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Subject: Lattice Group Key Exchange
Date: Wednesday, April 17, 2019 4:57:17 PM
Attachments: GKEM.pdf

Hi Daniel,

--Daniel

Constant-Round Group Key Exchange from the Ring-LWE Assumption

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Abstract. Group key-exchange protocols allow a set of N parties to agree on a shared, secret key by communicating over a public network. A number of solutions to this problem have been proposed over the years, mostly based on variants of Diffie-Hellman (two-party) key exchange. To the best of our knowledge, however, there has been almost no work looking at candidate *post-quantum* group key-exchange protocols.

Here, we propose a constant-round protocol for unauthenticated group key exchange (i.e., with security against a passive eavesdropper) based on the hardness of the Ring-LWE problem. By applying the Katz-Yung compiler using any post-quantum signature scheme, we obtain a (scalable) protocol for *authenticated* group key exchange with post-quantum security. Our protocol is constructed by generalizing the Burmester-Desmedt protocol to the Ring-LWE setting, which requires addressing several technical challenges.

Keywords: Ring learning with errors, Post-quantum cryptography, Group key exchange

1 Introduction

Protocols for (authenticated) key exchange are among the most fundamental and widely used cryptographic primitives. They allow parties communicating over an insecure public network to establish a common secret key, called a *session key*, permitting the subsequent use of symmetric-key cryptography for encryption and authentication of sensitive data. They can be used to instantiate so-called "secure channels" upon which higher-level cryptographic protocols often depend.

Most work on key exchange, beginning with the classical paper of Diffie and Hellman, has focused on two-party key exchange. However, many works have also explored extensions to the *group* setting [21, 29, 15, 30, 5, 6, 25, 14, 12, 13, 11, 17, 22, 16, 8, 2, 1, 24, 9, 31] in which N parties wish to agree on a common session key that they can each then use for encrypted/authenticated communication with the rest of the group.

The recent effort by NIST to evaluate and standardize one or more quantum-resistant public-key cryptosystems is entirely focused on digital signatures and two-party key encapsulation/key exchange,[1] and there has been an extensive amount of research over the past decade focused on designing such schemes. In contrast, we are aware of almost no[2] work on *group* key-exchange protocols with post-quantum security beyond the observation that a post-quantum group key-exchange protocol can be constructed from any post-quantum two-party protocol by having a designated group manager run independent two-party protocols with the $N - 1$ other parties, and then send a session key of its choice to the other parties encrypted/authenticated using each of the resulting keys. Such a solution is often considered unacceptable since it is highly asymmetric, requires additional coordination, is not contributory, and puts a heavy load on a single party who becomes a central point of failure.

1.1 Our Contributions

In this work, we propose a constant-round group key-exchange protocol based on the hardness of the Ring-LWE problem [27], and hence with (plausible) post-quantum security. We focus on constructing an *unauthenticated* protocol—i.e., one secure against a passive eavesdropper—since known techniques such as the Katz-Yung compiler [24] can then be applied to obtain an *authenticated* protocol secure against an active attacker.

The starting point for our work is the two-round group key-exchange protocol by Burmester and Desmedt [15, 16, 24], which is based on the decisional Diffie-Hellman assumption. Assume a group $\mathbb{G}$ of prime order q and a generator $g \in \mathbb{G}$ are fixed and public. The Burmester-Desmedt protocol run by parties $P_0, \ldots, P_{N-1}$ then works as follows:

1. In the first round, each party P_i chooses uniform $r_i \in \mathbb{Z}_q$ and broadcasts $z_i = g^{r_i}$ to all other parties.
2. In the second round, each party P_i broadcasts $X_i = (z_{i+1}/z_{i-1})^{r_i}$ (where the parties' indices are taken modulo N).

Each party P_i can then compute its session key sk_i as

$$\mathsf{sk}_i = (z_{i-1})^{Nr_i} \cdot X_i^{N-1} \cdot X_{i+1}^{N-2} \cdots X_{i+N-2}.$$

One can check that all the keys are equal to the same value $g^{r_0 r_1 + \cdots + r_{N-1} r_0}$.

In attempting to adapt their protocol to the Ring-LWE setting, we could fix a ring R_q and a uniform element $a \in R_q$. Then:

1. In the first round, each party P_i chooses "small" secret value $s_i \in R_q$ and "small" noise term $e_i \in R_q$ (with the exact distribution being unimportant in the present discussion), and broadcasts $z_i = as_i + e_i$ to the other parties.

[1] Note that CPA-secure key encapsulation is equivalent to two-round key-exchange (with passive security).

