and $e \leftarrow \chi_{\rho}$.
	Compute $\mathrm{k}_1 \coloneqq [-(\mathrm{k}_0 s + te)]_{\phi(x),p_0q}$.

Step 3.	Sample $\mathrm{k}_1' \leftarrow U(R_{n,p_1q})$ and $e_1' \leftarrow \chi_\rho$.
	Compute $\mathrm{k}_0' \coloneqq [-\mathrm{k}_1's + p_1s^2 + te_1']_{\phi(x),p_1q}$.
Step 4.	Return sk = s, pk = $(\mathrm{k}_0, \mathrm{k}_1)$, and ek = $(\mathrm{k}_0', \mathrm{k}_1')$.

BGV encryption and decryption. We define the BGV public key encryption in Algorithm 10. Decryption of a BGV ciphertext is given in Algorithm 11. When we refer to a "BGV ciphertext" in these algorithms and lemmas, we mean an ordered pair $\mathrm{ct} = (a, b) \in R_{n,q}^2$ satisfying

$$b + as \equiv m + te \bmod (\phi(x), q)$$

for some noise term $e \in R_n$ and given message $m \in R_{n,t}$. Lemma 3.6 provides proof of correctness for encryption, as well as the corresponding noise bound resulting from encryption.

Algorithm 10. Modified BGV encryption

	BGV.Encrypt(m_0, pk)
Input:	$m_0 \in R_{n,t}$ message,
	$\mathrm{pk} = (\mathrm{k}_0, \mathrm{k}_1) \in R_{n,p_0q}^2$ public key.
Output:	$\mathrm{ct}_0' = (a_0', b_0') \in R_{n,q}^2$ BGV ciphertext.
Step 1.	Sample $u \leftarrow U(R_{n,3})$, and sample $e_1, e_2 \leftarrow \chi_\rho$.
Step 2.	Compute $(a_0, b_0) \in R_{n,p_0q}^2$, where
	$a_0 \coloneqq [\mathrm{k}_0u + te_1]_{\phi(x),p_0q}$,
	$b_0 \coloneqq [\mathrm{k}_1u + te_2]_{\phi(x),p_0q}$.
Step 3.	Compute
	$(a_0', b_0^*) \coloneqq \mathrm{BGV.Modreduce}((a_0, b_0), p_0q, q)$,
	$b_0' \coloneqq [b_0^* + m_0]_q$.
Step 4.	Return $\mathrm{ct}_0' = (a_0', b_0') \in R_{n,q}^2$.

Lemma 3.6. *Let* ct_0' *be the output of Algorithm 10. Suppose that* $\|s\|_\infty = 1$, $t|(p_0 - 1)$, $t|(q - 1)$, *and* $p_0 > \frac{4\delta_R^2 + 2\delta_R}{\delta_R - 2}$. *Then* ct_0' *is a BGV ciphertext with noise bounded by* ρ.

Algorithm 11. BGV decryption

	BGV.Decrypt(ct_0, sk)
Input:	$\mathrm{ct}_0 = (a_0, b_0) \in R_{n,q}^2$ BGV ciphertext,
	$\mathrm{sk} = s \in R_{n,3}$ secret key.
Output:	$m_0 \in R_{n,t}$ message.
Step 1.	Compute $c \coloneqq [b_0 + a_0s]_{\phi(x),q}$.
Step 2.	Compute $m_0 \coloneqq [c]_t$.
Step 3.	Return m_0.

Just as with BFV, the condition on p_0 can be discussed in the more general case for any choice of u and s, in which the condition on p_0 is $p_0 > \frac{2\delta_R^2(\|u\|_\infty + \|s\|_\infty) + 2\delta_R}{\delta_R\|s\|_\infty - 2}$ and the resulting noise term e_0' satisfies $\|e_0'\| < \rho$. Similar to BFV as well, we argue that when $\delta_R \geq 16$, the condition on p_0 in Lemma 3.6 is satisfied when p_0 is chosen so that $p_0 \geq 5\delta_R + 3$ as per our parameter specifications since

$$5\delta_R + 3 > \frac{32}{7}\delta_R + \frac{16}{7} = \frac{16}{7}(2\delta_R + 1) = \frac{2\delta_R(2\delta_R + 1)}{\frac{7}{8}\delta_R} \geq \frac{4\delta_R^2 + 2\delta_R}{\delta_R - 2}.$$

Regarding decryption, the proof of correctness for Algorithm 11 is straightforward. Simply observe that $[[b_0 + a_0 s]_{\phi(x),q}]_t = [m_0 + te_0]_t = m_0$. The key observation here is that in order for correctness to hold, it is required that $\|m_0 + te_0\|_\infty < q/2$. That is, fully reducing $b_0 + a_0 s$ modulo q will actually yield the correct polynomial $m_0 + te_0$. The worst-case bound on $\|m_0 + te_0\|_\infty$ is $t/2 + tE$ if $\|e_0\|_\infty \leq E$. Hence, it suffices to require $E < \frac{q}{2t} - \frac{1}{2}$. This is very similar to the condition given earlier needed for correct BFV decryption.

BGV additions and linear combinations. Additions and linear combinations for BGV can be done using Algorithm 6. The argument is similar to Lemma 3.3 for BFV ciphertexts and results in the same noise bound of $M(E + 1)$. It is worth noting that divisibility of $q - 1$ by t yields no noise improvement for BGV addition, and the noise bound of $M(E + 1)$ holds for any t and q.

BGV multiplication. Again, multiplication incurs large noise increase during homomorphic computation. Unlike BFV, there is no requirement that t divides $q - 1$, and the noise analysis for BGV multiplication is simpler than BFV, as no scaling by t/q is required after computing the necessary components given from ct_0 and ct_1. Algorithm 12 outlines the procedure for BGV multiplication. Lemma 12 provides for proof of correctness and the corresponding noise bound.

Algorithm 12. BGV multiplication

	BGV.Multiply(ct_0, ct_1)
Input:	$\mathrm{ct}_0 = (a_0, b_0)$, $\mathrm{ct}_1 = (a_1, b_1) \in R_{n,q}^2$ ciphertexts.
Output:	$(c_0', c_1', c_2') \in R_{n,q}^3$.
Step 1.	Compute $c_0' \coloneqq [b_0 b_1]_{\phi(x),q}$, $c_1' \coloneqq [b_1 a_0 + b_0 a_1]_{\phi(x),q}$, $c_2' \coloneqq [a_0 a_1]_{\phi(x),q}$.
Step 2.	Return (c_0', c_1', c_2').

Lemma 3.7. *Suppose the inputs of Algorithm 12 are BGV ciphertexts for messages m_0 and m_1, respectively, both with noise bounded by E. Let (c_0', c_1', c_2') be the output of Algorithm 12. Then*

$$c_0' + c_1' s + c_2' s^2 \equiv [m_0 m_1]_{\phi(x),t} + te' \bmod (\phi(x), q) \tag{2}$$

with $\|e'\|_\infty \leq 2\delta_R t(E^2 + 1)$.

BGV relinearization. We can relinearize a BGV ciphertext to rewrite the left hand side of equation (2) as a linear form in only s rather than s and s^2. For BGV, a slightly modified evaluation key must be generated, as well as a slightly modified relinearization algorithm. Algorithms 9 and 13 give the BGV evaluation key generation and relinearization respectively, which we have based on the algorithms in [16]. Lemma 3.8 provides proof of correctness of Algorithm 13 and the corresponding noise bound. Algorithm 13 and the result of Lemma 3.8 can be combined with Algorithm 12 and the result of Lemma 12, respectively, to obtain a full BGV multiplication operation.

Lemma 3.8. *Let* (c_0, c_1) *be the output of Algorithm 13 and suppose the input* (c_0', c_1', c_2') *satisfies* (2) *in* Lemma 3.7. *If* $p_1 \geq 6q$ *and* $\delta_R \geq 16$, *then* (c_0, c_1) *is a BGV ciphertext with noise bounded by* $2\delta_R t(E^2 + 1) + \frac{1}{8}\delta_R^2 \|s\|_\infty$.

Algorithm 13. BGV relinearization.

	BGV.Relinearize((c_0', c_1', c_2'), ek)
Input:	$(c_0', c_1', c_2') \in R_{n,q}^3$,
	ek = $(\mathrm{k}_0', \mathrm{k}_1') \in R_{n,p_1q}^2$ evaluation key.
Output:	$(c_0, c_1) \in R_{n,q}^2$.
Step 1.	Compute $\beta_0 \coloneqq [c_2'\mathrm{k}_0']_{\phi(x),p_1q}$ and $\beta_1 \coloneqq [c_2'\mathrm{k}_1']_{\phi(x),p_1q}$.
Step 2.	Compute $\omega_0 \coloneqq [-t^{-1}\beta_0]_{p_1}$ and $\omega_1 \coloneqq [-t^{-1}\beta_1]_{p_1}$.
Step 3.	Compute $d_0' \coloneqq \frac{\beta_0 + t\omega_0}{p_1}$ and $d_1' \coloneqq \frac{\beta_1 + t\omega_1}{p_1}$.
Step 4.	Compute $c_0 \coloneqq [c_0' + d_0']_q$ and $c_1 \coloneqq [c_1' + d_1']_q$.
Step 5.	Return (c_0, c_1).

3.3 Modified CKKS scheme

In this section, we will discuss the CKKS scheme [7]. CKKS allows for homomorphic encryption for arithmetic of approximate numbers rather than arithmetic exactly as BFV and BGV do. This is done by first taking in data as some vector over $\mathbb{C}$, mapping the components into R_n, and then performing the homomorphic computation before mapping back to a vector over $\mathbb{C}$. This process of mapping to and from the $\mathbb{C}$-vector space is known as the *encoding* and *decoding* procedures, respectively. Throughout Section 3.3 and whenever referring to CKKS, we will always assume $\phi(x) = x^n + 1$, where n is a power of two. Thus, $\delta_R = n$.

Message encoding and decoding. Recall that $R_n = \mathbb{Z}[x]/(\phi(x))$ and $R_{n,q} = \mathbb{Z}_q[x]/(\phi(x))$. Let $\mathbb{H} = \{z \in \mathbb{C}^n : z_j = \overline{z}_{n-j+1}\}$. Define two mappings:

$$\pi : \mathbb{H} \to \mathbb{C}^{n/2},$$
$$\sigma : \mathbb{C}[x]/(\phi(x)) \to \mathbb{C}^n.$$

Here, π is the projection of $\mathbb{H}$ onto $\mathbb{C}^{n/2}$, by keeping only the first half of the entries for each vector in $\mathbb{H}$, and σ is the canonical embedding map defined as follows. Note that the polynomial $\phi(x) = x^n + 1$ has n complex roots, say $\zeta_1, \zeta_2, \ldots, \zeta_n$ in any fixed order, which are all primitive roots of unity of order $2n$. Given a polynomial $h \in \mathbb{C}[x]/(\phi(x))$, σ is defined via

$$\sigma(h) = (h(\zeta_1), h(\zeta_2), \ldots, h(\zeta_n)) \in \mathbb{C}^n.$$

That is, σ evaluates h at all the roots of $\phi(x)$ and stores the evaluations as a vector. Note that both π and σ serve as isomorphisms of vector spaces over $\mathbb{C}$, so π^{-1} and σ^{-1} exist. In practice, σ is computed via a fast Fourier transform (FFT), and σ^{-1} by an inverse fast Fourier transform (FFT^{-1}).

The purpose of these mappings is that given a message vector $z \in \mathbb{C}^{n/2}$, we want to convert it into a polynomial in R_n whose values at ζ_i correspond to $w = \pi^{-1}(z)$, hence polynomial multiplication corresponds to component-wise multiplication for message vectors. We now must map $\pi^{-1}(z)$ into R_n. Given $\zeta_1, \zeta_2, \ldots, \zeta_n$ in a fixed order such that $(\zeta_1, \zeta_2, \ldots, \zeta_n) \in \mathbb{H}$, σ then serves as an isomorphism between $\mathbb{R}[x]/(\phi(x))$ and $\mathbb{H}$. So, for $w \in \mathbb{H}$, we can compute $\sigma^{-1}(w) \in \mathbb{R}[x]/(\phi(x))$ and then round each coefficient to obtain an element in R_n.

It is worth noting that most texts use a technique called *coordinate-wise random rounding* instead of rounding to the nearest integer [25]. However, we will use the closest integer rounding. As we will see, this step

of rounding causes accuracy loss in the message. To avoid this, we scale by some positive integer Δ to preserve some desired precision of our message in the end result. The message encoding function is defined as follows:

$$\mathrm{Ecd}(z, \Delta) = \lfloor \sigma^{-1}(\Delta \pi^{-1}(z)) \rceil \in R_n,$$

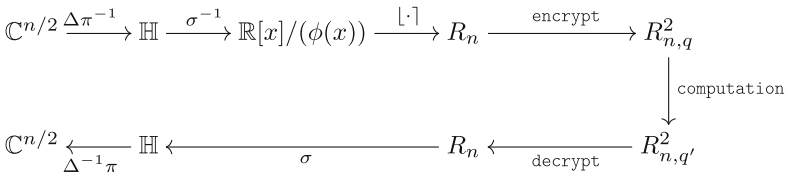
for any message $z \in \mathbb{C}^{n/2}$, and the message decoding function is defined as follows:

$$\mathrm{Dcd}(m, \Delta) = \pi(\sigma(\Delta^{-1} m)),$$

for any polynomial $m \in R_n$.

The encryption and decryption procedures for CKKS then map between R_n and $R_{n,q}$. A high level overview of the mappings in CKKS is shown below. Note that q' is used for the integer modulus of the ciphertext space after homomorphic computation, as we may have a different integer modulus if we perform any modulus reduction.

Overview of CKKS Mappings



Note the scaling factor Δ affects the ending precision and is usually chosen proportionally to the moduli gaps, which is discussed later in this section. It is also worth mentioning that although Ecd is defined for all messages $z \in \mathbb{C}^{n/2}$, in practice z is taken in the space of fixed precision numbers of some length, which is a subset of $\mathbb{C}^{n/2}$.

The remainder of Section 3.3 is devoted to the homomorphic computation in $R^2_{n,q}$ for the CKKS scheme. A significant observation for CKKS is how the homomorphic computation relates to the computation in $\mathbb{C}^{n/2}$. In particular, for vectors $z, z' \in \mathbb{C}^{n/2}$, we denote $z \circ z'$ as the Hadamard product of z and z' (i.e., the vector obtained from component-wise multiplication between z and z'). Homomorphic multiplication in $R^2_{n,q}$ of two ciphertexts corresponds with the Hadamard product of the respective vectors in $\mathbb{C}^{n/2}$, whereas homomorphic addition corresponds with standard vector addition of the respective message vectors.

CKKS rescaling. Regarding modulus reduction in CKKS, a similar procedure known as *rescaling* occurs. The rescaling procedure is identical to the modulus reduction for BFV outlined in Algorithm 1. That is,

$$\text{CKKS.Modreduce} = \text{BFV.Modreduce}.$$

The main difference is the purpose of the procedure. Rather than using modulus reduction as a form of noise control, it is used here to control precision. For two message encodings $m_0, m_1 \in R_n$, ciphertext multiplication yields an encryption of the product $m_0 m_1$, which takes up some less significant bits (LSBs). We rescale the corresponding ciphertext of $m_0 m_1$ to get rid of the lower significant digits to perform further computation where we want to keep only a fixed number of digits. While the rescaling serves a different purpose than modulus reduction, we can still discuss bounds on the corresponding error term achieved. Lemma 3.9 outlines our worst-case noise bound. When we refer to a "CKKS ciphertext" in these algorithms and lemmas, we mean an ordered pair $\mathrm{ct} = (a, b) \in R^2_{n,q}$ satisfying

$$b + as \equiv m + e \bmod (\phi(x), q)$$

for some noise term $e \in \mathbb{R}[x]/(\phi(x), q)$ and $m = \mathrm{Ecd}(z, \Delta) \in \mathbb{R}[x]/(\phi(x), q)$ for some $z \in \mathbb{C}^{n/2}$.

Lemma 3.9. *Suppose the input of Algorithm 1 is a CKKS ciphertext with noise bounded by E. Let* ct'_0 *be the output of Algorithm 1. Then,*

$$b'_0 + a'_0 s \equiv \frac{q}{Q} m_0 + e'_0 \bmod (\phi(x), q)$$

and $\|e'_0\|_\infty \leq \frac{q}{Q} E + \frac{1 + \delta_R \|s\|_\infty}{2}$. *Furthermore, if* $Q/q > \frac{2E}{\delta_R \|s\|_\infty - 1}$, *then* $\|e'_0\|_\infty < \rho$.

A notable difference in CKKS rescaling is that the algorithm returns an encryption of $\frac{q}{Q} m_0$ rather than the original m_0 encoding. As mentioned, this is intentional, as we wish to reduce the size of m_0 since bit usage

becomes an issue. The reason we use a modulus reduction algorithm rather than simply trying to scale down the ciphertext is because we are taking entries modulo Q. For a ciphertext $(a_0, b_0) \in R_{n,Q}^2$, if we computed a scaled ciphertext $(\lfloor a_0/\Delta \rceil, \lfloor b_0/\Delta \rceil)$ for some scaling factor Δ, we would first need to write $b_0 + a_0 s \equiv m_0 + e_0 + Qr$ for a polynomial $r \in R_n$. This would result in a term approximately equal to Qr/Δ after dividing through by Δ, which is no longer equivalent to 0 mod Q and would result in a huge noise term. However, it is still important to choose Q/q to be approximately Δ, or whatever desired scaling factor is needed. Accuracy of the approximation relies on this size of Q/q. When not concerned with RNS representation, we can simply choose $\Delta = Q/q$ exactly. In the RNS variant of CKKS [21], a bound is placed on the gap between Q/q and Δ to ensure some precision, while still allowing for coprime moduli Q and q.

CKKS key generation. For CKKS, the keys used are generated in the same way that the BFV keys are generated. In this case, we refer the reader to Algorithm 3 for generation of the CKKS keys, which again includes the secret key sk, the public key pk, and the evaluation key ek.