[2] The protocol of Ding et al. [19] has no security proof; the work of Boneh et al. [10] shows a framework for constructing a group key-exchange protocol with plausible post-quantum security but without a concrete instantiation.

2. In the second round, each party P_i chooses a second "small" noise term $e'_i \in R_q$ and broadcasts $X_i = (z_{i+1} - z_{i-1}) \cdot s_i + e'_i$.

Each party can then compute a session key b_i as

$$b_i = N \cdot s_i \cdot z_{i-1} + (N-1) \cdot X_i + (N-2) \cdot X_{i+1} + \cdots + X_{i+N-2}.$$

The problem, of course, is that (due to the noise terms) these session keys computed by the parties will *not* be equal. They will, however, be "close" to each other if the $\{s_i, e_i, e'_i\}$ are all sufficiently small, so we can add an additional reconciliation step to ensure that all parties agree on a common key k.

This gives a protocol that is correct, but proving security (even for a passive eavesdropper) is more difficult than in the case of the Burmester-Desmedt protocol. Here we informally outline the main difficulties and how we address them. First, we note that trying to prove security by direct analogy to the proof of security for the Burmester-Desmedt protocol (cf. [24]) fails; in the latter case, it is possible to use the fact that, for example,

$$(z_2/z_0)^{r_1} = z_1^{r_2 - r_0},$$

whereas in our setting the analogous relation does not hold. In general, the natural proof strategy here is to switch all the $\{z_i\}$ values to uniform elements of R_q, and similarly to switch the $\{X_i\}$ values to uniform subject to the constraint that their sum is approximately 0 (i.e., subject to the constraint that $\sum_i X_i \approx 0$). Unfortunately this cannot be done by simply invoking the Ring-LWE assumption $O(N)$ times; in particular, the first time we try to invoke the assumption, say on the pair $(z_1 = as_1 + e_1,\ X_1 = (z_2 - z_0)\ s_1 + e'_1)$, we need $z_2 - z_0$ to be uniform—which, in contrast to the analogous requirement in the Burmester-Desmedt protocol (for the value z_2/z_0), is not the case here. Thus, we must somehow break the circularity in the mutual dependence of the $\{z_i, X_i\}$ values.

Toward this end, let us look more carefully at the distribution of $\sum_i X_i$. We may write

$$\sum_i X_i = \sum_i (e_{i+1} s_i - e_{i-1} s_i) + \sum_i e'_i.$$

Consider now changing the way X_0 is chosen: that is, instead of choosing $X_0 = (z_1 - z_{N-1}) s_0 + e'_0$ as in the protocol, we instead set $X_0 = -\sum_{i=1}^{N-1} X_i + e'_0$ (where e'_0 is from the same distribution as before). Intuitively, as long as the standard deviation of e'_0 is large enough, these two distributions of X_0 should be "close" (as they both satisfy $\sum_i X_i \approx 0$). This, in particular, means that we need the distribution of e'_0 to be different from the distribution of the $\{e'_i\}_{i>0}$, as the standard deviation of the former needs to be larger than the latter.

We can indeed show that when we choose e'_0 from an appropriate distribution then the Rényi divergence between the two distributions of X_0, above, is bounded by a polynomial. With this switch in the distribution of X_0, we have broken the circularity and can now use the Ring-LWE assumption to switch the distribution of z_0 to uniform, followed by the remaining $\{z_i, X_i\}$ values.

Unfortunately, bounded Rényi divergence does not imply statistical closeness. However, polynomially bounded Rényi divergence *does* imply that any event occurring with negligible probability when X_0 is chosen according to the second distribution also occurs with negligible probability when X_0 is chosen according to the first distribution. For these reasons, we change our security goal from an "indistinguishability-based" one (namely, requiring that, given the transcript, the real session key is indistinguishable from uniform) to an "unpredictability-based" one (namely, given the transcript, it should be infeasible to compute the real session key). In the end, though, once the parties agree on an unpredictable value k they can hash it to obtain the final session key $\mathsf{sk} = \mathsf{H}(k)$; this final value sk will be indistinguishable from uniform if H is modeled as a random oracle.

2 Preliminaries

2.1 Notation

Let $\mathbb{Z}$ be the ring of integers, and let $[N] = \{0, 1, \ldots, N-1\}$. If χ is a probability distribution over some set S, then $x_0, x_1, \ldots, x_{\ell-1} \leftarrow \chi$ denotes independently sampling each x_i from distribution χ. We let $\mathrm{Supp}(\chi) = \{x : \chi(x) \neq 0\}$. Given an event E, we use E to denote its complement. Let $\chi(E)$ denote the probability that event E occurs under distribution χ. Given a polynomial p_i, let $(p_i)_j$ denote the jth coefficient of p_i. Let $\log(X)$ denote $\log_2(X)$, and $\exp(X)$ denote e^X. $\mathrm{poly}(\lambda)$ denotes a polynomial in term of λ.

2.2 Ring Learning with Errors

Informally, the (decisional) version of the Ring Learning with Errors (Ring-LWE) problem is: for some secret ring element s, distinguish many random "noisy ring products" with s from elements drawn uniformly from the ring. More precisely, the Ring-LWE problem is parameterized by (R, q, χ, ℓ) as follows:

1. $R = \mathbb{Z}[X]/(f(X))$ is a ring for some irreducible polynomial $f(X)$ in the indeterminate X. In this paper, we restrict to the case of $f(X) = X^n + 1$ where n is a power of 2. In later sections, we let R be parameterized by n.
2. q is a modulus defining the quotient ring $R_q := R/qR = \mathbb{Z}_q[X]/(f(X))$. We restrict to the case that q is prime and $q = 1 \bmod 2n$.
3. $\chi = (\chi_s, \chi_e)$ is a pair of noise distributions over R_q (with χ_s the *secret key* distribution and χ_e the *error* distribution) that are concentrated on "short" elements, for an appropriate definition of "short."
4. ℓ is the number of samples provided to the adversary.

Formally, the Ring-LWE problem is to distinguish between ℓ samples independently drawn from one of two distributions. The first distribution is generated by choosing secret $s \leftarrow \chi_s$ and then outputting

$$(a_i, b_i = s \cdot a_i + e_i) \in R_q \times R_q$$

for $i \in [\ell]$, where each a_i is uniform in R_q and each $e_i \leftarrow \chi_e$ is drawn from the error distribution. In the second distribution, each sample (a_i, b_i) is simply uniform in $R_q \times R_q$.

Let A_{n,q,χ_s,χ_e} be the distribution that outputs the Ring-LWE sample $(a_i, b_i = s \cdot a_i + e_i)$ as above. We denote by $\mathsf{Adv}^{\mathsf{RLWE}}_{n,q,\chi_s,\chi_e,f}(\mathsf{B})$ the advantage of algorithm B in distinguishing distributions A_{n,q,χ_s,χ_e} and $U(R_q^2)$.

We define $\mathsf{Adv}^{\mathsf{RLWE}}_{n,q,\chi_s,\chi_e,f}(t)$ to be the maximum advantage of any adversary running in time t. Note that in later sections, we write $\mathsf{Adv}_{n,q,\chi,f}$ if $\chi = \chi_s = \chi_e$ for simplicity.

The Ring-LWE Noise Distribution. The noise distribution χ (here we assume $\chi_s = \chi_e$, though this is not necessary) is usually a discrete Gaussian distribution on $R_q^\vee$ or in our case R_q (see [18] for details of the distinction, especially for concrete implementation purposes). Formally, in case of power of two cyclotomic rings, the discrete Gaussian distribution can be sampled by drawing each coefficient independently from the 1-dimensional discrete Gaussian distribution over $\mathbb{Z}$ with parameter σ, which is supported on $\{x \in \mathbb{Z} : -q/2 \leq x \leq q/2\}$ and has density function

$$D_{\mathbb{Z}_q,\sigma}(x) = \frac{e^{\frac{-\pi x^2}{\sigma^2}}}{\sum_{x=-\infty}^{\infty} e^{\frac{-\pi x^2}{\sigma^2}}}.$$

2.3 Rényi divergence

The Rényi divergence (RD) is a measure of closeness for two probability distributions. For any two discrete probability distributions P and Q such that $\mathrm{Supp}(P) \subseteq \mathrm{Supp}(Q)$, we define

$$\mathrm{RD}_2(P||Q) = \sum_{x\in\mathrm{Supp}(P)} \frac{P(x)^2}{Q(x)}.$$

Rényi divergence has a probability preservation property that can be considered the multiplicative analogue of statistical distance.

Proposition 1. *Given discrete distributions P and Q with $\mathrm{Supp}(P) \subseteq \mathrm{Supp}(Q)$, let $E \subseteq \mathrm{Supp}(Q)$ be an arbitrary event. We have*

$$Q(E) \geq P(E)^2/\mathrm{RD}_2(P||Q).$$

This property implies that as long as $\mathrm{RD}_2(P||Q)$ is bounded by $\mathrm{poly}(\lambda)$, any event E that occurs with negligible probability $Q(E)$ under distribution Q also occurs with negligible probability $P(E)$ under distribution P. We refer to [27, 26] for the formal proof.

The following theorem bounds the Rényi divergence between Gaussian distributions, which allows the "noise flooding" technique to be used even with polynomial modulus q.