CKKS encryption and decryption. The encryption algorithm is given by Algorithm 14, and decryption by Algorithm 15. Note that CKKS encryption in Algorithm 14 uses Algorithm 1 as a subroutine, which is the rescaling. From Step 2 of Algorithm 15, we obtain m_0'. However, recall that $b_0 + a_0 s \equiv m_0 + e_0 \bmod (\phi(x), q)$, so in this case, we really have that $m_0' = m_0 + e_0$. In other words, m_0' is a close approximation of m_0 so long as $\|e_0\|_\infty$ is small. We do not include proofs that decryption works, as it is directly apparent from the algorithm that it decrypts to an approximation of the desired message. We also note that many texts, such as the original CKKS paper in the study by Cheon et al. [7], have separate steps for encoding/decoding and encryption/decryption. We include the encoding or decoding in the encryption or decryption algorithms, respectively. Aside from the encoding step, the encryption algorithm for CKKS is actually identical to a BFV encryption with $D_q = 1$. Lemma 3.10 provides for the corresponding noise bound after encryption. As with the other schemes, choosing $p_0 \geq 5\delta_R + 3$ ensures the condition for p_0 in Lemma 3.10 holds. The condition on p_0 can also be generalized to $p_0 > \frac{2\delta_R^2(\|u\|_\infty + \|s\|_\infty) + 2\delta_R}{\delta_R \|s\|_\infty - 1}$ with the resulting noise still satisfying $\|e_0'\| < \rho$.

Algorithm 14. Modified CKKS encryption

	CKKS.Encrypt(z, Δ, pk)
Input:	$z \in \mathbb{C}^{n/2}$ message, $\Delta \in \mathbb{N}$ scaling factor, pk = $(k_0, k_1) \in R_{n,p_0q}^2$ public key.
Output:	ct = $(a, b) \in R_{n,q}^2$ CKKS ciphertext.
Step 1.	Encode z by computing $m \coloneqq \text{Ecd}(z, \Delta)$.
Step 2.	Compute ct $\coloneqq$ BFV.Encrypt(m, 1, pk).
Step 3.	Return ct = $(a, b) \in R_{n,q}^2$.

Algorithm 15. CKKS decryption

	CKKS.Decrypt(ct, sk)
Input:	ct = $(a, b) \in R_{n,q}^2$ CKKS ciphertext, sk = $s \in R_{n,3}$ secret key.
Output:	$z \in \mathbb{C}^{n/2}$ message.
Step 1.	Compute $m \coloneqq [b + as]_{\phi(x),q}$.
Step 2.	Decode m by computing $z \coloneqq \text{Dcd}(m, \Delta)$.
Step 3.	Return z.

Lemma 3.10. *Let* ct *be the output of Algorithm 14. Suppose* $||s||_\infty = 1$ *and* $p_0 > \frac{4\delta_R^2 + 2\delta_R}{\delta_R - 1}$. *Then* ct *is a CKKS ciphertext with noise bounded by* ρ.

CKKS additions and linear combinations. We can perform additions and linear combinations with CKKS ciphertexts using Algorithm 6. The resulting ciphertext obtained from Algorithm 6 has a slightly different noise bound than BFV and BGV. The reason for this is that the encoded messages which the ciphertexts represent are in R_n instead of $R_{n,t}$, so reduction modulo t is not necessary with CKKS. The result is summarized in Lemma 3.11.

Lemma 3.11. *Suppose the inputs of Algorithm 6 are CKKS ciphertexts each with noise bounded by E and suppose* $\sum_{i=0}^{k-1}|a_i| \leq M$. *Let* ct_0' *be the output of Algorithm 6. Then,* ct_0' *is a CKKS ciphertext with noise bounded by ME*.

CKKS multiplication. Multiplication in CKKS follows the same process as BGV, which is given in Algorithm 12. Thus,

$$\text{CKKS.Multiply} = \text{BGV.Multiply}.$$

The difference is only a slightly different noise bound, due to the plaintexts not being in $R_{n,t}$ and thus not needing reduction modulo t. Lemma 3.12 outlines the result and proof of the corresponding noise bound.

Lemma 3.12. *Suppose the inputs of Algorithm 12 are CKKS ciphertexts for messages* m_0 *and* m_1, *respectively, both with noise bounded by E. Suppose that* $||m_0||_\infty \leq t/2$ *and* $||m_1||_\infty \leq t/2$. *Let* (c_0', c_1', c_2') *be the output of Algorithm 12. Then*

$$c_0' + c_1's + c_2's^2 \equiv m_0m_1 + e' \bmod (\phi(x), q) \tag{3}$$

with $||e'||_\infty \leq Et\delta_R + E^2\delta_R$.

CKKS relinearization. As with the other schemes, full multiplication can then be achieved by including the relinearization process discussed in Algorithm 8, so

$$\text{CKKS.Relinearize} = \text{BFV.Relinearize}.$$

The proof is almost identical to the proof in Lemma 3.5. The result for CKKS is given in Lemma 3.13.

Lemma 3.13. *Let* (c_0, c_1) *be the output of Algorithm 8 and suppose the input* (c_0', c_1', c_2') *satisfies* (3) *in* Lemma 3.12, *with* $||m_0||_\infty \leq t/2$ *and* $||m_1||_\infty \leq t/2$. *If* $p_1 \geq 6q$ *and* $\delta_R \geq 16$, *then* (c_0, c_1) *is a CKKS ciphertext with noise bounded by* $Et\delta_R + E^2\delta_R + \frac{1}{8}\delta_R^2||s||_\infty$.

Note on BFV versus CKKS. Initially, the formatting of ciphertexts in BFV and CKKS seem very similar. For a message $m_0 \in R_{n,t}$, a BFV ciphertext $\mathrm{ct}_0 = (a_0, b_0)$ satisfies $b_0 + a_0s \equiv D_qm_0 + e_0 \bmod (\phi(x), q)$. In CKKS, the encoding step for a message $z_0 \in \mathbb{C}^{n/2}$ scales our message by a factor of Δ. That is, our CKKS ciphertext $\mathrm{ct}_0 = (a_0, b_0)$ satisfies $b_0 + a_0s \equiv m_0 + e_0 \bmod (\phi(x), q)$, where $m_0 = \mathrm{Ecd}(z_0, \Delta)$. In both equations for BFV and CKKS, we have a scaling factor attached to our message. In BFV, the message m_0 is directly multiplied by D_q, while in CKKS, Δ multiplies the original message z_0 and is implicitly hiding in m_0. Nonetheless, both have a scaling factor. In multiplication of both schemes, this scaling factor initially compounds in the first step. The two schemes handle this issue differently however, with BFV rescaling the individual degree 2 ciphertext components in Step 2 of Algorithm 2, before taking the computed polynomials modulo q. CKKS on the other hand computes the initial polynomials in multiplication and immediately reduces them modulo q, relinearizes the ciphertext, and then rescales the hidden Δ^2 back to Δ using modulus reduction in Algorithm 1. Part of the reason these schemes differ in where they rescale is due to the fact that $D_q \gg \Delta$. Since $D_q = \lfloor q/t \rfloor$ and $t < q/2$, $D_q^2 > q$ and rescaling must occur before taking components modulo q. On the other hand, Δ in CKKS has more freedom in choice, as it is a parameter chosen by the user that can influence accuracy in the approximation of end result of computation.

3.4 Comparison to other noise bound analyses

Our noise analysis differs from previous works [6,7,16,17] in that we derive *worst-case* bounds based on worst-case bound assumptions on the error distribution. As a result, our correctness guarantees are deterministic; there is no probability of decryption error. In addition, we simplify the derived bounds into clean, closed-form expressions that will be useful for subsequent sections. This simplification comes at the cost of slightly looser bounds overall, with the effect being most pronounced in BFV multiplication. A detailed comparison of BFV multiplication appears near the end of Section 3.1.

For most operations, our bounds are very close to the worst-case results presented in the study by Kim et al. [16], though with a few important distinctions. First, in all of our modulus reduction lemmas, we explicitly bound the ratio Q/q to ensure that the noise remains below a fixed threshold, which supports more precise composability in computations. Second, our modified encryption procedure enables much smaller choices of p_0 by performing modulus reduction *before* message embedding. This results in fresh ciphertexts with noise bounded by a constant, while preserving correctness. The same structure can be extended to CKKS rescaling during encryption. Since the message bits are not yet introduced at that stage, rescaling does not degrade precision.

In addition, under basic assumptions on δ_R, we are able to significantly simplify the relinearization noise bounds. This is especially helpful in regimes with small plaintext modulus t, where relinearization noise can be more significant.

For concrete estimates, many works adopt $\delta_R = 2\sqrt{n}$ as an expansion factor that holds with high probability. In contrast, we use the exact value $\delta_R = n$ in later derivations. In the case of CKKS [7,17], the most significant differences in noise growth appear in fresh encryptions, relinearization (i.e., key switching), and rescaling. Across these operations, our analysis yields a dominant noise term of approximately n, compared to $\sqrt{n}$ in the aforementioned works. This is primarily due to our adoption of a strict worst-case model and the associated choice of expansion factor. For average-case analyses such as those presented in the study by Costache et al. [17], direct comparison is more nuanced, as noise growth depends on the variance of sampled noise terms.

4 Leveled schemes and RNS variants

For practical computation, we employ a *leveled homomorphic encryption scheme* rather than a fully homomorphic one. Unlike fully homomorphic schemes – which support an unlimited number of operations via costly bootstrapping – a leveled scheme supports a predetermined number of homomorphic operations, making it more efficient for realistic workloads. In this section, we outline leveled versions of the BFV, BGV, and CKKS schemes. The core idea is to carry out computations at decreasing modulus levels: perform a fixed number of operations at a given modulus, then reduce both the modulus and the noise to enable further computation.

Let $q_\ell > q_{\ell-1} > \ldots > q_0 > 1$ be distinct primes, and define

$$Q_i = \prod_{j=0}^{i} q_j, \quad 0 \le i \le \ell.$$

We refer to Q_i as the *modulus at level i* or simply the *level-i modulus*. In Section 4.1, we describe how to select each q_i so that a budgeted operation, called a *depth-1 multiplication*, can be performed at level Q_i. Specifically, if the input ciphertexts at level Q_i have noise bounded by ρ, then the resulting ciphertext – after multiplication and modulus switching to Q_{i-1} – continues to maintain the same noise bound ρ.

Section 4.2 details how ciphertext operations are performed in the RNS. By the Chinese remainder theorem, any polynomial $a \in R_{n,Q_\ell}$ can be represented as follows:

$$[a]_{\mathcal{B}} = (a^{(0)}, a^{(1)}, \ldots, a^{(\ell)}),$$

where $a^{(i)} \coloneqq a \bmod q_i$, and $\mathcal{B} = (q_0, q_1, \ldots, q_\ell)$ is called the *modulus basis* (or simply the *basis*). This representation is referred to as the RNS form of a.

All ciphertexts, public keys, and evaluation keys are stored in RNS form with respect to appropriate modulus bases. A key advantage of RNS is that addition and multiplication of polynomials can be performed *component-wise*, independently across the q_i. However, operations such as modulus reduction and relinearization are more involved. In Section 4.2, we describe how these operations are implemented in RNS for the BFV, BGV, and CKKS schemes, and we present the associated noise bounds.

4.1 Budgeted operations at each level

For a collection of ciphertexts, we want to know how much homomorphic computation we can perform before ciphertext noise becomes too big to the point that no further computation can be performed. To do this, we introduce the concept of a depth-1 multiplication computation.

Definition 4.1. (Depth-1 multiplication) Suppose we have a collection of messages. For fixed k_1 and k_2, we say that we can perform a depth-1 multiplication if we can perform $2k_2$ groups of $k_1 - 1$ additions, followed by one round of k_2 multiplications, followed by $k_2 - 1$ additions.

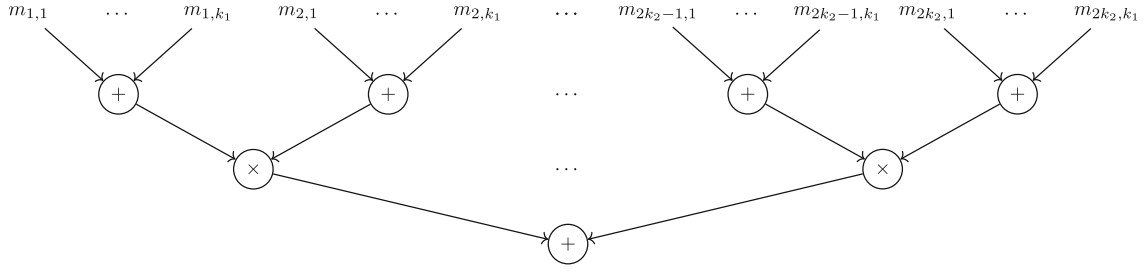


Figure 1: Plaintext depth-1 multiplication.

Figure 1 shows an arbitrary depth-1 multiplication with $2k_2k_1$ plaintexts, where $m_{j,k}$ is a plaintext for each $j = 1,\dots, 2k_2$, $k = 1,\dots, k_1$. Our goal to derive a bound on q_i so that one can compute depth-1 multiplication homomorphically at each level i. To perform a depth-1 multiplication homomorphically for BFV, BGV, and CKKS, we introduce Algorithm 16.

Algorithm 16. Depth-1 multiplication

	Depth1($\mathrm{ct}_{j,k}$, ek_i, Q_i, Q_{i-1})
Input:	$\mathrm{ct}_{j,k} \in R^2_{n,Q_i}$, $j = 1,\dots, 2k_2$, $k = 1,\dots, k_1$ ciphertexts,
	$\mathrm{ek}_i \in R^2_{n,Q_i}$ evaluation key at level i,
	$Q_i \in \mathbb{N}$ integer modulus,
	$Q_{i-1} \in \mathbb{N}$ integer modulus with $Q_i = q_iQ_{i-1}$.
Output:	$\mathrm{ct} \in R^2_{n,Q_{i-1}}$.
Step 1.	For j from 1 to $2k_2$ do
	$\mathrm{ct}_j \coloneqq \mathrm{Linearcombo}(\mathrm{ct}_{j,1}, \dots, \mathrm{ct}_{j,k_1}, 1, \dots, 1)$.
Step 2.	Initialize $\mathrm{ct} \coloneqq (0,0,0)$.
	For j from 1 to k_2 do
	$\mathrm{ct} \coloneqq \mathrm{ct} + \mathrm{Multiply}(\mathrm{ct}_{2j-1}, \mathrm{ct}_{2j})$.
Step 3.	Compute $\mathrm{ct} \coloneqq \mathrm{Relinearize}(\mathrm{ct}, \mathrm{ek}_i)$.
Step 4.	Compute $\mathrm{ct} \coloneqq \mathrm{Modreduce}(Q_i, Q_{i-1}, \mathrm{ct})$.
Step 5.	Return ct.

In Algorithm 16, we remark that Multiply, Relinearize, and Modreduce call the respective algorithms for the inputted ciphertext type. For example, if each $\mathrm{ct}_{j,k}$ is a BFV ciphertext, Algorithm 16 will use BFV.Multiply, BFV.Relinearize, and BFV.Modreduce, while Linearcombo is identical for all three ciphertext types. Note that $\mathrm{ek}_i \in R^2_{n,p_1Q_i}$ is assumed to match the ciphertext type of the $\mathrm{ct}_{j,k}$'s. Our relinearization takes place after summing together our ct_j's obtained from Step 2, which are each a polynomial triple. This slightly improves our bounds, and is better from a computational perspective since we are only running Relinearize once in Algorithm 16.

To guarantee the amount of computation we can perform, we want to choose q_i so that the output of Algorithm 16 is always a ciphertext with noise bounded by ρ when all the $\mathrm{ct}_{j,k}$ inputs have noise bounded by ρ. The precise bound on q_i is presented in Lemmas 4.1 and 4.2 for BFV and BGV, respectively.

Lemma 4.1. *For any* $1 \le i \le \ell$, *suppose* $q_i > 9k_1k_2tn^2$ *and* $\delta_R \ge 16$. *Then, for a collection of BFV ciphertexts at level* i *all with noise bounded by* ρ, *the output of Algorithm 16 is a BFV ciphertext at level* $i - 1$ *with noise bounded by* ρ.

Lemma 4.2. *For any* $1 \le i \le \ell$, *suppose* $q_i > 4k_1^2k_2tn^2$ *and* $\delta_R \ge 16$. *Then, for a collection of BGV ciphertexts at level* i *all with noise bounded by* ρ, *the output of Algorithm 16 is a BGV ciphertext at level* $i - 1$ *with noise bounded by* ρ.

For a similar depth-1 result in CKKS, we must use caution when finding conditions for q_i. The reason for this is that in CKKS, we do not have much flexibility to choose q_i. For the standard scheme, it is always assumed that $q_i = \Delta$ for each $i \neq 0$. Thus, the best that we can do is to bound the error in general after computing the depth-1 algorithm. We cannot force the error down within a constant after rescaling without the assumption that $q_i \gg \Delta$, which clearly contradicts the size of Δ needed in CKKS. We give the result on bounding CKKS noise below in Lemma 4.3.

Lemma 4.3. *Let* i *be such that* $1 \le i \le \ell$, *and suppose that* $n^2 \le \Delta$ *and* $q_i = \Delta$. *Suppose we have a collection of CKKS ciphertexts at level* i *all with noise bounded by* E. *Furthermore, suppose that the corresponding messages* $z_j \in \mathbb{C}^{n/2}$ *each satisfy* $||z_j||_\infty \le Z$. *Then Algorithm 16 results in a CKKS ciphertext at level* $i - 1$ *with noise bounded by* $2k_1k_2nEZ + \frac{k_1k_2En}{\Delta} + \frac{k_1^2k_2E^2n}{\Delta} + \frac{1}{8}$.