Theorem 2.1 ([7]). *Fix $m, q, \lambda \in \mathbb{Z}$, a bound B, and the 1-dimensional discrete Gaussian distribution $D_{\mathbb{Z}_q,\sigma}$ such that $B < \sigma < q$. Moreover, let $e \in \mathbb{Z}$ be such that $|e| \leq B$. If $\sigma = \Omega(B\sqrt{m/\log\lambda})$, then*

$$\mathrm{RD}_2((e + D_{\mathbb{Z}_q,\sigma})^m||D^m_{\mathbb{Z}_q,\sigma}) \leq \exp(2\pi m(B/\sigma)^2) = \mathrm{poly}(\lambda),$$

where X^m denotes m independent samples from X.

2.4 Generic Key Reconciliation

In this subsection, we define a generic, one round, two-party key reconciliation mechanism which allows both parties to derive the same key from an approximately agreed upon ring element. A key reconciliation mechanism KeyRec consists of two algorithms recMsg and recKey, parameterized by security parameter 1^λ as well as β_{Rec}. In this context, Alice and Bob hold "close" values b_A and b_B, respectively, and wish to generate a shared value k. The abstract mechanism KeyRec is defined as follows:

1. Bob computes $\mathsf{recMsg}(b_B)$ which outputs a reconciliation message m^{rec} and a final key k_B. Bob sends the reconciliation message m^{rec} to Alice.
2. Once receiving m^{rec}, Alice computes $\mathsf{recKey}(b_A, m^{\mathsf{rec}})$, which outputs a final key $k_A =\in \{0, 1\}^\lambda$.

Correctness. Given $b_A, b_B \in R_q$, if each coefficient of $b_B - b_A$ is bounded by β_{Rec} then it is guaranteed that $k_A = k_B$.

Security. A key reconciliation mechanism KeyRec is secure if the subsequent two distribution ensembles are computationally indistinguishable.
$\mathsf{Exe}_{\mathsf{KeyRec}}(\lambda)$: A draw from this helper distribution is performed by initiating the key reconciliation protocol among two honest parties and outputting (m^{rec}, k_B); i.e. the reconciliation message m^{rec} and (Bob's) key k_B of the protocol execution.

We denote by $\mathsf{Adv}_{\mathsf{KeyRec}}(\mathcal{B})$ the advantage of adversary distinguishing the distributions below.

$$\{(m^{\mathsf{rec}}, k_B) \mid b_B \leftarrow \mathsf{U}(R_q), (m^{\mathsf{rec}}, k_B) \leftarrow \mathsf{Exe}_{\mathsf{KeyRec}}(\lambda, b_B)\}_{\lambda\in\mathbb{N}},$$
$$\{(m^{\mathsf{rec}}, k') \mid b_B \leftarrow \mathsf{U}(R_q), (m^{\mathsf{rec}}, k_B) \leftarrow \mathsf{Exe}_{\mathsf{KeyRec}}(\lambda, b_B), k' \leftarrow U_\lambda\}_{\lambda\in\mathbb{N}},$$

where U_λ denotes the uniform distribution over λ bits.

We define $\mathsf{Adv}_{\mathsf{KeyRec}}(t)$ to be the maximum advantage of any adversary running in time t.

Key reconciliation mechanisms from the literature. The notion of key reconciliation was first introduced by Ding et al. [19] in his work on two-party, lattice-based key exchange. It was later used in several works on two-party key exchange, including [28, 32, 4].

In the key reconciliation mechanisms of Peikert [28], Zhang et al. [32] and Alkim et al. [4], the agreed-upon key $k = k_A = k_B$ is close to each of the original values b_A, b_B held by the parties. When instantiating our group key exchange (GKE) protocol with this type of key-reconciliation mechanism, our final GKE protocol is contributory. In other cases [3], the agreed-upon key is determined by Bob; instantiating our GKE protocol with this type of key-reconciliation mechanism yields a non-contributory protocol.

3 Group Key Exchange

A group key-exchange protocol allows a session key to be established among $N > 2$ parties. Following prior work [23, 14, 12, 13], we will use the term group key exchange (GKE) to denote a protocol secure against a *passive* (eavesdropping) adversary and will use the term authenticated group key exchange (GAKE) to denote a protocol secure against an *active* adversary, who controls all communication channels. Fortunately, the work of Katz and Yung [23] presents a compiler that takes any GKE protocol and transforms it into a GAKE protocol. The underlying tool required for this transform is any post-quantum signature scheme which is strongly unforgeable under adaptive chosen message attack (EUF-CMA). We may thus focus our attention on achieving GKE in the remainder of this work.

In GKE setting, the adversary gets to see a single transcript generated by an execution of the protocol. Given the transcript, the adversary must distinguish the real key from a fake key that is generated uniformly at random and independently of the transcript.

Formally, for security parameter $\lambda \in \mathbb{N}$, we define the following distribution:

$\mathsf{Execute}_{\Pi}^{\mathsf{O}_H}(\lambda)$: A draw from this distribution is performed by sampling a classical random oracle H from distribution H, initiating the GKE protocol Π among N honest parties with security parameter λ relative to H, and outputting $(\mathsf{trans}, \mathsf{sk})$—the transcript trans and key sk of the protocol execution.

Consider the following distributions:

$$\{(\mathsf{trans}, \mathsf{sk}) \mid (\mathsf{trans}, \mathsf{sk}) \leftarrow \mathsf{Execute}_{\Pi}^{\mathsf{O}_H}(\lambda)\}_{\lambda \in \mathbb{N}},$$
$$\{(\mathsf{trans}, \mathsf{sk}') \mid (\mathsf{trans}, \mathsf{sk}) \leftarrow \mathsf{Execute}_{\Pi}^{\mathsf{O}_H}(\lambda), \mathsf{sk}' \leftarrow U_\lambda\}_{\lambda \in \mathbb{N}},$$

where U_λ denotes the uniform distribution over λ bits. Let $\mathsf{Adv}^{\mathsf{GKE},\mathsf{O}_H}(\mathsf{A})$ denote the advantage of adversary A with classical access to the sampled oracle , distinguishing the distributions above.

To enable a concrete security analysis, we define $\mathsf{Adv}^{\mathsf{GKE},\mathsf{O}_H}(t, q_{\mathsf{O}_H})$ to be the maximum advantage of any adversary running in time t and making at most q_{O_H} queries to the random oracle. Security holds even if the adversary sees multiple executions by a hybrid argument.

In the next section we will define our GKE scheme and prove that it satisfies the notion of GKE.

4 A Group Key-Exchange Protocol

In this section, we present our group key exchange construction, Π, which runs key reconciliation protocol KeyRec as a subroutine. Let KeyRec be parametrized by β_{Rec}. The protocol has two security parameters λ and ρ. λ is the computational security parameter. ρ is the statistical parameter. In this setting, N players $P_0, \ldots, P_{N-1}$ plan to generate a shared session key. The players' indices are taken modulo N.

The structure of the protocol is as follows: All parties agree on "close" keys $b_0 \approx \cdots \approx b_{N-1}$ after the second round. Player $N-1$ then initiates a key reconciliation protocol to allow all users to agree on the same key $k = k_0 = \ldots = k_{N-1}$. Since we are only able to prove that k is difficult to compute for an eavesdropping adversary (but may not be indistinguishable from random), we hash k using random oracle H to get the final shared key sk.

Public setting: $R_q = \mathbb{Z}_q[x]/(x^n+1)$, $a \leftarrow \mathsf{U}(R_q)$.

Round 1: Each player P_i samples $s_i, e_i \leftarrow \chi_{\sigma_1}$ and broadcasts $z_i = as_i + e_i$.
Round 2: Player P_0 samples $e'_0 \leftarrow \chi_{\sigma_2}$ and each of the other players P_i samples $e'_i \leftarrow \chi_{\sigma_1}$. Each P_i broadcasts $X_i = (z_{i+1} - z_{i-1})s_i + e'_i$.
Key Computation (Round 3):

- Player P_{N-1} proceeds as follows:
  1. Samples $e'_{N-1} \leftarrow \chi_{\sigma_1}$ and computes $b_{N-1} = z_{N-2}Ns_{N-1} + X_{N-1} \cdot (N-1) + X_0 \cdot (N-2) + \cdots + X_{N-3} + e'_{N-1}$.
  2. Computes $(m^{\mathsf{rec}}_{N-1}, k_{N-1}) = \mathsf{recMsg}(b_{N-1})$ and broadcasts m^{rec}_{N-1}.
  3. Obtains session key $\mathsf{sk}_{N-1} = \mathsf{H}(k_{N-1})$.
- Each player P_i (except P_{N-1}) proceeds as follows:
  1. Computes $b_i = z_{i-1}Ns_i + X_i \cdot (N-1) + X_{i+1} \cdot (N-2) + \cdots + X_{i+N-2}$.
  2. Computes $k_i = \mathsf{recKey}(b_i, m^{\mathsf{rec}}_{N-1})$, and obtains session key $\mathsf{sk}_i = \mathsf{H}(k_i)$.