One special case of a depth-1 multiplication is the inner product of vectors. That is, given vectors of messages $\mathbf{m} = (m_1, \ldots, m_k) \in R^k_{n,t}$ and $\mathbf{m}' = (m_1', \ldots, m_k') \in R^k_{n,t}$, we want to compute $\langle \mathbf{m}, \mathbf{m}' \rangle \in R_{n,t}$ homomorphically. This can be thought of as one round of k products between the corresponding ciphertexts of $m_1, \ldots, m_k$ and $m_1', \ldots, m_k'$, followed by $k - 1$ additions to sum them together. This is simply a depth-1 multiplication with $k_1 = 1$ and $k_2 = k$. Alternatively, a depth-1 multiplication with $k_1 = k$ and $k_2 = 1$ allows for k ciphertexts to be added together for two separate groups, followed by a single multiplication between the two sums. We argue that our proposed model allows for some more flexibility from a theoretical perspective, as we provide for additions both before and after multiplication at each modulus level.

Remark. Algorithm 16 also works for groups of ciphertext inputs of size less than k_1 with arbitrary linear combinations in Step 1. That is, we can compute

$$\mathrm{ct}_j \coloneqq \mathrm{Linearcombo}(\mathrm{ct}_{j,1}, \ldots, \mathrm{ct}_{j,k_j'}, \alpha_{j,1}, \ldots, \alpha_{j,k_j'})$$

so long as $k_j' \le k_1$ for each j. Furthermore, if for each j, we have

$$\sum_{\omega=1}^{k_j'} |\alpha_{j,\omega}| \le k_1,$$

then Lemmas 4.1, 4.2, and 4.3 still apply.

4.2 Operations in the residue number system

Implementations of homomorphic encryption [28,31–33] take advantage of the RNS variants of schemes [19,21–23]. In our modified leveled homomorphic schemes, we would require that each q_i be chosen coprime to one another in order to use the Chinese remainder theorem. Additions and multiplications are computed componentwise (except for BFV multiplication), which provides major computational advantage over computation modulo large integers. However, algorithms for modulus reductions and relinearization need to be modified to avoid operations in large integers.

Basis conversion in RNS. Suppose $q = q_0 \dots q_{k-1}$ and $p = q_k \dots q_{k+\ell-1}$, where $q_0,\dots, q_{k-1}, q_k \dots, q_{k+\ell-1}$ are distinct primes. Denote

$$\mathcal{B} = (q_0, \dots, q_{k-1}), \quad C = (q_k, \dots, q_{k+\ell-1})$$

as two arbitrary ordered sets which we call *bases*. For an element $a \in R_{n,q}$, we denote $[a]_{\mathcal{B}} \in \prod_{j=0}^{k-1} R_{n,q_j}$ as the vector of CRT components of a in basis $\mathcal{B}$. That is,

$$[a]_{\mathcal{B}} = (a^{(0)}, a^{(1)}, \dots, a^{(k-1)}) = (a^{(j)})_{0 \le j \le k-1}$$

where $a^{(j)} \coloneqq a \bmod q_j$ for $0 \le j < k$. We need to compute $a \bmod p$, that is, $[a]_C$. To do this, let $\hat{q}_j = q/q_j \in \mathbb{Z}$ and $r_j = \hat{q}_j^{-1} a^{(j)} \bmod q_j$ for $0 \le j \le k-1$, where $\|r_j\|_\infty \le q_j/2$. Let

$$\tilde{a} = \sum_{j=0}^{k-1} \hat{q}_j r_j.$$

One can check that $\tilde{a} \equiv a^{(j)} \pmod{q_j}$ for $0 \le j \le k-1$, hence $\tilde{a} \equiv a \pmod{q}$. Then one can compute $\tilde{a} \bmod q_j$ for $k \le j \le k + \ell - 1$ to get $[\tilde{a}]_C$. This yields Algorithm 17 below from [21–23].

Algorithm 17. Fast basis conversion

	$\mathrm{Conv}([a]_{\mathcal{B}}, \mathcal{B}, C)$
Input:	$\mathcal{B} = (q_0, \dots, q_{k-1})$ with $q = q_0 \dots q_{k-1}$, $C = (q_k, \dots, q_{k+\ell-1})$ with $p = q_k \dots q_{k+\ell-1}$, $[a]_{\mathcal{B}} = (a^{(0)}, \dots, a^{(k-1)})$, RNS representation of $a \in R_{n,q}$ in basis $\mathcal{B}$.
Output:	$[\tilde{a}]_C = (\tilde{a}^{(0)}, \dots, \tilde{a}^{(\ell-1)})$, RNS representation of $\tilde{a} \in R_{n,p}$ in basis C.
Step 1.	For $0 \le i \le k-1$, compute $r_i \coloneqq [a^{(i)} \cdot \hat{q}_i^{-1}]_{q_i}$.
Step 2.	For $0 \le i \le \ell - 1$, compute $\tilde{a}^{(i)} \coloneqq \left[\sum_{j=0}^{k-1} \hat{q}_j \cdot r_j\right]_{q_{k+i}}$.
Step 3.	Return $[\tilde{a}]_C = (\tilde{a}^{(0)}, \dots, \tilde{a}^{(\ell-1)})$.

Lemma 4.4. ([21–23]) *Suppose the input of Algorithm 17 is the RNS representation in basis $\mathcal{B}$ of an element $a \in R_{n,q}$. Then, the output $[\tilde{a}]_C$ is the RNS representation in basis C of an element $\tilde{a} \in R_{n,p}$ satisfying*

$$\tilde{a} = a + q \cdot e$$

for some $e \in R_n$ satisfying $\|\tilde{a}\|_\infty \le q \cdot k/2$ and $\|e\|_\infty \le k/2$.

We note that the bound on e follows from the fact that $\|\tilde{a}\|_\infty \le qk/2$. This means that $\tilde{a}$ is only an approximation of a. There are other fast basis conversions in the literature, which give an exact switch (e.g., see [22]). For our purposes in the analysis of relinearization error, the approximate switching in Algorithm 17 will suffice.

Modulus reduction in RNS. Let p be a factor of Q, say $Q = qp$. For any polynomial $a \in R_{n,Q}$, we need to compute the rounding:

$$\left\lfloor \frac{q \cdot a}{Q} \right\rceil = \left\lfloor \frac{a}{p} \right\rceil \in R_{n,q}.$$

Suppose $q = q_0 \dots q_{k-1}$ and $p = q_k \dots q_{k+\ell-1}$ where $q_0, \dots, q_{k-1}, q_k, \dots, q_{k+\ell-1}$ are distinct primes. In RNS, a is represented as $(a^{(0)}, \dots, a^{(k+\ell-1)})$, where

$$a \equiv a^{(i)} \pmod{q_i}, \quad 0 \le i \le k + \ell - 1. \tag{4}$$

By the Chinese remainder theorem, the solution a to equation (4) is unique modulo Q. Note that, for any two polynomials a and b with $a \equiv b \pmod{Q}$, we have

$$\left\lfloor \frac{q \cdot a}{Q} \right\rceil \equiv \left\lfloor \frac{q \cdot b}{Q} \right\rceil \pmod{q}.$$

This is true even if q is not a factor of Q. Hence, we can use any solution a to (4) in the rounding.

Define $\hat{Q}_i = Q/q_i$ and

$$r_i \coloneqq \hat{Q}_i^{-1} a^{(i)} \bmod q_i, \quad 0 \le i \le k + \ell - 1,$$

where the coefficients of r_i are bounded by $q_i/2$. Then a solution of (4) is

$$a = \sum_{i=0}^{k+\ell-1} \hat{Q}_i \, r_i = p \sum_{i=0}^{k-1} \frac{q}{q_i} r_i + q \sum_{i=k}^{k+\ell-1} \frac{p}{q_i} r_i.$$

Note that

$$\frac{a}{p} = \sum_{i=0}^{k-1} \frac{q}{q_i} r_i + \sum_{i=k}^{k+\ell-1} \frac{q}{q_i} r_i.$$

The first sum has integer coefficients, and we only need to round the second sum. Let

$$w = \left\lfloor \sum_{i=k}^{k+\ell-1} \frac{q}{q_i} r_i \right\rceil.$$

Then

$$\frac{a}{p} = \sum_{i=0}^{k-1} \frac{q}{q_i} r_i + w + e, \tag{5}$$

where $e \in \mathbb{R}[x]/(\phi(x))$ with $\|e\|_\infty \le 1/2$. Also, note that, for $0 \le j \le k - 1$,

$$\sum_{i=0}^{k-1} \frac{q}{q_i} r_i \equiv p^{-1} a^{(j)} \bmod q_j.$$

This gives us the Algorithm 18 that matches Algorithm 1.

Algorithm 18. RNS BFV modulus reduction (v1)

RNS.BFV.Modreduce($[a]_{\mathcal{D}}, \mathcal{D}, \mathcal{B}$)

Input: $\mathcal{D} = (q_0, \dots, q_{k+\ell-1})$,
$\mathcal{B} = (q_0, \dots, q_{k-1})$ with $q = q_0 \dots q_{k-1}$ and $p = q_k \dots q_{k+\ell-1}$,
$[a]_{\mathcal{D}} = (a^{(0)}, \dots, a^{(k+\ell-1)})$, RNS representation of $a \in R_{n,Q}$ in basis $\mathcal{D}$.

Output: $[b]_{\mathcal{B}} = (b^{(0)}, \dots, b^{(k-1)})$, RNS representation of $b \in R_{n,q}$ in basis $\mathcal{B}$.

Step 1. For $k \le i \le k + \ell - 1$ and $\hat{p}_i = p/q_i$, compute
$r_i \coloneqq [\hat{p}_i^{-1} \cdot a^{(i)}]_{q_i}$.

Step 2. Compute
$w \coloneqq \left\lfloor \sum_{i=k}^{k+\ell-1} \frac{q}{q_i} \cdot r_i \right\rceil$ in R_n.

Step 3. For $0 \le j \le k-1$, compute
$b^{(j)} \coloneqq [p^{-1}a^{(j)} + w]_{q_j}$.

Step 4. Return $[b]_{\mathcal{B}} = (b^{(0)}, \ldots, b^{(k-1)}) \in \prod_{i=0}^{k-1} R_{n,q_i}$.

The w above gives a rounding error at most 1/2, however, it might be too expensive to compute, as its coefficients are too large. Next, we derive a faster rounding method, with a slightly larger rounding error. Let $\hat{p}_i = p/q_i$ and $v_i \coloneqq \hat{p}_i^{-1} a^{(i)} \bmod q_i$ with $\|v_i\|_\infty \le p/2$ for $k \le i \le k + \ell - 1$. Then

$$v = \sum_{i=k}^{k+\ell-1} \hat{p}_i v_i$$

satisfies $v \equiv a \pmod{p}$ and $\|v\|_\infty \le (p\ell)/2$. Let $u = \frac{a-v}{p}$, which has integer coefficients. Then

$$\frac{a}{p} = u + e, \tag{6}$$

where $e = v/p \in \mathbb{R}[x]/(\phi(x))$ with $\|e\|_\infty \le \ell/2$. In RNS, $u \bmod q$ can be obtained by first computing $v \bmod q_j$, $0 \le j \le k-1$, via basis conversion from p to q. Algorithm 19 describes the procedure, while Lemma 4.5 shows the noise bound. We exclude the proof of Lemma 4.5 from our Appendix, as it is clear from the previous discussion.

Algorithm 19. RNS BFV modulus reduction (v2)

RNS.BFV.Modreduce($[a]_{\mathcal{D}}, \mathcal{D}, \mathcal{B}$)

Input: $\mathcal{D} = (q_0, \ldots, q_{k+\ell-1})$,
$\mathcal{B} = (q_0, \ldots, q_{k-1})$ with $q = q_0 \ldots q_{k-1}$ and $p = q_k \ldots q_{k+\ell-1}$,
$[a]_{\mathcal{D}} = (a^{(0)}, \ldots, a^{(k+\ell-1)})$, RNS representation of $a \in R_{n,Q}$ in basis $\mathcal{D}$.

Output: $[b]_{\mathcal{B}} = (b^{(0)}, \ldots, b^{(k-1)})$, RNS representation of $b \in R_{n,q}$ in basis $\mathcal{B}$.

Step 1. Let $C = \mathcal{D} \backslash \mathcal{B} = (q_k, \ldots, q_{k+\ell-1})$. Compute
$(v^{(0)}, \ldots, v^{(k-1)}) \coloneqq \mathrm{Conv}((a^{(k)}, \ldots, a^{(k+\ell-1)}), C, \mathcal{B})$.

Step 2. For $0 \le j \le k-1$, compute
$b^{(j)} \coloneqq p^{-1} \cdot (a^{(j)} - v^{(j)}) \bmod q_j$.

Step 3. Return $[b]_{\mathcal{B}} = (b^{(0)}, \ldots, b^{(k-1)}) \in \prod_{i=0}^{k-1} R_{n,q_i}$.

Lemma 4.5. *Let the RNS representation of $a \in R_{n,Q}$ be the input of Algorithm 19 and the RNS representation of $b \in R_{n,q}$ the output. Then,*

$$\frac{a}{p} = b + e$$

for some $e \in \mathbb{R}[x]/(\phi(x))$ with $\|e\|_\infty \le \ell/2$.

Next, we show a BGV modulus reduction in RNS that matches Algorithm 2. When q is a factor of Q, say $Q = qp$, note that $(-a_0 q t^{-1}) \bmod Q$ is the same as $q(-a_0 t^{-1} \bmod p)$. Hence, Algorithm 2 can be simplified as follows:

$$\omega_a \coloneqq [-a_0 t^{-1}]_p \quad \text{and} \quad \omega_b \coloneqq [-b_0 t^{-1}]_p$$
$$a_0' \coloneqq \left[\frac{a_0 + t\ \omega_a}{p}\right]_q \quad \text{and} \quad b_0' \coloneqq \left[\frac{b_0 + t\ \omega_b}{p}\right]_q.$$

Algorithm 20 describes the above procedure for reduction of one polynomial, while Lemma 4.6 shows the noise bound. We exclude the proof of Lemma 4.6, since it is clear from the discussion and the proof of Lemma 2.3.

Algorithm 20. RNS BGV modulus reduction

RNS.BGV.Modreduce($[a]_{\mathcal{D}}, \mathcal{D}, \mathcal{B}$)

Input: $\mathcal{D} = (q_0, q_1, \ldots, q_{k+\ell-1})$,
$\mathcal{B} = (q_0, \ldots, q_{k-1})$ and $p = q_k \cdots q_{k+\ell-1}$,
$[a]_{\mathcal{D}} = (a^{(0)}, \ldots, a^{(k+\ell-1)})$, RNS representation of $a \in R_{n,Q}$ in basis $\mathcal{D}$.

Output: $[b]_{\mathcal{B}} = (b^{(0)}, \ldots, b^{(k-1)})$, RNS representation of $b \in R_{n,q}$ in basis $\mathcal{B}$.

Step 1. For $k \le i \le k + \ell - 1$, compute
$w^{(i)} \coloneqq [-t^{-1}a^{(i)}]_{q_i}$.

Step 2. Let $C = \mathcal{D} \backslash \mathcal{B} = (q_k, \ldots, q_{k+\ell-1})$. Compute
$(u^{(0)}, \ldots, u^{(k-1)}) \coloneqq \mathrm{Conv}((w^{(k)}, \ldots, w^{(k+\ell-1)}), C, \mathcal{B})$.

Step 3. For $0 \le i \le k - 1$, compute
$b_i \coloneqq [p^{-1}(a^{(i)} + tu^{(i)})]_{q_i}$.

Step 4. Return $(b^{(0)}, \ldots, b^{(k-1)}) \in \prod_{i=0}^{k-1} R_{n,q_i}$.

Lemma 4.6. *Let the RNS representation of $a \in R_{n,Q}$ be the input of Algorithm 20 and the RNS representation of $b \in R_{n,q}$ the output. Then,*

$$\frac{a}{p} = b + t \cdot e$$

for some $e \in \mathbb{R}[x]/(\phi(x))$ with $\|e\|_\infty \le \ell/2$.

For the CKKS rescaling procedure in RNS, we can again use the same procedure as the fast BFV modulus reduction in Algorithm 19. Said otherwise,

$$\text{RNS.CKKS.Modreduce} = \text{RNS.BFV.Modreduce}.$$

Note here that we specifically use (v2) of the RNS BFV Modulus reduction for RNS CKKS. Likewise, we can also use Lemma 4.5 for RNS CKKS when discussing the noise bound after performing RNS.CKKS.Modreduce.

Relinearization in RNS. For the rest of this section, fix $Q_\ell = q_0 \cdots q_\ell$ and $P = p_0 \cdots p_{k-1}$ with $q_0 \cdots q_\ell, p_0 \cdots p_{k-1}$ all coprime. For $0 \le i \le \ell$, let $Q_i = q_0 \cdots q_i$ and fix the ordered bases

$$\mathcal{B} = (q_0, \ldots, q_i),$$
$$C = (p_0, \ldots, p_{k-1}),$$
$$\mathcal{D} = \mathcal{B} \cup C = (q_0, \ldots, q_i, p_0, \ldots, p_{k-1}).$$

The goal of our relinearization algorithms in RNS will again be to closely match the previously outlined relinearizations for the classic variants in Algorithms 8 and 13. We first introduce the evaluation key generations for the three schemes, followed by their relinearization procedures. We should note that we only discuss the evaluation key generation here. The generations of other keys (secret and public encryption keys) are similarly RNS versions of Algorithms 3 and 10.

We begin with RNS BFV. The procedure for evaluation key generation is given in Algorithm 21. Observe that this evaluation key generation is the same as the evaluation key generation in Algorithm 3, only computed in RNS.

Algorithm 21. RNS BFV evaluation key generation

RNS.BFV.ek.Keygen(sk, $(q_0, \ldots, q_\ell, p_0, \ldots, p_{k-1})$)

Input: sk = $s \in R_{n,3}$ secret key,
$(q_0, \ldots, q_\ell, p_0, \ldots, p_{k-1})$ full RNS basis.