4.1 Correctness

The following claim states that each party derives the same session key sk_i, with all but negligible probability, as long as $\chi_{\sigma_1}, \chi_{\sigma_2}$ satisfy the constraint $(N^2+2N)\cdot\sqrt{n}\rho^{3/2}\sigma_1^2 + (\frac{N^2}{2}+1)\sigma_1 + (N-2)\sigma_2 \le \beta^{\mathsf{Rec}}$, where β^{Rec} is the parameter from the KeyRec protocol.

Theorem 4.1. *If the parameters in the group key exchange protocol Π satisfy the constraints $(N^2+2N)\cdot\sqrt{n}\rho^{3/2}\sigma_1^2 + (\frac{N^2}{2}+1)\sigma_1 + (N-2)\sigma_2 \le \beta_{\mathsf{Rec}}$, then each player derives the same key with probability at least $1 - 2\cdot 2^{-\rho}$.*

Proof. We refer to Appendix A for the detailed proof. □

5 Security Proof

The following theorem shows that protocol Π is a passively secure group key-exchange protocol. We remark that we prove security of the protocol for a classical attacker only; in particular, we allow the attacker only classical access to H We believe the protocol can be proven secure even against attackers that are allowed to make quantum queries to H, but leave proving this to future work.

Theorem 5.1. *If the parameters in the group key exchange protocol Π satisfy the constraints $2N\sqrt{n}\lambda^{3/2}\sigma_1^2 + (N-1)\sigma_1 \le \beta_{Rényi}$ and $\sigma_2 = \Omega(\beta_{Rényi}\sqrt{n/\log\lambda})$,* *and if* H *is modeled as a random oracle, then for any algorithm running in time t, making at most* q *queries to the random oracle, we have:*

$$\mathsf{Adv}^{\mathsf{GKE},\mathcal{O}^H}_{\Pi}(t,\mathsf{q}) \le 2^{-\lambda+1} + \frac{\left(N\cdot \mathsf{Adv}^{\mathsf{RLWE}}_{n,q,\chi_{\sigma_1},3}(t_1) + \mathsf{Adv}_{\mathsf{KeyRec}}(t_2) + \frac{\mathsf{q}}{2^{\lambda}}\right)\cdot \exp\left(2\pi n(\beta_{R\acute{e}nyi}/\sigma_2)^2\right)}{1-2^{-\lambda+1}},$$

where $t_1 = t + O(N)\cdot t_{\mathsf{ring}}$, $t_2 = t + O(N)\cdot t_{\mathsf{ring}}$ *and where* t_{ring} *is defined as the (maximum) time required to perform operations in* R_q*.*

Proof. Consider the joint distribution of (T,sk), where $\mathsf{T} = (\{z_i\},\{X_i\}, m^{\mathsf{rec}}_{N-1})$ is the transcript of an execution of the protocol Π, and sk is the final shared session key. The distribution of (T,sk) is denoted as Real. Proceeding via a sequence of experiments, we will show that under the Ring-LWE assumption, an adversary having negligible success probability in guessing k_{N-1} as input to the random oracle in the Ideal experiment (to be formally defined) also has negligible success probability in the Real experiment.

Furthermore, in Ideal, the input k_{N-1} to the random oracle is uniformly random, which means that the adversary has $\mathsf{negl}(\lambda)$ probability of guessing k_{N-1} in Ideal when $\mathsf{q} = \text{poly}(\lambda)$. Finally, we argue that the above is sufficient to prove the GKE security of the scheme, because in the random oracle model, the output of the random oracle on k_{N-1} – i.e. the agreed upon key – looks uniformly random to an adversary who does not query k_{N-1}. We now proceed with the formal proof.

Let Query be the event that k_{N-1} is among the adversary $\mathcal{A}$'s random oracle queries and denote by $\Pr_i[\mathsf{Query}]$ the probability that event Query happens in *Experiment i*.

Experiment 0. This is the original experiment. In this experiment, the distribution of (T,sk) is as follows, denoted Real :

$$\mathsf{Real} := \left\{ \begin{array}{l} a \leftarrow R_q;\ \forall i: s_i, e_i \leftarrow \chi_{\sigma_1};\\ \forall i: z_i = a s_i + e_i;\\ e'_1,\ldots,e'_{N-1} \leftarrow \chi_{\sigma_1}; e'_0 \leftarrow \chi_{\sigma_2};\\ \forall i: X_i = (z_{i+1} - z_{i-1})s_i + e'_i;\\ e''_{N-1} \leftarrow \chi_{\sigma_1};\\ b_{N-1} = z_{N-2} N s_{N-1} + e''_{N-1} + X_{N-1}\cdot(N-1)+\\ \quad X_0\cdot(N-2)+\cdots+X_{N-3};\\ (m^{\mathsf{rec}}_{N-1}, k_{N-1}) = \mathsf{recMsg}(b_{N-1});\ \mathsf{sk} = \mathsf{H}(k_{N-1});\\ \mathsf{T} = (z_0,\ldots,z_{N-1},X_0,\ldots,X_{N-1},m^{\mathsf{rec}}_{N-1}) \end{array} : (\mathsf{T},\mathsf{sk}) \right\}.$$