Output: ek = $(\tilde{k}_0^{(j)}, \tilde{k}_1^{(j)})_{0\le j\le k+\ell} \in \prod_{j=0}^{\ell} R_{n,q_j}^2 \times \prod_{j=0}^{k-1} R_{n,p_j}^2$ evaluation key.

Step 1. Sample $(\tilde{k}_0^{(0)}, \ldots, \tilde{k}_0^{(k+\ell)}) \leftarrow U(\prod_{j=0}^{\ell} R_{n,q_j} \times \prod_{j=0}^{k-1} R_{n,p_j})$ and $\tilde{e} \leftarrow \chi_\rho$.

Step 2. For $0 \le j \le \ell$ compute
$\tilde{k}_1^{(j)} \coloneqq [-\tilde{k}_0^{(j)} s + [P]_{q_j} s^2 + \tilde{e}]_{\phi(x),q_j}$.

Step 3. For $0 \le j \le k-1$ compute
$\tilde{k}_1^{(\ell+1+j)} \coloneqq [-\tilde{k}_0^{(\ell+1+j)} s + \tilde{e}]_{\phi(x),p_j}$.

Step 4. Return ek = $(\tilde{k}_0^{(j)}, \tilde{k}_1^{(j)})_{0\le j\le k+\ell}$.

For the RNS BGV scheme, the evaluation key generation is again a similar approach to the original evaluation key generation from Algorithm 9. Algorithm 22 gives the procedure for RNS BGV.

Algorithm 22. RNS BGV evaluation key generation

RNS.BGV.ek.Keygen(sk, $(q_0, \ldots, q_\ell, p_0, \ldots, p_{k-1})$)

Input: sk = $s \in R_{n,3}$ secret key,
$(q_0, \ldots, q_\ell, p_0, \ldots, p_{k-1})$ full RNS basis.

Output: ek = $(\tilde{k}_0^{(j)}, \tilde{k}_1^{(j)})_{0\le j\le k+\ell} \in \prod_{j=0}^{\ell} R_{n,q_j}^2 \times \prod_{j=0}^{k-1} R_{n,p_j}^2$ evaluation key.

Step 1. Sample $(\tilde{k}_0^{(0)}, \ldots, \tilde{k}_0^{(k+\ell)}) \leftarrow U(\prod_{j=0}^{\ell} R_{n,q_j} \times \prod_{j=0}^{k-1} R_{n,p_j})$ and $\tilde{e} \leftarrow \chi_\rho$.

Step 2. For $0 \le j \le \ell$ compute
$\tilde{k}_1^{(j)} \coloneqq [-\tilde{k}_0^{(j)} s + [P]_{q_j} s^2 + t\tilde{e}]_{\phi(x),q_j}$.

Step 3. For $0 \le j \le k-1$ compute
$\tilde{k}_1^{(\ell+1+j)} \coloneqq [-\tilde{k}_0^{(\ell+1+j)} s + t\tilde{e}]_{\phi(x),p_j}$.

Step 4. Return ek = $(\tilde{k}_0^{(j)}, \tilde{k}_1^{(j)})_{0\le j\le k+\ell}$.

For the RNS CKKS scheme, the evaluation key generation is exactly the same as the evaluation key generation for RNS BFV:

$$\text{RNS.CKKS.ek.Keygen} = \text{RNS.BFV.ek.Keygen}.$$

We are now ready to discuss the full RNS realinearization procedures. In these algorithms, we assume that we have obtained a vector of components $(c_0^{(j)}, c_1^{(j)}, c_2^{(j)})_{0\le j\le i} \in \prod_{j=0}^{i} R_{n,q_j}^3$ which are the RNS representation of some $(c_0, c_1, c_2) \in R_{n,Q_i}^3$ that is obtained after initial RNS multiplication for BFV, BGV, or CKKS. The initial RNS multiplication operations for BGV and CKKS are simply computed componentwise modulo the q_j's. The procedure for BFV is more complicated. Specifically, Step 2 of Algorithm 7 in which we must scale down components without modular reduction is difficult to perform in RNS representation. We refer the reader to the previous studies [22,23] for details for the details on the initial RNS BFV multiplication.

For all three schemes, the RNS linearization procedure is given in Algorithm 23. Here, RNS.Modreduce is the RNS modulus reduction procedure for the chosen scheme, and ek is the corresponding evaluation key for that scheme. For instance, if we choose to run RNS relinearization for BGV, we use RNS.BGV.Modreduce

in Step 4 with the corresponding evaluation key ek for BGV. We introduce Lemmas 4.7, 4.8, and 4.9 to prove correctness of the algorithm and our noise bound for RNS variant of BFV, BGV, and CKKS, respectively.

Algorithm 23. RNS relinearization

RNS.Relinearize($[(c_0, c_1, c_2)]_{\mathcal{B}}$, $[\mathrm{ek}]_{\mathcal{D}}$)

Input: $[(c_0, c_1, c_2)]_{\mathcal{B}} = (c_0^{(j)}, c_1^{(j)}, c_2^{(j)})_{0\le j\le i} \in \prod_{j=0}^{i} R_{n,q_j}^3$,

$[\mathrm{ek}]_{\mathcal{D}} = (\tilde{k}_0^{(j)}, \tilde{k}_1^{(j)})_{0\le j\le i+k} \in \prod_{j=0}^{i} R_{n,q_j}^2 \times \prod_{j=0}^{k-1} R_{n,p_j}^2$ evaluation key.

Output: $[\mathrm{ct}]_{\mathcal{B}} = (a^{(j)}, b^{(j)})_{0\le j\le i}$ RNS ciphertext.

Step 1. Compute $(\tilde{c}_2^{(0)}, \ldots, \tilde{c}_2^{(k-1)}) \coloneqq \mathrm{Conv}((c_2^{(0)}, \ldots, c_2^{(i)}), \mathcal{B}, C)$.

Step 2. For $0 \le j \le i$ compute

$\hat{a}^{(j)} \coloneqq [c_2^{(j)} \tilde{k}_0^{(j)}]_{\phi(x),q_j}$,

$\hat{b}^{(j)} \coloneqq [c_2^{(j)} \tilde{k}_1^{(j)}]_{\phi(x),q_j}$.

Step 3. For $0 \le j \le k-1$ compute

$\hat{a}^{(i+1+j)} \coloneqq [\tilde{c}_2^{(j)} \tilde{k}_0^{(i+1+j)}]_{\phi(x),p_j}$,

$\hat{b}^{(i+1+j)} \coloneqq [\tilde{c}_2^{(j)} \tilde{k}_1^{(i+1+j)}]_{\phi(x),p_j}$.

Step 4. Compute

$(\hat{c}_1^{(0)}, \ldots, \hat{c}_1^{(i)}) \coloneqq \mathrm{RNS.Modreduce}((\hat{a}_0^{(0)}, \ldots, \hat{a}_0^{(i+k)}), \mathcal{D}, \mathcal{B})$,

$(\hat{c}_0^{(0)}, \ldots, \hat{c}_0^{(i)}) \coloneqq \mathrm{RNS.Modreduce}((\hat{b}_0^{(0)}, \ldots, \hat{b}_0^{(i+k)}), \mathcal{D}, \mathcal{B})$.

Step 5. For $0 \le j \le i$ compute

$a^{(j)} \coloneqq [c_1^{(j)} + \hat{c}_1^{(j)}]_{q_j}$,

$b^{(j)} \coloneqq [c_0^{(j)} + \hat{c}_0^{(j)}]_{q_j}$.

Step 6. Return $\mathrm{ct} = (a^{(j)}, b^{(j)})_{0\le j\le i}$.

Lemma 4.7. *Let* ct *be the output of Algorithm 23 and suppose the input* $(c_0^{(j)}, c_1^{(j)}, c_2^{(j)})_{0\le j\le i}$ *is the RNS representation of some* $(c_0, c_1, c_2) \in R_{n,Q_i}^3$ *satisfying*

$$c_0 + c_1 s + c_2 s^2 \equiv D_{Q_i}[m_0 m_1]_{\phi(x),t} + e' \bmod (\phi(x), Q_i)$$

for $\|e'\|_\infty \le E$. *If* $P \ge 6Q_i$, $\delta_R \ge 16$, *and* $k > i$, *then* ct *is the RNS representation of a BFV ciphertext with noise bounded by* $E + \frac{1}{8}\delta_R^2 k$.

Lemma 4.8. *Let* ct *be the output of Algorithm 23 and suppose the input* $(c_0^{(j)}, c_1^{(j)}, c_2^{(j)})_{0\le j\le i}$ *is the RNS representation of some* $(c_0, c_1, c_2) \in R_{n,Q_i}^3$ *satisfying*

$$c_0 + c_1 s + c_2 s^2 \equiv [m_0 m_1]_{\phi(x),t} + te' \bmod (\phi(x), Q_i)$$

for $\|e'\|_\infty \le E$. *If* $P \ge 6Q_i$, $\delta_R \ge 16$, *and* $k > i$, *then* ct *is the RNS representation of a BGV ciphertext with noise bounded by* $E + \frac{1}{8}\delta_R^2 k$.

Lemma 4.9. *Let* ct *be the output of Algorithm 23 and suppose the input* $(c_0^{(j)}, c_1^{(j)}, c_2^{(j)})_{0\le j\le i}$ *is the RNS representation of some* $(c_0, c_1, c_2) \in R_{n,Q_i}^3$ *satisfying*

$$c_0 + c_1 s + c_2 s^2 \equiv m_0 m_1 + e' \bmod (\phi(x), Q_i),$$

for $\|e'\|_\infty \le E$. *If* $P \ge 6Q_i$, $\delta_R \ge 16$, *and* $k > i$, *then* ct *is the RNS representation of a CKKS ciphertext with noise bounded by* $E + \frac{1}{8}\delta_R^2 k$.

In addition to the relinearization technique we opt for, we should also mention that there exists versions of the alternate relinearization technique from Section 3.1 in RNS [22,23]. In the RNS version, the element c_2 is essentially expanded into a bit decomposition twice: an expansion in the q_j's first, and then another expansion in a fixed base B for each component corresponding to each q_j. Though this technique is certainly viable, we opt for the outlined technique due to the smaller size of ek. In practice, both relinearization techniques are used together in a method known as *hybrid key switching* [34]. In practice, the noise bound for hybrid key switching is quite similar to what we have outlined in Lemmas 4.7, 4.8, and 4.9. The bound for hybrid key switching ranges from about $E + \frac{3}{8}\delta_R^2 k$ to $E + \frac{10}{8}\delta_R^2 k$ with our approach, using the noise bound from the study by Kim et al. [16] and practical estimates of d_{num} from [9]. We refer the reader to [16,34] for more details on hybrid key switching.

5 Lattices, security, and attacks

The security of homomorphic encryption schemes is based on the LWE problem over finite fields, which can be reduced to lattice problems. In this section, we will give an overview of these lattice problems as well as various attacks on LWE. As this article is more focused on noise reduction in homomorphic encryption schemes, we only provide a brief overview of security and attacks. For a more in-depth discussion on security, we refer the reader to various sources [24,30,35]. Decision-RLWE can be shown to be as hard as many worst-case lattice problems [36]. There is also a brief mention of security reductions from RLWE to LWE in [30]. We will discuss attacks on classic LWE rather than RLWE, since RLWE problems can be easily converted into LWE problems. Furthermore, all the best attack algorithms are for LWE instead of RLWE.

5.1 Lattices and lattice problems

Let $m \ge 1$. A subset $\Lambda \subseteq \mathbb{R}^m$ is called a lattice if Λ is a discrete additive subgroup of $\mathbb{R}^m$ and each point of Λ is isolated (i.e., no points in Λ are arbitrarily close to each other). In general, a lattice can be generated in the following way: given a matrix $B = (b_1, \ldots, b_n) \in \mathbb{R}^{m\times n}$ with independent columns, a lattice can be defined via

$$\Lambda = \{y \in \mathbb{R}^m : y = Bx, x \in \mathbb{Z}^n\}.$$

Here, $\Lambda \subseteq \mathbb{R}^m$ is an n-dimensional lattice. We call B a lattice basis. For a lattice Λ defined by lattice basis B, the volume $\text{vol}(\Lambda)$ is defined as follows:

$$\text{vol}(\Lambda) = \sqrt{\det(B^T B)},$$

which can be proved to be independent of the choice of basis. Let

$$\lambda(\Lambda) = \min\{\|x\|_2 : x \in \Lambda, x \ne 0\}.$$

Definitions 5.1, 5.2, and 5.3 describe a few instances of well-studied lattice problems for a lattice Λ [30].

Definition 5.1. (SVP) The shortest vector problem (SVP) is as follows: given a basis B of Λ, find a vector $v \in \Lambda$ such that $\|v\|_2 = \lambda(\Lambda)$.

Definition 5.2. (γ-SVP) The γ-approximate shortest vector problem (γ-SVP) is as follows: given a basis B of Λ, find a nonzero vector $v \in \Lambda$ such that $\|v\|_2 \leq \gamma \cdot \lambda(\Lambda)$.

Definition 5.3. (γ-GapSVP) The γ-gap shortest vector problem (γ-GapSVP) is as follows: given a basis B of Λ and a real number $r > 0$, output "yes" if $\lambda(\Lambda) \leq r$ and output "no" if $\lambda(\Lambda) > \gamma \cdot r$.

These lattice problems are examples of NP-hard problems in the worst case. Regev [24] provides a reduction from an instance of $\gamma-$GapSVP to Decision-LWE [35], meaning that LWE is at least as hard as γ-GapSVP. For potential attacks on LWE based schemes, we will discuss attack strategies outlined in the previous studies [24,30].

q-ary Lattices. In lattice-based cryptography, a class of lattices that is particularly important is q-ary lattices. We say Λ is a q-ary lattice if Λ is a lattice such that $q\mathbb{Z}^m \subseteq \Lambda \subseteq \mathbb{Z}^m$. In particular, given a matrix $A \in \mathbb{Z}^{m\times n}$ of rank n modulo q, the following are q-ary lattices

$$\Lambda_q(A) = \{x \in \mathbb{Z}^m : x \equiv Ay \bmod q \text{ for some } y \in \mathbb{Z}^n\},$$
$$\Lambda_q'(A) = \{x \in \mathbb{Z}^m : x^T A \equiv 0 \bmod q\}.$$

It is also worth mentioning that although the matrix A defines both of these lattices, A is not necessarily a lattice basis for these lattices. To find a lattice basis of $\Lambda_q(A)$, we can perform column operations on A modulo q, permuting the rows if necessary, to obtain a matrix

$$\begin{pmatrix} I_n \\ A_1 \end{pmatrix} \in \mathbb{Z}_q^{m\times n},$$

where I_n is an $n \times n$ identity matrix and $A_1 \in \mathbb{Z}_q^{(m-n)\times n}$. Let

$$B = \begin{pmatrix} I_n & 0 \\ A_1 & qI_{m-n} \end{pmatrix} \in \mathbb{Z}^{m\times m}. \tag{7}$$

Then B has rank m and is a lattice basis for $\Lambda_q(A)$. Since $\Lambda_q(A)$ is a full rank lattice, $\mathrm{vol}(\Lambda_q(A)) = |\det(B)| = q^{m-n}$.

To obtain a lattice basis for $\Lambda_q'(A)$, let A_1 be as mentioned earlier. Note that the solution space for $x^T A = 0 \bmod q$ is spanned by the columns of the matrix

$$\begin{pmatrix} -A_1^T \\ I_{m-n} \end{pmatrix} \in \mathbb{Z}^{m\times(m-n)}.$$

Then, a basis for $\Lambda_q'(A)$ is

$$B' = \begin{pmatrix} -A_1^T & qI_n \\ I_{m-n} & 0 \end{pmatrix} \in \mathbb{Z}^{m\times m}.$$

The volume of this lattice is $\mathrm{vol}(\Lambda_q'(A)) = q^n$.

Gaussian heuristic. By the Gaussian Heuristic, for a lattice Λ of rank m, we expect its shortest vector to be of length

$$\sqrt{\frac{m}{2\pi \cdot \exp(1)}}\,\mathrm{vol}(\Lambda)^{1/m}$$

on average [24,37], where $\exp(1) = 2.7182\ldots$ is the exponential function evaluated at 1. As the lattice $\Lambda_q(A)$ has volume q^{m-n} and A_1 is uniform random in $\mathbb{Z}_q^{(m-n)\times n}$, we can use the Gaussian Heuristic and expect the shortest vector in $\Lambda_q(A)$ to be of length

$$\sqrt{\frac{m}{2\pi \cdot \exp(1)}}\,q^{1-n/m} \tag{8}$$

on average. Similarly for $\Lambda_q'(A)$, we expect its shortest vector to be of length

$$\sqrt{\frac{m}{2\pi \cdot \exp(1)}}\, q^{n/m}$$

on average.

5.2 LWE attack strategies

First, recall the LWE problems outlined in Section 2.2. Fix $s \in \mathbb{Z}_q^n$, which is secret. We sample $e_i \leftarrow \chi(\mathbb{Z})$ from some desired distribution χ such that $\|e_i\|_\infty \le \rho$, where ρ is a desired parameter. Then, we sample a uniform random $a_i \in \mathbb{Z}_q^n$ and calculate b_i via $b_i = [-\langle a_i, s\rangle + e_i]_q$. The ordered pair $(a_i, b_i) \in \mathbb{Z}_q^n \times \mathbb{Z}_q$ is called an *LWE sample*. The Search-LWE problem is to find s given many LWE samples. The decision-LWE problem is given many samples that are either LWE samples or sampled uniform randomly, decide which distributions the samples are drawn from [2,24]. If we sample m times, we can instead think of the LWE samples as the matrix equation

$$b \equiv As + e \bmod q, \tag{9}$$

where $b \in \mathbb{Z}_q^m$ is the vector of b_i's, $A \in \mathbb{Z}_q^{m\times n}$ is the matrix of a_i's, and $e \in \mathbb{Z}_q^m$ is the vector of e_i's.