Since $\mathsf{Adv}^{\mathsf{GKE},\mathcal{O}^H}_{\Pi}(t,\mathsf{q}) + \frac{1}{2} = \Pr_0[\mathsf{Query}]\cdot 1 + \Pr_0[\overline{\mathsf{Query}}]\cdot \frac{1}{2}$, we have

$$\mathsf{Adv}^{\mathsf{GKE},\mathcal{O}^H}_{\Pi}(t,\mathsf{q}) \le \Pr_0[\mathsf{Query}]. \qquad (1)$$

In the remainder of the proof, we focus on bounding $\Pr_0[\mathsf{Query}]$.

Experiment 1. In this experiment, X_0 is replaced by $X_0' = -\sum_{i=1}^{N-1} X_i + e_0'$. The remainder of the experiment is exactly the same as *Experiment 0*. The corresponding distribution of $(\mathsf{T}, \mathsf{sk})$ is as follows, denoted Dist_1:

$$\mathsf{Dist}_1 := \left\{ \begin{array}{l} a \leftarrow \mathsf{U}(R_q);\ \forall i: s_i, e_i \leftarrow \chi_{\sigma_1}; \\ \forall i: z_i = as_i + e_i; \\ e_1', \ldots, e_{N-1}' \leftarrow \chi_{\sigma_1}; e_0' \leftarrow \chi_{\sigma_2} \\ X_0' = -\sum_{i=1}^{N-1} X_i + e_0';\ i > 0: X_i = (z_{i+1} - z_{i-1})s_i + e_i' \\ e_{N-1}'' \leftarrow \chi_{\sigma_1}; \\ b_{N-1} = z_{N-2}Ns_{N-1} + e_{N-1}'' + X_{N-1} \cdot (N-1) + \\ \quad X_0 \cdot (N-2) + \cdots + X_{N-3}; \\ (m_{N-1}^{\mathsf{rec}}, k_{N-1}) = \mathsf{recMsg}(b_{N-1}); \mathsf{sk} = \mathsf{H}(k_{N-1}); \\ \mathsf{T} = (z_0, \ldots, z_{N-1}, X_0, \ldots, X_{N-1}, m_{N-1}^{\mathsf{rec}}). \end{array} : (\mathsf{T}, \mathsf{sk}) \right\}.$$

Claim. If $2N\sqrt{n}\lambda^{3/2}\sigma_1^2 + (N-1)\sigma_1 \leq \beta_{\mathsf{R\acute{e}nyi}}$, we have

$$\Pr_0[\mathsf{Query}] \leq \sqrt{\Pr_1[\mathsf{Query}] \cdot \frac{\exp(2\pi n(\beta_{\mathsf{R\acute{e}nyi}}/\sigma_2)^2)}{1 - 2^{-\lambda+1}}} + 2^{-\lambda+1}. \qquad (2)$$

Proof. Let Error be the difference between the distribution of X_0 in *Experiment 0* and the distribution of X_0' in *Experiment 1*, denoted $\mathsf{Error} = X_0 - X_0' = \sum_{i=0}^{N-1}(s_i e_{i+1} + s_i e_{i-1}) + \sum_{i=1}^{N-1} e_i'$. It is straightforward to verify that the distribution of X_0 in *Experiment 0* is

$$-\left(as_1 s_0 - as_{N-1}s_0 - \sum_{i=0}^{N-1}(e_{i+1}s_i + e_{i-1}s_i) - \sum_{i=1}^{N-1} e_i' \right) + \mathsf{Error} + \chi_{\sigma_2},$$

and the distribution of X_0' in *Experiment 1* is

$$-\left(as_1 s_0 - as_{N-1}s_0 - \sum_{i=0}^{N-1}(e_{i+1}s_i + e_{i-1}s_i) - \sum_{i=1}^{N-1} e_i' \right) + \chi_{\sigma_2}.$$

For simplicity, we let brick denote $as_1 s_0 - as_{N-1}s_0 - \sum_{i=0}^{N-1}(e_{i+1}s_i + e_{i-1}s_i) - \sum_{i=1}^{N-1} e_i'$.