Dual attacks via SVP. To solve Decision-LWE, we employ a dual attack. Let A be the matrix defined in equation (9). In dual attacks, we wish to find a short vector $v \in \Lambda_q'(A)$. Since $\langle v, b\rangle = \langle v, As + e\rangle = \langle v, e\rangle \bmod q$ and e is short, $\langle v, b\rangle$ is small. If an adversary can find a short vector $v \in \Lambda_q'(A)$, then the adversary can solve the Decision-LWE problem with a fair amount of confidence, since $\langle v, b\rangle$ would likely not be small for true random b. Thus, the attacker can distinguish LWE samples from true random samples with advantage. We refer the reader to [8,24,38,39] for more details on dual attacks and the exact advantages based on sizes of e and v.

Primal attacks via SVP. A common attack strategy for Search-LWE is with SVP. Let A be the matrix defined in equation (9) and B be computed from A as in equation (7). Let

$$\tilde{B} = \begin{pmatrix} B & b \\ 0 & 1 \end{pmatrix} \in \mathbb{Z}^{(m+1)\times(m+1)},$$

and let $\tilde{\Lambda}$ be the lattice defined by $\tilde{B}$, which has volume $\mathrm{vol}(\tilde{\Lambda}) = q^{m-n}$. Then equation (9) means that the vector $\begin{pmatrix} e \\ 1 \end{pmatrix}$ is in $\tilde{\Lambda}$. By the Gaussian Heuristic in equation (8), we expect the shortest vector in the lattice generated by B to be of length about

$$\sqrt{\frac{m}{2\pi \cdot \exp(1)}}\, q^{1-n/m}$$

since the entries of A_1 are uniform random in $\mathbb{Z}_q$. If $\|e\|_2$ is smaller than this, then the shortest vector in $\tilde{\Lambda}$ is likely to be $\begin{pmatrix} e \\ 1 \end{pmatrix}$ with a significant probability. Thus, when e is small, we can solve SVP for the lattice $\tilde{\Lambda}$ to find e, and in turn solve LWE. The error distribution χ is crucial in determining the expected size of e, which we discuss thoroughly in the next subsection. We refer the reader to previous studies [30,39,40] for more details on primal attacks.

Lattice basis reduction algorithms. Several algorithms employ these strategies, as well as others, to solve LWE. Algorithms in practice for solving lattice problems include algorithms such as LLL [41], BKW [42], and BKZ [43], which are discussed thoroughly in previous studies [8] and [24]. We will primarily discuss BKZ, as it seems to currently be the best algorithm for lattice reduction. The basic idea behind BKZ is to solve SVP for sublattices of dimension k, which is known as the *block size* in BKZ.

Let $B = (b_0, \ldots, b_{m-1})$ be a lattice basis for Λ, ordered so that b_0 is the shortest vector in B. Then, there is a constant γ_0 so that

$$\|b_0\|_2 = \gamma_0^m \operatorname{vol}(\Lambda)^{1/m}.$$

We call γ_0^m the Hermite factor and γ_0 the root-Hermite factor of the lattice basis B. The Hermite factor is crucial in determining cost and runtime of lattice reduction algorithms, especially BKZ. For a block size k in BKZ, Chen [44] shows that it is expected that the algorithm output a lattice basis B with γ_0 satisfying

$$\lim_{m\to\infty} \gamma_0 \approx \left(\frac{k}{2\pi \cdot \exp(1)}(\pi k)^{\frac{1}{k}}\right)^{\frac{1}{2(k-1)}}. \tag{10}$$

In practice, the limiting factor from equation (10) is used to estimate γ_0 for a finite dimensional lattice [24]. One can compute a lattice basis B with root-Hermite factor γ_0 with the estimate in equation (10). For ease of analysis, γ_0 is approximated by either $\gamma_0 = k^{\frac{1}{2k}}$ or $\gamma_0 = 2^{\frac{1}{k}}$. Albrecht et al. [24] show that for block sizes $50 \le k \le 250$, $2^{\frac{1}{k}}$ actually approximates the estimate of γ_0 from equation (10) better than $k^{\frac{1}{2k}}$.

Our goal is to obtain a basis with a specific target size for $\|b_0\|_2$ while we are allowed to determine the lattice dimension m, which is the number of LWE samples used. This can be difficult since the relationship between γ_0 and $\|b_0\|_2$ depends on m. In this scenario, one of the earlier estimates for γ_0 (e.g., $\gamma_0 = k^{\frac{1}{2k}}$ or $\gamma_0 = 2^{\frac{1}{k}}$) can be used to determine m based on a chosen block size k. In practice, the best root-Hermite factor γ_0 that can be obtained via lattice reduction algorithms is about $\gamma_0 \approx 1.1011$ to $\gamma_0 \approx 1.1013$ [45], and consistently lower values do not seem currently achievable. For a lattice of volume q^n (e.g., $\Lambda_q'(A)$), the study [46] gives the optimal size of m as

$$m = \sqrt{\frac{n\log(q)}{\log(\gamma_0)}}$$

for use in lattice reduction algorithms to obtain the best result.

The study by van de Pol and Smart [45] introduces a technique which does not rely on first knowing γ_0 to find m. Instead, a security level is first chosen, then the best possible γ_0 obtained from BKZ is found based on the security level, a chosen lattice dimension m, and the underlying error sampling distribution (and in particular, its standard deviation). The homomorphic encryption standard and several implementations use $\chi = D_{\mathbb{Z},\alpha q}$ with $\alpha q = 8$, resulting in a standard deviation of $\sigma \approx 3.2$. In our proposed schemes, χ is a discrete uniform distribution in $[-n, n]$ with standard deviation

$$\sigma = \sqrt{\frac{(2n+1)^2 - 1}{12}} = \sqrt{\frac{4n^2 + 4n}{12}} \approx \frac{n}{\sqrt{3}}.$$

Note that in all these techniques, the dual q-ary lattice $\Lambda_q'(A)$ is used instead of the lattice $\Lambda_q(A)$. Though, $\Lambda_q(A)$ and $\Lambda_q'(A)$ are equivalent to one another up to normalization (see, e.g., [46]).

After deciding on the block size k and optimal dimension m, one can estimate the cost of BKZ as follows. For algorithms to solve SVP for lattices of rank k, the fastest known classical algorithm (using sieving) [47] runs in time $2^{0.292k+o(k)}$. The fastest known quantum algorithm [48] runs in time $2^{0.265k+o(k)}$. With BKZ, we can have up to $8m$ calls to an oracle solving SVP using a sieving algorithm, meaning that the total costs for BKZ we can expect are $8m \cdot 2^{0.292k+o(k)}$ classically and $8m \cdot 2^{0.265k+o(k)}$ quantumly [8].

In summary, to estimate the total cost of attack, an appropriate block size k and lattice dimension m must first be determined based on parameters n, q, and the standard deviation of the chosen error sampling distribution χ. After finding this expected size of e, we can calculate the root-Hermite factor γ_0, which then allows us to determine the block size k and the total cost of BKZ. Though a thorough security analysis is still to be conducted on our strategy for picking leveled homomorphic encryption parameters, extensive research and estimates have been performed on the best known algorithms for solving LWE as briefly shown above. We point the reader to previous studies [8,9,24,49] for a more complete discussion and further references on security.

6 Conclusion

We have presented a detailed mathematical foundation for the BGV, BFV, and CKKS homomorphic encryption schemes, aligning our work with the functionalities proposed in recent homomorphic encryption standards. By providing protocol algorithms and correctness proofs, we have ensured that these schemes are not only theoretically sound but also practical for implementation. Our proposed improvements, particularly in noise management and leveled homomorphic computation, enhance the efficiency and applicability of these schemes by reducing ciphertext expansion and storage requirements. In future works, we plan to further analyze the impacts of these noise bounds on precision accuracy in CKKS.

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Appendix

A Proofs of Lemmas

Proof of Lemma 2.1. By assumption, $b_0 \equiv -a_0 s + e_0 \bmod (\phi(x), Q)$. Therefore, there exists $r \in R_n$ such that

$$b \equiv -a_0 s + e_0 + Qr \bmod \phi(x).$$

Since $a_0' = \left\lfloor \frac{qa_0}{Q} \right\rceil$ and $b_0' = \left\lfloor \frac{qb_0}{Q} \right\rceil$, there are polynomials $\varepsilon_1, \varepsilon_2 \in \mathbb{R}[x]/(\phi(x))$ such that $a_0' = \frac{qa_0}{Q} - \varepsilon_1$ and $b_0' = \frac{qb_0}{Q} - \varepsilon_2$ with $||\varepsilon_1||_\infty, ||\varepsilon_2||_\infty \le 1/2$. Then,

$$\begin{aligned} b_0' &= \frac{q}{Q} b_0 - \varepsilon_2 \\ &\equiv -\frac{q}{Q} a_0 s + \frac{q}{Q} e_0 + qr - \varepsilon_2 \bmod \phi(x) \\ &\equiv -a_0' s + \frac{q}{Q} e_0 + qr - \varepsilon_2 - \varepsilon_1 s \bmod \phi(x). \end{aligned}$$

Let $e_0' = \frac{q}{Q} e_0 - \varepsilon_2 - \varepsilon_1 s$. Then, $b_0' \equiv -a_0' s + e_0' \bmod (\phi(x), q)$. Note that $||e_0'||_\infty \le \frac{q}{Q} E + \frac{\delta_R ||s||_\infty + 1}{2}$. By assumption $Q/q > \frac{2E}{\delta_R ||s||_\infty - 1}$, so $||e_0'||_\infty < \delta_R ||s||_\infty$. □

Proof of Lemma 2.2. By assumption, we first note that $(a_0, b_0) \in R_{n,Q}^2$ satisfies $b_0 \equiv -a_0 s + D_Q m_0 + e_0 \bmod (\phi(x), Q)$. Therefore, there is some $r_Q \in R_n$ such that $b_0 + a_0 s \equiv D_Q m_0 + e_0 + Q r_Q \bmod \phi(x)$. Let $\varepsilon_Q = Q/t - D_Q$, $\varepsilon_q = q/t - D_q$, $\varepsilon_1 = qa_0/Q - a_0'$, and $\varepsilon_2 = qb_0/Q - b_0'$. Then,

$$b_0' = \frac{qb_0}{Q} - \varepsilon_2 \equiv -\frac{qa_0 s}{Q} + \frac{qD_Q}{Q} m_0 + \frac{qe_0}{Q} - \varepsilon_2 + qr_Q \bmod \phi(x).$$

Note that as $D_Q = Q/t - \varepsilon_Q$, we have that $qD_Q/Q = q/t - q\varepsilon_Q/Q$. Since $q/t = D_q + \varepsilon_q$, we have $qD_Q/Q = D_q + \varepsilon_q - q\varepsilon_Q/Q$. Therefore,

$$\begin{aligned} b_0' &\equiv -\frac{qa_0 s}{Q} + \frac{qD_Q}{Q} m_0 + \frac{qe_0}{Q} - \varepsilon_2 + qr_Q \bmod \phi(x) \\ &\equiv -a_0' s - \varepsilon_1 s + D_q m_0 + (\varepsilon_q - \frac{q\varepsilon_Q}{Q}) m_0 + \frac{qe_0}{Q} - \varepsilon_2 + qr_Q \bmod \phi(x). \end{aligned}$$

Let $e_0' = \frac{qe_0}{Q} + (\varepsilon_q - \frac{q\varepsilon_Q}{Q}) m_0 - \varepsilon_2 - \varepsilon_1 s$. Then, $b_0' + a_0' s \equiv D_q m_0 + e_0' \bmod (\phi(x), q)$. Furthermore if $Q > q$, $t|(Q-1)$, and $t|(q-1)$, then $D_Q = (Q-1)/t$ and $\varepsilon_Q = 1/t$. Similarly, $D_q = (q-1)/t$ and $\varepsilon_q = 1/t$. Then, $|\varepsilon_q - q\varepsilon_Q/Q| = \frac{1}{t}(1 - \frac{q}{Q}) < \frac{1}{t}$. So,

$$\begin{aligned} ||e_0'||_\infty &= \left\| \frac{qe_0}{Q} + (\varepsilon_q - q\varepsilon_Q/Q) m_0 - \varepsilon_2 - \varepsilon_1 s \right\|_\infty \\ &\le \frac{q}{Q} ||e_0||_\infty + |\varepsilon_q - q\varepsilon_Q/Q| \frac{t}{2} + ||\varepsilon_2||_\infty + ||\varepsilon_1 s||_\infty \\ &\le \frac{q}{Q} E + 1 + \frac{\delta_R ||s||_\infty}{2}. \end{aligned}$$

By assumption $Q/q > \frac{2E}{\delta_R ||s||_\infty - 2}$, so $||e_0'||_\infty < \delta_R ||s||_\infty$. □

Proof of Lemma 2.3. First, note that a_0' and b_0' are polynomials with integer coefficients since both $qa_0 + t\omega_a$ and $qb_0 + t\omega_b$ are equivalent to 0 modulo Q. Then, we have

$$b_0' + a_0' s \equiv \frac{qb_0 + t\omega_b}{Q} + \frac{qa_0 + t\omega_a}{Q} s \bmod \phi(x) \tag{A1}$$

$$\equiv \frac{q}{Q}(b_0 + a_0 s) + \frac{t}{Q}(\omega_b + \omega_a s) \bmod \phi(x) \tag{A2}$$

$$\equiv \frac{q}{Q}(m_0 + te_0) + \frac{t}{Q}(\omega_b + \omega_a s) + qr \bmod \phi(x) \tag{A3}$$

$$\equiv m_0 + t\left(\frac{q-Q}{Qt}m_0 + \frac{q}{Q}e_0 + \frac{1}{Q}(\omega_b + \omega_a s)\right) + qr \bmod \phi(x) \tag{A4}$$

$$\equiv m_0 + te_0' \bmod (\phi(x), q), \tag{A5}$$

where $e_0' = \frac{q-Q}{Qt}m_0 + \frac{q}{Q}e_0 + \frac{1}{Q}(\omega_b + \omega_a s)$. Since $Q > |q - Q|$, we have $|\frac{q-Q}{Qt}| < \frac{1}{t}$. Thus, we have that

$$\begin{aligned}
||e_0'||_\infty &= \left\| \frac{q-Q}{Qt}m_0 + \frac{q}{Q}e_0 + \frac{1}{Q}(\omega_b + \omega_a s) \right\|_\infty \\
&\le \left| \frac{q-Q}{Qt} \right| ||m_0||_\infty + \frac{q}{Q}||e_0||_\infty + \frac{1}{Q}||\omega_b||_\infty + \frac{1}{Q}||\omega_a s||_\infty \\
&\le \frac{1}{2} + \frac{q}{Q}E + \frac{1}{2} + \frac{\delta_R||s||_\infty}{2} \\
&= \frac{q}{Q}E + 1 + \frac{\delta_R||s||_\infty}{2}.
\end{aligned}$$

By assumption $Q/q > \frac{2E}{\delta_R||s||_\infty - 2}$, so $||e_0'||_\infty < \delta_R||s||_\infty$. □

Proof of Lemma 3.1. By assumption, the public key $\mathrm{pk} = (\mathrm{k}_0, \mathrm{k}_1)$ satisfies

$$\mathrm{k}_1 + \mathrm{k}_0 s \equiv e \bmod (\phi(x), p_0 q)$$

for some noise $e \in R_n$ with $||e||_\infty \le \rho$. Then,

$$\begin{aligned}
b_0 + a_0 s &\equiv \mathrm{k}_1 u + e_2 + (\mathrm{k}_0 u + e_1)s \bmod (\phi(x), p_0 q) \\
&\equiv -(\mathrm{k}_0 s + e)u + e_2 + (\mathrm{k}_0 u + e_1)s \bmod (\phi(x), p_0 q) \\
&\equiv -\mathrm{k}_0 su - eu + e_2 + \mathrm{k}_0 su + e_1 s \bmod (\phi(x), p_0 q) \\
&\equiv -eu + e_2 + e_1 s \bmod (\phi(x), p_0 q).
\end{aligned}$$

Let $e_0 = -eu + e_2 + e_1 s$. Then, $b_0 + a_0 s \equiv e_0 \bmod (\phi(x), p_0 q)$. Note that

$$\begin{aligned}
||e_0||_\infty &= || - eu + e_2 + e_1 s||_\infty \\
&\le \delta_R||e||_\infty||u||_\infty + ||e_2||_\infty + \delta_R||e_1||_\infty||s||_\infty \\
&\le \delta_R^2||u||_\infty + \delta_R + \delta_R^2||s||_\infty \\
&= \delta_R^2(||u||_\infty + ||s||_\infty) + \delta_R.
\end{aligned}$$

Since $||s||_\infty = ||u||_\infty = 1$, $||e_0||_\infty \le 2\delta_R^2 + \delta_R$. By assumption $p_0 > \frac{4\delta_R^2 + 2\delta_R}{\delta_R - 1}$, so BFV.Modreduce$(\mathrm{ct}_0, p_0 q, q)$ outputs $\mathrm{ct}_0' = (a_0', b_0^*)$ satisfying

$$b_0^* \equiv -a_0' s + e_0' \bmod (\phi(x), q)$$

with $||e_0'||_\infty \le \rho$ by Lemma 2.1. Since $b_0' = [b_0^* + D_q m_0]_q$, we have

$$b_0' + a_0' s \equiv D_q m_0 + e_0' \bmod (\phi(x), q).$$

□

Proof of Lemma 3.2. By assumption, $b_0 + a_0 s \equiv D_q m_0 + e_0 \bmod (\phi(x), q)$ with $||e_0||_\infty < (D_q - 1)/2$. There is a polynomial $r \in R_n$ such that $[b_0 + a_0 s]_{\phi(x),q} = D_q m_0 + e_0 + qr \bmod \phi(x)$. Thus,

$$tc = t[b_0 + a_0 s]_{\phi(x),q} = tD_q m_0 + te_0 + tqr = qm_0 - m_0 + te_0 + tqr.$$

Then,

$$\left[\left\lfloor \frac{tc}{q} \right\rceil\right]_t = \left[m_0 + \left\lfloor -\frac{m_0}{q} + \frac{t}{q}e_0 \right\rceil + tr\right]_t = \left[m_0 + \left\lfloor \frac{t}{q}(e_0 - \frac{1}{t}m_0) \right\rceil\right]_t = m_0.$$

The last equality follows from the fact that $\left\lfloor \frac{t}{q}(e_0 - \frac{1}{t}m_0) \right\rceil = 0$, as $||e_0||_\infty < (D_q - 1)/2$. □

Proof of Lemma 3.3. By assumption, each $\mathrm{ct}_i = (a_i, b_i) \in R_{n,q}^2$ satisfies

$$b_i + a_i s \equiv D_q m_i + e_i \bmod (\phi(x), q)$$

with $||e_i||_\infty \le E$ for all $i = 0, \ldots, k-1$. Then,

$$\mathrm{ct}_0' = \left[\sum_{i=0}^{k-1} \alpha_i \mathrm{ct}_i\right]_q = \left(\left[\sum_{i=0}^{k-1} \alpha_i a_i\right]_q, \left[\sum_{i=0}^{k-1} \alpha_i b_i\right]_q\right).$$

Since $t|(q-1)$ and $D_q = (q-1)/t$, we have $D_q t \equiv -1 \bmod q$. There exists $r \in R_n$ such that

$$\sum_{i=0}^{k-1} \alpha_i m_i = \left[\sum_{i=0}^{k-1} \alpha_i m_i\right]_t + tr$$

with $||r||_\infty \le M$. Then,

$$\begin{aligned}
\sum_{i=0}^{k-1} \alpha_i b_i + \sum_{i=0}^{k-1} \alpha_i a_i s &\equiv D_q\left(\sum_{i=0}^{k-1} \alpha_i m_i\right) + \sum_{i=0}^{k-1} \alpha_i e_i \bmod (\phi(x), q) \\
&\equiv D_q\left[\sum_{i=0}^{k-1} \alpha_i m_i\right]_t + \sum_{i=0}^{k-1} \alpha_i e_i + D_q t r \bmod (\phi(x), q) \\
&\equiv D_q\left[\sum_{i=0}^{k-1} \alpha_i m_i\right]_t + \sum_{i=0}^{k-1} \alpha_i e_i - r \bmod (\phi(x), q).
\end{aligned}$$

Let $e' = \sum_{i=0}^{k-1} \alpha_i e_i - r$. Then, we have

$$\sum_{i=0}^{k-1} \alpha_i b_i + \sum_{i=0}^{k-1} \alpha_i a_i s \equiv D_q\left[\sum_{i=0}^{k-1} \alpha_i m_i\right]_t + e' \bmod (\phi(x), q).$$

So, ct_0' is a BFV ciphertext with noise term e' and

$$||e'||_\infty \le \sum_{i=0}^{k-1} |\alpha_i| ||e_i||_\infty + ||r||_\infty \le ME + M = M(E+1).$$ □

Proof of Lemma 3.4. By assumption, we have for $i = 0,1$,

$$b_i + a_i s \equiv D_q m_i + e_i \bmod (\phi(x), q), \tag{A6}$$

with $||a_i||_\infty \le q/2$, $||b_i||_\infty \le q/2$, and $||e_i||_\infty \le E$. We can rewrite (A6) as follows:

$$b_i + a_i s \equiv D_q m_i + e_i + q r_i \bmod \phi(x), \tag{A7}$$

where

$$||r_i||_\infty \le \frac{1}{q}||b_i + a_i s||_\infty \le \frac{1}{q}\left(\frac{q}{2} + \frac{q}{2}\delta_R ||s||_\infty\right) \le \frac{1}{2}(1 + \delta_R ||s||_\infty) \le \delta_R ||s||_\infty.$$

On the one hand,

$$(t/q)(b_0 + a_0 s)(b_1 + a_1 s) \equiv (t/q)(b_0 b_1 + (b_1 a_0 + b_0 a_1)s + a_0 a_1 s^2) \bmod \phi(x)$$

$$\begin{aligned} &\equiv (t/q)(c_0 + c_1 s + c_2 s^2) \bmod \phi(x) \\ &\equiv c_0' + c_1' s + c_2' s^2 + (\varepsilon_0' + \varepsilon_1' s + \varepsilon_2' s^2) \bmod \phi(x), \end{aligned}$$

with $\|\varepsilon_i'\|_\infty \le 1/2$ for $0 \le i \le 2$. Let $r_m \in R_n$ so that $m_0 m_1 = [m_0 m_1]_{\phi(x),t} + t r_m$ with $\|r_m\|_\infty \le t\delta_R/4$. Then on the other hand, noting that $(t/q)D_q = (t/q)(q/t - 1/t) = 1 - 1/q$ and $tD_q \equiv -1 \bmod q$, we have

$$\begin{aligned} &(t/q)(D_q m_0 + e_0 + q r_0)(D_q m_1 + e_1 + q r_1) \\ &\quad \equiv (t/q)(D_q^2 m_0 m_1 + D_q(m_0 e_1 + m_1 e_0) + q(e_0 r_1 + r_0 e_1) + e_0 e_1 + q D_q(m_0 r_1 + r_0 m_1) + q^2 r_0 r_1) \bmod \phi(x) \\ &\quad \equiv (1 - 1/q) D_q [m_0 m_1]_{\phi(x),t} + (1 - 1/q) D_q t r_m + (1 - 1/q)(m_0 e_1 + m_1 e_0) + t(e_0 r_1 + r_0 e_1) + (t/q) e_0 e_1 \\ &\qquad + t D_q(m_0 r_1 + r_0 m_1) + t q r_0 r_1 \bmod \phi(x) \\ &\quad \equiv D_q [m_0 m_1]_{\phi(x),t} - (D_q/q)[m_0 m_1]_{\phi(x),t} + D_q t r_m - (D_q t/q) r_m + (m_0 e_1 + m_1 e_0) - (1/q)(m_0 e_1 + m_1 e_0) \\ &\qquad + t(e_0 r_1 + r_0 e_1) + (t/q) e_0 e_1 + t D_q(m_0 r_1 + r_0 m_1) + t q r_0 r_1 \bmod \phi(x) \\ &\quad \equiv D_q [m_0 m_1]_{\phi(x),t} - r_m + (m_0 e_1 + m_1 e_0) + t(e_0 r_1 + r_0 e_1) - (m_0 r_1 + r_0 m_1) - (D_q/q)[m_0 m_1]_{\phi(x),t} \\ &\qquad - (1/q)(m_0 e_1 + m_1 e_0) + (t/q) e_0 e_1 \bmod (\phi(x), q). \end{aligned}$$

Let

$$\varepsilon_1 = \varepsilon_0' + \varepsilon_1' s + \varepsilon_2' s^2 \quad \text{and}$$
$$\varepsilon_2 = -r_m + (m_0 e_1 + m_1 e_0) + t(e_0 r_1 + r_0 e_1) - (m_0 r_1 + r_0 m_1) - (D_q/q)[m_0 m_1]_{\phi(x),t} - (1/q)(m_0 e_1 + m_1 e_0) + (t/q) e_0 e_1.$$

Let $e' = \varepsilon_2 - \varepsilon_1$. Then,

$$c_0' + c_1' s + c_2' s^2 \equiv D_q [m_0 m_1]_{\phi(x),t} + e' \bmod (\phi(x), q).$$

We now turn to the noise bound for e'. First, note that

$$\|\varepsilon_1\|_\infty = \|\varepsilon_0' + \varepsilon_1' s + \varepsilon_2' s^2\|_\infty \le \frac{1}{2} + \frac{\delta_R \|s\|_\infty}{2} + \frac{\delta_R^2 \|s\|_\infty^2}{2} = \frac{(1 + \delta_R \|s\|_\infty)^2}{2}.$$

To simplify the noise analysis for ε_2, we break down ε_2 into two pieces. Let

$$\omega_1 = -r_m + (m_0 e_1 + m_1 e_0) + t(e_0 r_1 + r_0 e_1) - (m_0 r_1 + r_0 m_1) \quad \text{and}$$
$$\omega_2 = -(D_q/q)[m_0 m_1]_{\phi(x),t} - (D_q t/q) r_m - (1/q)(m_0 e_1 + m_1 e_0) + (t/q) e_0 e_1.$$

Then, $\varepsilon_2 = \omega_1 + \omega_2$. For ω_1, we have

$$\begin{aligned} \|\omega_1\|_\infty &\le \|r_m\|_\infty + \|m_0 e_1 + m_1 e_0\|_\infty + t\|e_0 r_1 + r_0 e_1\|_\infty + \|m_0 r_1 + r_0 m_1\|_\infty \\ &\le \frac{t\delta_R}{4} + Et\delta_R + 2Et\delta_R^2 \|s\|_\infty + t\delta_R^2 \|s\|_\infty. \end{aligned}$$

For ω_2, note that since $D_q < q/t$ and $q > 2t$, we have

$$\begin{aligned} \|\omega_2\|_\infty &\le \frac{D_q}{q}\|[m_0 m_1]_{\phi(x),t}\|_\infty + \frac{D_q t}{q}\|r_m\|_\infty + \frac{1}{q}\|m_0 e_1 + m_1 e_0\|_\infty + \frac{t}{q}\|e_0 e_1\|_\infty \\ &< \frac{1}{t}\|[m_0 m_1]_{\phi(x),t}\|_\infty + \|r_m\|_\infty + \frac{1}{q}\|m_0 e_1 + m_1 e_0\|_\infty + \frac{t}{q}\|e_0 e_1\|_\infty \\ &\le \frac{1}{2} + \frac{t\delta_R}{4} + \frac{Et\delta_R}{q} + \frac{E\delta_R}{2} \\ &\le \frac{1}{2} + \frac{t\delta_R}{4} + E\delta_R. \end{aligned}$$

The bound on $\frac{t}{q}\|e_0 e_1\|_\infty$ follows from the fact that since e_0, e_1 are ciphertext noise terms, so by assumption

$\|e_i\|_\infty \le E < (D_q - 1)/2 < D_q/2 < q/(2t)$ for $i = 0, 1$. So,

$$\frac{t}{q}\|e_0 e_1\|_\infty \le \frac{t}{q}E^2\delta_R < \frac{E\delta_R}{2}.$$

By assumption, $\delta_R \ge 16$, $E \ge 1$, $t \ge 2$, and $\|s\|_\infty = 1$. So,

$$\begin{aligned}
\|e'\|_\infty &\le \frac{t\delta_R}{4} + Et\delta_R + 2Et\delta_R^2\|s\|_\infty + t\delta_R^2\|s\|_\infty + \frac{1}{2} + \frac{t\delta_R}{4} + E\delta_R + \frac{(\delta_R\|s\|_\infty + 1)^2}{2} \\
&\le \frac{t\delta_R}{2} + Et\delta_R + 3Et\delta_R^2\|s\|_\infty + 1 + E\delta_R + \frac{\delta_R^2\|s\|_\infty^2}{2} + \delta_R\|s\|_\infty \\
&= Et\delta_R^2\|s\|_\infty^2\left(\frac{1}{E\delta_R\|s\|_\infty^2} + \frac{1}{\delta_R^2\|s\|_\infty^2} + \frac{3}{\|s\|_\infty} + \frac{1}{Et\delta_R^2\|s\|_\infty^2} + \frac{1}{t\delta_R\|s\|_\infty^2} + \frac{1}{2Et} + \frac{1}{Et\delta_R\|s\|_\infty}\right) \\
&\le Et\delta_R^2\|s\|_\infty^2\left(\frac{1}{16} + \frac{1}{16^2} + 3 + \frac{1}{2\cdot 16^2} + \frac{1}{2\cdot 16} + \frac{1}{2\cdot 2} + \frac{1}{2\cdot 16}\right) \\
&< 3.5Et\delta_R^2\|s\|_\infty^2 \\
&= 3.5tE\rho^2.
\end{aligned}$$

□

Proof of Lemma 3.5. Notice that as $\beta_1 \equiv c_2'\mathrm{k}_1' \bmod (\phi(x), p_1 q)$, we can write

$$\beta_1 \equiv c_2'\mathrm{k}_1' + p_1 q\alpha_1 \bmod \phi(x)$$

for some $\alpha_1 \in \mathbb{Z}$. Then,

$$d_1' = \left\lfloor \frac{\beta_1}{p_1} \right\rfloor \equiv \left\lfloor \frac{c_2'\mathrm{k}_1' + p_1 q\alpha_1}{p_1} \right\rfloor \equiv \frac{c_2'\mathrm{k}_1'}{p_1} + q\alpha_1 + \varepsilon_1 \equiv \frac{c_2'\mathrm{k}_1'}{p_1} + \varepsilon_1 \bmod (\phi(x), q)$$

for some $\varepsilon_1 \in \mathbb{R}[x]/(\phi(x))$ with $\|\varepsilon_1\|_\infty \le 1/2$. Note also that as $c_2'\mathrm{k}_0' \equiv -c_2'\mathrm{k}_1' s + c_2' e_1' + c_2' p_1 s^2 \bmod (\phi(x), p_1 q)$, we have that for some $\alpha_0 \in \mathbb{Z}$,

$$\beta_0 \equiv c_2'\mathrm{k}_0' \equiv -c_2'\mathrm{k}_1' s + c_2' e_1' + c_2' p_1 s^2 + p_1 q\alpha_0 \bmod \phi(x).$$

Then,

$$\begin{aligned}
d_0' &= \left\lfloor \frac{\beta_0}{p_1} \right\rfloor \\
&\equiv \left\lfloor \frac{-c_2'\mathrm{k}_1' s + c_2' e_1' + c_2' p_1 s^2 + p_1 q\alpha_0}{p_1} \right\rfloor \bmod (\phi(x), q) \\
&\equiv \left\lfloor \frac{-c_2'\mathrm{k}_1' s}{p_1} + \frac{c_2' e_1'}{p_1} + c_2' s^2 + q\alpha_0 \right\rfloor \bmod (\phi(x), q) \\
&\equiv \frac{-c_2'\mathrm{k}_1' s}{p_1} + \frac{c_2' e_1'}{p_1} + c_2' s^2 + q\alpha_0 + \varepsilon_0 \bmod (\phi(x), q) \\
&\equiv \frac{-c_2'\mathrm{k}_1' s}{p_1} + \frac{c_2' e_1'}{p_1} + c_2' s^2 + \varepsilon_0 \bmod (\phi(x), q)
\end{aligned}$$

for some $\varepsilon_0 \in \mathbb{R}[x]/(\phi(x))$ with $\|\varepsilon_0\|_\infty \le 1/2$. Therefore,

$$\begin{aligned}
d_0' + d_1' s &\equiv c_2' s^2 - \frac{c_2'\mathrm{k}_1' s}{p_1} + \frac{c_2' e_1'}{p_1} + \varepsilon_0 + \frac{c_2'\mathrm{k}_1' s}{p_1} + \varepsilon_1 s \bmod (\phi(x), q) \\
&\equiv c_2' s^2 + \frac{c_2' e_1'}{p_1} + \varepsilon_0 + \varepsilon_1 s \bmod (\phi(x), q)
\end{aligned}$$

Let $e'' = \frac{c_2' e_1'}{p_1} + \varepsilon_0 + \varepsilon_1 s$. Then, $d_0' + d_1' s \equiv c_2' s^2 + e'' \bmod (\phi(x), q)$. By assumption, (c_0', c_1', c_2') satisfies

$$c_0' + c_1' s + c_2' s^2 \equiv D_q[m_0 m_1]_{\phi(x),t} + e' \bmod (\phi(x), q).$$

Let $e^* = e' + e''$. Then,

$$\begin{aligned} c_1 + c_0 s &\equiv c_0' + d_0' + c_1' s + d_1' s \bmod (\phi(x), q) \\ &\equiv c_0' + c_1' s + c_2' s^2 + e'' \bmod (\phi(x), q) \\ &\equiv D_q[m_0 m_1]_{\phi(x),t} + e' + e'' \bmod (\phi(x), q) \\ &\equiv D_q[m_0 m_1]_{\phi(x),t} + e^* \bmod (\phi(x), q). \end{aligned}$$

Now, we turn to the noise bounds on e'' and e^*. Note that as $e_1' \leftarrow \chi_\rho$, we have $\|e_1'\|_\infty \le \rho = \delta_R \|s\|_\infty$. If $p_1 \ge 6q$ and $\delta_R \ge 16$, then

$$\begin{aligned} \|e''\|_\infty &= \left\| \frac{c_2' e_1'}{p_1} + \varepsilon_0 + \varepsilon_1 s \right\|_\infty \\ &\le \frac{q}{2p_1} \delta_R^2 \|s\|_\infty + \frac{1}{2} + \frac{\delta_R \|s\|_\infty}{2} \\ &\le \delta_R^2 \|s\|_\infty \left(\frac{1}{12} + \frac{1}{\delta_R^2 \|s\|_\infty} + \frac{1}{2\delta_R} \right) \\ &< \frac{1}{8} \delta_R^2 \|s\|_\infty. \end{aligned}$$

By Lemma 7, we have that $\|e'\|_\infty \le 3.5Et\rho^2 = 3.5Et\delta_R^2 \|s\|_\infty^2$. As relinearization introduces additional noise e'' bounded by $\frac{1}{8}\delta_R^2 \|s\|_\infty$, we have

$$\begin{aligned} \|e^*\|_\infty &\le 3.5Et\delta_R^2 \|s\|_\infty^2 + \frac{1}{8}\delta_R^2 \|s\|_\infty \\ &= Et\delta_R^2 \|s\|_\infty^2 \left(3.5 + \frac{1}{8Et\|s\|_\infty} \right) \\ &\le Et\delta_R^2 \|s\|_\infty^2 \left(3.5 + \frac{1}{16} \right) \\ &< 3.6Et\delta_R^2 \|s\|_\infty^2 \\ &= 3.6Et\rho^2. \end{aligned}$$

□

Proof of Lemma 3.6. By assumption, the public key $\mathrm{pk} = (\mathrm{k}_0, \mathrm{k}_1)$ satisfies

$$\mathrm{k}_1 + \mathrm{k}_0 s \equiv te \bmod (\phi(x), p_0 q)$$

for some noise $e \in R_n$ with $\|e\|_\infty \le \rho$. Then,

$$\begin{aligned} b_0 + a_0 s &\equiv \mathrm{k}_1 u + te_2 + (\mathrm{k}_0 u + te_1) s \bmod (\phi(x), p_0 q) \\ &\equiv -(\mathrm{k}_0 s + te) u + te_2 + (\mathrm{k}_0 u + te_1) s \bmod (\phi(x), p_0 q) \\ &\equiv -\mathrm{k}_0 su - teu + te_2 + \mathrm{k}_0 su + te_1 s \bmod (\phi(x), p_0 q) \\ &\equiv -t(eu + e_2 + e_1 s) \bmod (\phi(x), p_0 q) \\ &\equiv te_0 \bmod (\phi(x), p_0 q). \end{aligned}$$

Let $e_0 = -eu + e_2 + e_1 s$. Then, $b_0 + a_0 s \equiv te_0 \bmod (\phi(x), p_0 q)$. Since $\|s\|_\infty = \|u\|_\infty = 1$, we have $\|e_0\|_\infty \le 2\delta_R^2 + \delta_R$. By assumption $p_0 > \frac{4\delta_R^2 + 2\delta_R}{\delta_R - 2}$, so BGV.Modreduce$(\mathrm{ct}_0, p_0 q, q)$ outputs $\mathrm{ct}_0' = (a_0', b_0^*)$ satisfying

$$b_0^* \equiv -a_0' s + te_0' \bmod (\phi(x), q)$$

with $\|e_0'\|_\infty \le \rho$ by Lemma 2.3. Since $b_0' = [b_0^* + m_0]_q$, we have

$$b_0' + a_0' s \equiv m_0 + te_0' \bmod (\phi(x), q).$$

□

Proof of Lemma 3.7. By assumption, we have for $i = 0,1$,

$$b_i + a_i s \equiv m_i + te_i \bmod (\phi(x), q)$$

with $\|e_i\|_\infty \le E$. Let $r_m \in R_n$ such that $m_0 m_1 = [m_0 m_1]_{\phi(x),t} + tr_m$. Note that $\|r_m\|_\infty \le t$. We have that

$$\begin{aligned} c_0' + c_1' s + c_2' s^2 &\equiv b_0 b_1 + (b_1 a_0 + b_0 a_1)s + a_0 a_1 s^2 \bmod (\phi(x), q) \\ &\equiv (b_0 + a_0 s)(b_1 + a_1 s) \bmod (\phi(x), q) \\ &\equiv (m_0 + te_0)(m_1 + te_1) \bmod (\phi(x), q) \\ &\equiv m_0 m_1 + t(m_0 e_1 + m_1 e_0 + te_0 e_1) \bmod (\phi(x), q) \\ &\equiv [m_0 m_1]_{\phi(x),t} + t(m_0 e_1 + m_1 e_0 + te_0 e_1 + r_m) \bmod (\phi(x), q). \end{aligned}$$

Let $e' = m_0 e_1 + m_1 e_0 + te_0 e_1 + r_m$. Then,

$$c_0' + c_1' s + c_2' s^2 \equiv [m_0 m_1]_{\phi(x),t} + te' \bmod (\phi(x), q)$$

with

$$\|e'\|_\infty \le \delta_R t E + \delta_R t E^2 + t \le 2\delta_R t E^2 + t \le 2\delta_R t(E^2 + 1). \qquad \square$$

Proof of Lemma 3.8. First, observe that both $\beta_0 + t\omega_0$ and $\beta_1 + t\omega_1$ are divisible by p_1 since for $i = 0, 1$,

$$\beta_i + t\omega_i = \beta_i + t[-t^{-1}\beta_i]_{p_1} \equiv \beta_i + t(-t^{-1}\beta_i) \equiv 0 \bmod p_1,$$

so $d_0' = \frac{\beta_0 + t\omega_0}{p_1}$ and $d_1' = \frac{\beta_1 + t\omega_1}{p_1}$ have integer coefficients. Then,

$$\begin{aligned} d_0' + d_1' s &\equiv \frac{\beta_0 + t\omega_0}{p_1} + \frac{\beta_1 + t\omega_1}{p_1} s \bmod (\phi(x), q) \\ &\equiv \frac{c_2' \mathrm{k}_0' + \omega_0}{p_1} + \frac{c_2' \mathrm{k}_1' + \omega_1}{p_1} s \bmod (\phi(x), q) \\ &\equiv \frac{c_2'(-\mathrm{k}_1' s + p_1 s^2 + te_1') + t\omega_0}{p_1} + \frac{c_2' \mathrm{k}_1' + t\omega_1}{p_1} s \bmod (\phi(x), q) \\ &\equiv c_2' s^2 + t\frac{c_2' e_1' + \omega_0 + \omega_1 s}{p_1} \bmod (\phi(x), q). \end{aligned}$$

Let $e'' = \frac{c_2' e_1'}{p_1} + \frac{\omega_0 + \omega_1 s}{p_1}$, which has integer coefficients. Then, $d_0' + d_1' s \equiv c_2' s^2 + te''$. By assumption, (c_0', c_1', c_2') satisfies

$$c_0' + c_1' s + c_2' s^2 \equiv [m_0 m_1]_{\phi(x),t} + te' \bmod (\phi(x), q),$$

where $\|e'\|_\infty \le 2\delta_R t(E^2 + 1)$. Let $e^* = e' + e''$. Then,

$$\begin{aligned} c_1 + c_0 s &\equiv c_0' + d_0' + c_1' s + d_1' s \bmod (\phi(x), q) \\ &\equiv c_0' + c_1' s + c_2' s^2 + te'' \bmod (\phi(x), q) \\ &\equiv [m_0 m_1]_{\phi(x),t} + t(e' + e'') \bmod (\phi(x), q) \\ &\equiv [m_0 m_1]_{\phi(x),t} + te^* \bmod (\phi(x), q). \end{aligned}$$

Then, note that

$$\begin{aligned} \|e''\|_\infty &\le \left\| \frac{c_2' e_1'}{p_1} \right\|_\infty + \left\| \frac{\omega_0 + \omega_1 s}{p_1} \right\|_\infty \\ &\le \frac{q}{2p_1}\delta_R^2 \|s\|_\infty + \frac{p_1}{2p_1} + \frac{p_1}{2p_1}\delta_R \|s\|_\infty \\ &\le \frac{1}{12}\delta_R^2 \|s\|_\infty + \frac{1}{2} + \frac{\delta_R \|s\|_\infty}{2} \\ &= \delta_R^2 \|s\|_\infty \left(\frac{1}{12} + \frac{1}{2\delta_R^2 \|s\|_\infty} + \frac{1}{2\delta_R} \right) \end{aligned}$$

$$\leq \delta_R^2 ||s||_\infty \left(\frac{1}{12} + \frac{1}{512} + \frac{1}{32}\right)$$
$$< \frac{1}{8}\delta_R^2 ||s||_\infty.$$

The bound on e^* then follows immediately. □

Proof of Lemma 3.9. By assumption, we first note that $(a_0, b_0) \in R_{n,Q}^2$ satisfies

$$b_0 \equiv -a_0 s + m_0 + e_0 \bmod (\phi(x), Q).$$

Therefore, there is some integer $r \in \mathbb{Z}$ such that $b_0 + a_0 s \equiv m_0 + e_0 + Qr \bmod \phi(x)$. Let $\varepsilon_1 = qa_0/Q - a_0'$, and $\varepsilon_2 = qb_0/Q - b_0'$. Then,

$$\begin{aligned} b_0' &= \frac{qb_0}{Q} - \varepsilon_2 \\ &\equiv -\frac{qa_0 s}{Q} + \frac{q}{Q}m_0 + \frac{qe_0}{Q} - \varepsilon_2 + qr \bmod \phi(x) \\ &\equiv -a_0' s - \varepsilon_1 s + \frac{q}{Q}m_0 + \frac{qe_0}{Q} - \varepsilon_2 + qr \bmod \phi(x). \end{aligned}$$

Let $e_0' = \frac{q}{Q}e_0 - \varepsilon_2 - \varepsilon_1 s$. Then, $b_0' + a_0' s \equiv \frac{q}{Q}m_0 + e_0' \bmod (\phi(x), q)$. Therefore,

$$||e_0'||_\infty \leq \frac{q}{Q}E + \frac{1 + \delta_R ||s||_\infty}{2}.$$

By assumption $Q/q > \frac{2E}{\delta_R ||s||_\infty - 1}$, so $||e_0'||_\infty < \delta_R ||s||_\infty$. □

Proof of Lemma 3.10. We consider step 2 of Algorithm 14, in which we compute $\mathrm{ct} \coloneqq \mathrm{BFV.Encrypt}(m, 1, \mathrm{pk})$. By an argument identical to the proof of Lemma 3.1, steps 1–3 of Algorithm 4 produce an ordered pair $(a_0, b_0) \in R_{n,p_0 q}^2$ satisfying

$$b_0 + a_0 s \equiv e_0 \bmod (\phi(x), p_0 q)$$

for $e_0 \in R_n$ with $||e_0||_\infty \leq 2\delta_R^2 + \delta_R$. By assumption $p_0 > \frac{4\delta_R^2 + 2\delta_R}{\delta_R - 1}$, so $(a_0', b_0^*) \coloneqq \mathrm{CKKS.Modreduce}(p_0 q, q, \mathrm{ct}_0)$ satisfies

$$b_0^* \equiv -a_0' s + e_0' \bmod (\phi(x), q)$$

with $||e_0'||_\infty \leq \rho$ by Lemma 3.9. Since $b_0' = [b_0^* + m]_q$, we have

$$b_0' + a_0' s \equiv m + e_0' \bmod (\phi(x), q).$$

Let $(a, b) = (a_0', b_0')$. Then, $\mathrm{ct} = (a, b)$ is the output of Algorithm 14 and is a CKKS ciphertext with noise bounded by ρ. □

Proof of Lemma 3.11. By assumption, each $\mathrm{ct}_i = (a_i, b_i) \in R_{n,q}^2$ satisfies

$$b_i + a_i s \equiv m_i + e_i \bmod (\phi(x), q)$$

for some noise term e_i with $||e_i||_\infty \leq E$. Then,

$$\mathrm{ct}_0' = \left[\sum_{i=0}^{k-1} a_i \mathrm{ct}_i\right]_q = \left(\left[\sum_{i=0}^{k-1} a_i a_i\right]_q, \left[\sum_{i=0}^{k-1} a_i b_i\right]_q\right).$$

Let $e' = \sum_{i=0}^{k-1} a_i e_i$. Then,

$$\sum_{i=0}^{k-1} a_i b_i + \sum_{i=0}^{k-1} a_i a_i s \equiv \sum_{i=0}^{k-1} a_i m_i + \sum_{i=0}^{k-1} a_i e_i \equiv \sum_{i=0}^{k-1} a_i m_i + e' \bmod (\phi(x), q).$$

So, ct_0' is a CKKS ciphertext with noise term e' and

$$||e'||_\infty \leq \sum_{i=0}^{k-1} |a_i| ||e_i||_\infty \leq ME.$$

□

Proof of Lemma 3.12. By assumption, we have for $i = 0, 1$,

$$b_i + a_i s \equiv m_i + e_i \bmod (\phi(x), q)$$

with $\|e_i\|_\infty \le E$. Let $e' = m_0 e_1 + m_1 e_0 + e_0 e_1$. Then,

$$\begin{aligned} c_0' + c_1' s + c_2' s^2 &= (b_0 + a_0 s)(b_1 + a_1 s) \\ &\equiv (m_0 + e_0)(m_1 + e_1) \bmod (\phi(x), q) \\ &\equiv m_0 m_1 + m_0 e_1 + m_1 e_0 + e_0 e_1 \bmod (\phi(x), q) \\ &\equiv m_0 m_1 + e' \bmod (\phi(x), q). \end{aligned}$$

The bound on e' then follows immediately. □

Proof of Lemma 3.13. It is clear from the proof of Lemma 3.5 that $d_0' + d_1' s \equiv c_2' s^2 + e'' \bmod (\phi(x), q)$ where $e'' = \frac{c_2' e_1'}{p_1} + \varepsilon_0 + \varepsilon_1 s$ and $\varepsilon_0, \varepsilon_1 \in \mathbb{R}[x]/(\phi(x))$ with $\|\varepsilon_0\|_\infty \le 1/2$ and $\|\varepsilon_1\|_\infty \le 1/2$. By assumption, $c_0' + c_1' s + c_2' s^2 \equiv m_0 m_1 + e' \bmod (\phi(x), q)$ with $\|e'\|_\infty \le Et\delta_R + E^2\delta_R$, so

$$\begin{aligned} c_1 + c_0 s &\equiv c_0' + d_0' + c_1' s + d_1' s \bmod (\phi(x), q) \\ &\equiv c_0' + c_1' s + c_2' s^2 + e'' \bmod (\phi(x), q) \\ &\equiv m_0 m_1 + e' + e'' \bmod (\phi(x), q) \end{aligned}$$

Let $e^* = e' + e''$. Then, $c_1 + c_0 s \equiv m_0 m_1 + e^* \bmod (\phi(x), q)$. The bound on e^* follows immediately. □

Proof of Lemma 4.1. Suppose we have a collection of BFV ciphertexts, each with noise bounded by $\rho = \delta_R \|s\|_\infty$. By Lemma 3.3, computing

$$\mathrm{ct}_j \coloneqq \mathrm{Linearcombo}(\mathrm{ct}_{j,1}, \ldots, \mathrm{ct}_{j,k_1}, 1, \ldots, 1)$$

results in BFV ciphertexts ct_j for $j = 1, \ldots, 2k_2$, each with noise bounded by $k_1\delta_R\|s\|_\infty + k_1$. Since $\delta_R \ge 16$, we have

$$k_1\delta_R\|s\|_\infty + k_1 = \left(k_1 + \frac{k_1}{\delta_R\|s\|_\infty}\right)\delta_R\|s\|_\infty \le \frac{17}{16}k_1\delta_R\|s\|_\infty.$$

By Lemma 3.4, each Multiply(ct_{2j-1}, ct_{2j}) in step 2 of Algorithm 16 for $j = 1, \ldots, k_2$ results in a polynomial triple with noise bounded by

$$3.5\left(\frac{17}{16}k_1\delta_R\|s\|_\infty\right)t\delta_R^2\|s\|_\infty^2 = \frac{119}{32}k_1 t\delta_R^3\|s\|_\infty^3.$$

Summing all k_2 of these polynomial triples in step 2 of Algorithm 16 results in a polynomial triple with noise bounded by

$$\frac{119}{32}k_1k_2t\delta_R^3\|s\|_\infty^3 + k_2$$

by an equivalent argument to Lemma 3.3 with polynomial triples as the input. Notice,

$$\begin{aligned} \frac{119}{32}k_1k_2t\delta_R^3\|s\|_\infty^3 + k_2 &= \frac{119}{32}k_1k_2t\delta_R^3\|s\|_\infty^3\left(1 + \frac{32}{119k_1t\delta_R^3\|s\|_\infty^3}\right) \\ &\le \frac{119}{32}k_1k_2t\delta_R^3\|s\|_\infty^3\left(1 + \frac{32}{119 \cdot 2 \cdot 16^3}\right) \\ &\le \frac{30}{8}k_1k_2t\delta_R^3\|s\|_\infty^3. \end{aligned}$$

By the proof of Lemma 3.5, BFV.Relinearize introduces additional noise of at most $\frac{1}{8}\delta_R^2\|s\|_\infty$. So, after performing relinearization in step 3 of Algorithm 16, we have a BFV ciphertext with noise bounded by

$$\frac{30}{8}k_1k_2t\delta_R^3\|s\|_\infty^3 + \frac{1}{8}\delta_R^2\|s\|_\infty = \frac{30}{8}k_1k_2t\delta_R^3\|s\|_\infty^3\left(1 + \frac{1}{30k_1k_2t\delta_R\|s\|_\infty^2}\right)$$

$$\leq \frac{30}{8}k_1k_2t\delta_R^3||s||_\infty^3\left(1+\frac{1}{30\cdot 2\cdot 16}\right)$$
$$\leq \frac{31}{8}k_1k_2t\delta_R^3||s||_\infty^3.$$

So, a worst-case noise bound for a depth-1 multiplication is given by $\frac{31}{8}k_1k_2t\delta_R^3||s||_\infty^3$. Since $\delta_R - 2 \geq \frac{7}{8}\delta_R$ and $\delta_R||s||_\infty - 2 \geq \frac{7}{8}\delta_R||s||_\infty$, we have

$$\frac{2(\frac{31}{8}k_1k_2t\delta_R^3||s||_\infty^3)}{\delta_R||s||_\infty - 2} \leq \frac{\frac{62}{8}k_1k_2t\delta_R^3||s||_\infty^3}{\frac{7}{8}\delta_R||s||_\infty} < 9k_1k_2t\delta_R^2||s||_\infty^2.$$

As $\delta_R = n$ and $||s||_\infty = 1$, we have that $9k_1k_2t\delta_R^2||s||_\infty^2 = 9k_1k_2tn^2 < q_i$. By Lemma 2.2, BFV modulus reduction from Q_i to Q_{i-1} gives a new ciphertext with noise bounded by ρ. Thus, the lemma is proved. □

Proof of Lemma 4.2. Suppose we have a collection of BGV ciphertexts, each with noise bounded by $\rho = \delta_R||s||_\infty$. By a similar argument to Lemma 3.3, computing

$$\mathrm{ct}_j \coloneqq \mathrm{Linearcombo}(\mathrm{ct}_{j,1}, \ldots, \mathrm{ct}_{j,k_1}, 1, \ldots, 1)$$

results in BGV ciphertexts ct_j for $j = 1, \ldots, 2k_2$, each with noise bounded by $k_1\delta_R||s||_\infty + k_1$. Since $\delta_R \geq 16$, we have

$$k_1\delta_R||s||_\infty + k_1 = (k_1 + \frac{k_1}{\delta_R||s||_\infty})\delta_R||s||_\infty \leq \frac{17}{16}k_1\delta_R||s||_\infty.$$

By Lemma 3.7, each $\mathrm{Multiply}(\mathrm{ct}_{2j-1}, \mathrm{ct}_{2j})$ in step 2 of Algorithm 16 for $j = 1, \ldots, k_2$ results in a polynomial triple with noise bounded by

$$2t\delta_R\left(\left(\frac{17}{16}k_1\delta_R||s||_\infty\right)^2 + 1\right) = 2t\delta_R\left(\frac{289}{256}k_1^2\delta_R^2||s||_\infty^2 + 1\right).$$

Then,

$$2t\delta_R\left(\frac{289}{256}k_1^2\delta_R^2||s||_\infty^2 + 1\right) = \frac{289}{128}tk_1^2\delta_R^3||s||_\infty^2\left(1+\frac{256}{289k_1^2\delta_R^2||s||_\infty^2}\right)$$
$$\leq \frac{289}{128}tk_1^2\delta_R^3||s||_\infty^2\left(1+\frac{1}{289}\right)$$
$$= \frac{290}{128}tk_1^2\delta_R^3||s||_\infty^2.$$

Summing all k_2 of these polynomial triples in step 2 of Algorithm 16 results in a polynomial triple with noise bounded by

$$\frac{290}{128}k_1^2k_2t\delta_R^3||s||_\infty^3 + k_2.$$

Notice,

$$\frac{290}{128}k_1^2k_2t\delta_R^3||s||_\infty^3 + k_2 = \frac{290}{128}k_1^2k_2t\delta_R^3||s||_\infty^3\left(1+\frac{128}{290k_1^2t\delta_R^3||s||_\infty^3}\right)$$
$$= \frac{290}{128}k_1^2k_2t\delta_R^3||s||_\infty^3\left(1+\frac{128}{290\cdot 2\cdot 16^3}\right)$$
$$\leq \frac{290}{128}k_1^2k_2t\delta_R^3||s||_\infty^3\left(1+\frac{1}{18560}\right)$$
$$\leq \frac{13}{8}k_1^2k_2t\delta_R^3||s||_\infty^3.$$

By the proof of Lemma 3.8, BGV.Relinearize introduces additional noise of at most $\frac{1}{8}\delta_R^2||s||_\infty^2$. So, after performing relinearization in step 3 of Algorithm 16, we have a BGV ciphertext with noise bounded by

$$\begin{aligned}\frac{13}{8}k_1^2k_2t\delta_R^3||s||_\infty^3 + \frac{1}{8}\delta_R^2||s||_\infty^2 &= \frac{13}{8}k_1^2k_2t\delta_R^3||s||_\infty^3\left(1 + \frac{1}{13k_1^2k_2t\delta_R||s||_\infty^2}\right)\\ &\le \frac{13}{8}k_1^2k_2t\delta_R^3||s||_\infty^3\left(1 + \frac{1}{416}\right)\\ &\le \frac{14}{8}k_1^2k_2t\delta_R^3||s||_\infty^3.\end{aligned}$$

So, a worst-case noise bound for a depth-1 multiplication is given by $\frac{14}{8}k_1^2k_2t\delta_R^3||s||_\infty^3$. Since $\delta_R - 2 \ge \frac{7}{8}\delta_R$ and $\delta_R||s||_\infty - 2 \ge \frac{7}{8}\delta_R||s||_\infty$, we have

$$\frac{2(\frac{14}{8}k_1^2k_2t\delta_R^3||s||_\infty^3)}{\delta_R||s||_\infty - 2} \le \frac{\frac{28}{8}k_1^2k_2t\delta_R^3||s||_\infty^3}{\frac{7}{8}\delta_R||s||_\infty} = 4k_1^2k_2t\delta_R^2||s||_\infty^2.$$

As $\delta_R = n$ and $||s||_\infty = 1$, we have that $4k_1^2k_2t\delta_R^2||s||_\infty^2 = 4k_1^2k_2tn^2 < q_i$. By Lemma 2.3, BGV modulus reduction from Q_i to Q_{i-1} gives a new ciphertext with noise bounded by ρ. Thus, the lemma is proved. □

Proof of Lemma 4.3. Suppose we have a collection of CKKS ciphertexts, each with noise bounded by E. Recall that $\delta_R||s||_\infty = n$. By Lemma 3.11, computing

$$\mathrm{ct}_j \coloneqq \mathrm{Linearcombo}(\mathrm{ct}_{j,1}, \ldots, \mathrm{ct}_{j,k_1}, 1, \ldots, 1)$$

results in CKKS ciphertexts ct_j for $j = 1, \ldots, 2k_2$, each with noise bounded by k_1E. Let $t = 2\Delta Z + 1$, and observe that for each encoding m_j of message z_j,

$$||m_j||_\infty \le \Delta||z_j||_\infty + \frac{1}{2} \le \Delta Z + \frac{1}{2} = t/2.$$

By Lemma 3.12, each $\mathrm{Multiply}(\mathrm{ct}_{2j-1}, \mathrm{ct}_{2j})$ in Step 2 of Algorithm 16 for $j = 1, \ldots, k_2$ results in a polynomial triple with noise bounded by

$$k_1Etn + k_1^2E^2n = k_1E(2\Delta Z + 1)n + k_1^2E^2n.$$

Summing all k_2 of these polynomial triples in Step 2 of Algorithm 16 results in a polynomial triple with noise bounded by

$$k_1k_2E(2\Delta Z + 1)n + k_1^2k_2E^2n$$

by an equivalent argument to Lemma 3.11 with polynomial triples as the input. By the proof of Lemma 3.13, CKKS.Relinearize introduces additional noise of at most $\frac{1}{8}n^2$. So, after performing relinearization in Step 3 of Algorithm 16, we have a CKKS ciphertext with noise bounded by

$$k_1k_2E(2\Delta Z + 1)n + k_1^2k_2E^2n + \frac{1}{8}n^2.$$

So, a worst-case noise bound for a depth-1 multiplication is given by $k_1k_2E(2\Delta Z + 1)n + k_1^2k_2E^2n + \frac{1}{8}n^2$. Performing a rescaling operation then gives noise bounded by

$$\frac{2k_1k_2E\Delta Zn}{\Delta} + \frac{k_1k_2En}{\Delta} + \frac{k_1^2k_2E^2n}{\Delta} + \frac{n^2}{8\Delta} \le 2k_1k_2nEZ + \frac{k_1k_2En}{\Delta} + \frac{k_1^2k_2E^2n}{\Delta} + \frac{1}{8}.$$ □

Proof of Lemma 4.7. By assumption the input of Algorithm 23 is the RNS representation in basis $\mathcal{B}$ of some $(c_0, c_1, c_2) \in R_{n,Q_i}^3$ satisfying

$$c_0 + c_1s + c_2s^2 \equiv D_{Q_i}[m_0m_1]_{\phi(x),t} + e' \bmod (\phi(x), Q_i).$$

for $\|e'\|_\infty \le E$. From Step 1, $(\tilde{c}_2^{(0)}, \ldots, \tilde{c}_2^{(k-1)}, c_2^{(0)}, \ldots, c_2^{(i)})$ is the RNS representation of $\tilde{c}_2 \in R_{n,PQ_i}$ in basis $\mathcal{D}$ satisfying $\tilde{c}_2 = c_2 + Q_i e$ for some $e \in R_n$ with $\|\tilde{c}_2\|_\infty \le Q_i(i+1)/2$ by Lemma 4.4. After Step 3, note $(\hat{a}^{(j)}, \hat{b}^{(j)})_{0\le j\le i+k}$ is the RNS representation of some $(\hat{a}, \hat{b}) \in R^2_{n,PQ_i}$ satisfying

$$\begin{aligned}
\hat{b} + \hat{a}s &\equiv \tilde{c}_2\tilde{k}_1 + \tilde{c}_2\tilde{k}_0 s \bmod (\phi(x), PQ_i)\\
&\equiv \tilde{c}_2(-\tilde{k}_0 s + Ps^2 + \tilde{e}) + \tilde{c}_2\tilde{k}_0 s \bmod (\phi(x), PQ_i)\\
&\equiv (c_2 + Q_i e)(-\tilde{k}_0 s + Ps^2 + \tilde{e}) + (c_2 + Q_i e)\tilde{k}_0 s \bmod (\phi(x), PQ_i)\\
&\equiv (c_2 + Q_i e)(Ps^2 + \tilde{e}) \bmod (\phi(x), PQ_i)\\
&\equiv c_2 Ps^2 + c_2\tilde{e} + PQ_i e s^2 + Q_i e\tilde{e} \bmod (\phi(x), PQ_i)\\
&\equiv c_2 Ps^2 + c_2\tilde{e} + Q_i e\tilde{e} \bmod (\phi(x), PQ_i)\\
&\equiv c_2 Ps^2 + \hat{e} \bmod (\phi(x), PQ_i)
\end{aligned}$$

for $\hat{e} = c_2\tilde{e} + Q_i e\tilde{e} = \tilde{c}_2\tilde{e}$. Note that as $\tilde{e} \leftarrow \chi_\rho$, we have $\|\hat{e}\|_\infty = \|\tilde{c}_2\tilde{e}\|_\infty \le Q_i(i+1)\rho^2/2$. Furthermore, there exists $\omega \in R_n$ such that

$$\hat{b} + \hat{a}s \equiv c_2 Ps^2 + \hat{e} + \omega PQ_i \bmod \phi(x).$$

By Lemma 4.5, Step 4 returns the RNS representation of some $\hat{c}_0 \in R_{n,Q_i}$ and $\hat{c}_1 \in R_{n,Q_i}$ satisfying

$$\begin{aligned}
\hat{c}_0 &= \frac{\hat{b}}{P} + \hat{e}_0,\\
\hat{c}_1 &= \frac{\hat{a}}{P} + \hat{e}_1
\end{aligned}$$

with $\|\hat{e}_0\|_\infty \le k/2$ and $\|\hat{e}_1\|_\infty \le k/2$. Finally, Step 6 returns the RNS representation of $(a, b) \in R^2_{n,Q_i}$, which satisfies

$$\begin{aligned}
b + as &\equiv c_0 + \hat{c}_0 + (c_1 + \hat{c}_1)s \bmod (\phi(x), Q_i)\\
&\equiv c_0 + c_1 s + P^{-1}(\hat{b} + \hat{a}s) + \hat{e}_0 + \hat{e}_1 s \bmod (\phi(x), Q_i)\\
&\equiv c_0 + c_1 s + P^{-1}(c_2 Ps^2 + \hat{e} + \omega PQ_i) + \hat{e}_0 + \hat{e}_1 s \bmod (\phi(x), Q_i)\\
&\equiv c_0 + c_1 s + c_2 s^2 + P^{-1}\hat{e} + \omega Q_i + \hat{e}_0 + \hat{e}_1 s \bmod (\phi(x), Q_i)\\
&\equiv D_{Q_i}[m_0 m_1]_{\phi(x),t} + e' + P^{-1}\hat{e} + \hat{e}_0 + \hat{e}_1 s \bmod (\phi(x), Q_i)\\
&\equiv D_{Q_i}[m_0 m_1]_{\phi(x),t} + e^* \bmod (\phi(x), Q_i)
\end{aligned}$$

for $e^* = e' + P^{-1}\hat{e} + \hat{e}_0 + \hat{e}_1 s$. We now turn to the noise term e^*. If $P \ge 6Q_i$, $\delta_R \ge 16$, and $k > i$, then

$$\begin{aligned}
\|e^*\|_\infty &\le \|e'\|_\infty + P^{-1}\|\hat{e}\|_\infty + \|\hat{d}_0 s\|_\infty + \|\hat{d}_1\|_\infty\\
&\le \|e'\|_\infty + \frac{Q_i(i+1)}{2P}\delta_R^2\|s\|_\infty + \frac{k}{2} + \frac{k\delta_R\|s\|_\infty}{2}\\
&\le \|e'\|_\infty + \delta_R^2 k\left(\frac{1}{12} + \frac{1}{\delta_R^2} + \frac{1}{2\delta_R}\right)\\
&\le E + \frac{1}{8}\delta_R^2 k.
\end{aligned}$$

□

Proof of Lemma 4.8. By assumption the input of Algorithm 23 is the RNS representation in basis $\mathcal{B}$ of some $(c_0, c_1, c_2) \in R^3_{n,Q_i}$ satisfying

$$c_0 + c_1 s + c_2 s^2 \equiv [m_0 m_1]_{\phi(x),t} + te' \bmod (\phi(x), Q_i).$$

for $\|e'\|_\infty \le E$. From Step 1, $(\tilde{c}_2^{(0)}, \ldots, \tilde{c}_2^{(k-1)}, c_2^{(0)}, \ldots, c_2^{(i)})$ is the RNS representation of $\tilde{c}_2 \in R_{n,PQ_i}$ in basis $\mathcal{D}$ satisfying $\tilde{c}_2 = c_2 + Q_i e$ for some $e \in R_n$ with $\|\tilde{c}_2\|_\infty \le Q_i(i+1)/2$ by Lemma 4.4. After Step 3, note $(\hat{a}^{(j)}, \hat{b}^{(j)})_{0\le j\le i+k}$

is the RNS representation of some $(\hat{a}, \hat{b}) \in R_{n,PQ_i}^2$ satisfying

$$\begin{aligned}
\hat{b} + \hat{a}s &\equiv \tilde{c}_2\tilde{k}_1 + \tilde{c}_2\tilde{k}_0 s \bmod (\phi(x), PQ_i) \\
&\equiv \tilde{c}_2(-\tilde{k}_0 s + Ps^2 + t\tilde{e}) + \tilde{c}_2\tilde{k}_0 s \bmod (\phi(x), PQ_i) \\
&\equiv (c_2 + Q_i e)(-\tilde{k}_0 s + Ps^2 + t\tilde{e}) + (c_2 + Q_i e)\tilde{k}_0 s \bmod (\phi(x), PQ_i) \\
&\equiv (c_2 + Q_i e)(Ps^2 + t\tilde{e}) \bmod (\phi(x), PQ_i) \\
&\equiv c_2 Ps^2 + tc_2\tilde{e} + PQ_i e s^2 + tQ_i e\tilde{e} \bmod (\phi(x), PQ_i) \\
&\equiv c_2 Ps^2 + t(c_2\tilde{e} + Q_i e\tilde{e}) \bmod (\phi(x), PQ_i) \\
&\equiv c_2 Ps^2 + t\hat{e} \bmod (\phi(x), PQ_i)
\end{aligned}$$

for $\hat{e} = c_2\tilde{e} + Q_i e\tilde{e} = \tilde{c}_2\tilde{e}$. Note that as $\tilde{e} \leftarrow \chi_\rho$, we have $\|\hat{e}\|_\infty = \|\tilde{c}_2\tilde{e}\|_\infty \le Q_i(i+1)\rho^2/2$. Furthermore, there exists $\omega \in R_n$ such that

$$\hat{b} + \hat{a}s \equiv c_2 Ps^2 + t\hat{e} + \omega PQ_i \bmod \phi(x).$$

By Lemma 4.6, Step 4 returns the RNS representation of some $\hat{c}_0 \in R_{n,Q_i}$ and $\hat{c}_1 \in R_{n,Q_i}$ satisfying

$$\hat{c}_0 = \frac{\hat{b}}{P} + t \cdot \hat{e}_0,$$
$$\hat{c}_1 = \frac{\hat{a}}{P} + t \cdot \hat{e}_1$$

with $\|\hat{e}_0\|_\infty \le k/2$ and $\|\hat{e}_1\|_\infty \le k/2$. Finally, Step 6 returns the RNS representation of $(a, b) \in R_{n,Q_i}^2$, which satisfies

$$\begin{aligned}
b + as &\equiv c_0 + \hat{c}_0 + (c_1 + \hat{c}_1)s \bmod (\phi(x), Q_i) \\
&\equiv c_0 + c_1 s + P^{-1}(\hat{b} + \hat{a}s) + t\hat{e}_0 + t\hat{e}_1 s \bmod (\phi(x), Q_i) \\
&\equiv c_0 + c_1 s + P^{-1}(c_2 Ps^2 + t\hat{e} + \omega PQ_i) + t\hat{e}_0 + t\hat{e}_1 s \bmod (\phi(x), Q_i) \\
&\equiv c_0 + c_1 s + c_2 s^2 + P^{-1}t\hat{e} + \omega Q_i + t\hat{e}_0 + t\hat{e}_1 s \bmod (\phi(x), Q_i) \\
&\equiv [m_0 m_1]_{\phi(x),t} + te' + P^{-1}t\hat{e} + t\hat{e}_0 + t\hat{e}_1 s \bmod (\phi(x), Q_i) \\
&\equiv [m_0 m_1]_{\phi(x),t} + te^* \bmod (\phi(x), Q_i)
\end{aligned}$$

for $e^* = e' + P^{-1}\hat{e} + \hat{e}_0 + \hat{e}_1 s$. The bound on e^* follows identically as in the proof of Lemma 4.7. □

Proof of Lemma 4.9. By assumption the input of Algorithm 23 is the RNS representation in basis $\mathcal{B}$ of some $(c_0, c_1, c_2) \in R_{n,Q_i}^3$ satisfying

$$c_0 + c_1 s + c_2 s^2 \equiv m_0 m_1 + e' \bmod (\phi(x), Q_i).$$

for $\|e'\|_\infty \le E$. By an identical argument to the proof of Lemma 4.7, the output of Algorithm 23 is the RNS representation in basis $\mathcal{B}$ of $(a, b) \in R_{n,Q_i}$ satisfying

$$b + as \equiv m_0 m_1 + e^*$$

with $\|e^*\|_\infty \le E + \frac{1}{8}\delta_R^2 k$. □