We begin by showing that the absolute value of each coefficient of Error is bounded by $\beta_{\mathsf{R\acute{e}nyi}}$ with all but negligible probability. Then by adding a "bigger" error $e_0' \leftarrow \chi_{\sigma_2}$, the small difference between distributions $\mathsf{brick} + \mathsf{Error} + \chi_{\sigma_2}$ (corresponding to Experiment 0) and $\mathsf{brick} + \chi_{\sigma_2}$ (corresponding to Experiment 1) can be "washed" away by applying Theorem 2.1.

For all coefficient indices j, note that $|\mathsf{Error}_j| = |(\sum_{i=0}^{N-1}(s_i e_{i+1} + s_i e_{i-1}) + \sum_{i=1}^{N-1} e'_i)_j|$. Let bound_λ denote the event that for all i and all coordinate indices j, $|(s_i)_j| \le c\sigma_1$, $|(e_i)_j| \le c\sigma_1$, $|(e'_{i\neq 0})_j| \le c\sigma_1$, $|(e'_{N-1})_j| \le c\sigma_1$, and $|(e'_0)_j| \le c\sigma_2$, where $c = \sqrt{\frac{2\lambda}{\pi \log e}}$. We denote by $\mathsf{bound}_{\mathsf{Err}}$ the event that $\forall j, |\mathsf{Error}_j| \le \beta_{\text{Rényi}}$. By replacing ρ with λ in Lemma A.1 and Lemma A.2, we have $\Pr[\mathsf{bound}_\lambda] \ge 1-2^{-\lambda}$ and $\Pr[|(s_i e_j)_v| \ge \sqrt{n}\lambda^{3/2}\sigma_1^2 \mid \mathsf{bound}_\lambda] \le 2^{-2\lambda+1}$. By Union Bound, we have $\Pr[\forall j, |\mathsf{Error}_j| \le 2N\sqrt{n}\lambda^{3/2}\sigma_1^2 + (N-1)\sigma_1 \mid \mathsf{bound}_\lambda] \ge 1 - 2N \cdot 2n2^{-2\lambda}$. Under the assumption that $4Nn \le 2^\lambda$ and using similar argument as in Equations (11) and (12) of Lemma A.2, we conclude that

$$\Pr[\mathsf{bound}_{\mathsf{Err}}] \ge 1 - 2^{-\lambda+1}. \tag{3}$$

For a fixed $\mathsf{Error} \in R_q$, we note that $\mathsf{Error} + x_{\sigma_2}$, x_{σ_2} are n-dimensional distributions.

Since $\sigma_2 = \Omega(\beta_{\text{Rényi}}\sqrt{n/\log\lambda})$, assuming that for all j, $|\mathsf{Error}_j| < \beta_{\text{Rényi}}$, by Theorem 2.1, we have

$$RD_2(\mathsf{Error} + x_{\sigma_2} \| x_{\sigma_2}) \le \exp(2\pi n(\beta_{\text{Rényi}}/\sigma_2)^2) = \mathrm{poly}(\lambda). \tag{4}$$

In addition, the remaining part brick of Dist_1 is identical to Real. Therefore we may view Real in *Experiment 0* as a function of a random variable sampled from $\mathsf{Error} + x_{\sigma_2}$ and take Dist_1 in *Experiment 1* as a function of a random variable sampled from x_{σ_2}.

Recall that Query is the event that k_{N-1} is contained in the set of random oracle queries issued by adversary A. Note that Error_j is defined in both *Experiment 0* and *Experiment 1*. We denote by $\Pr_0[\mathsf{bound}_{\mathsf{Err}}]$ (resp. $\Pr_1[\mathsf{bound}_{\mathsf{Err}}]$) the probability that event $\mathsf{bound}_{\mathsf{Err}}$ occurs in *Experiment 0* (resp. *Experiment 1*) and define $\Pr_0[\overline{\mathsf{bound}_{\mathsf{Err}}}]$, $\Pr_1[\overline{\mathsf{bound}_{\mathsf{Err}}}]$ analogously. Let Real' (resp. Dist'_1) denote the random variable Real (resp. Dist_1), conditioned on the event $\mathsf{bound}_{\mathsf{Err}}$. Therefore, we have

$$\begin{aligned}
\Pr_0[\mathsf{Query}] &= \Pr_0[\mathsf{Query}|\mathsf{bound}_{\mathsf{Err}}] \cdot \Pr_0[\mathsf{bound}_{\mathsf{Err}}] + \Pr_0[\mathsf{Query}|\overline{\mathsf{bound}_{\mathsf{Err}}}] \cdot \Pr_0[\overline{\mathsf{bound}_{\mathsf{Err}}}] \\
&\le \Pr_0[\mathsf{Query}|\mathsf{bound}_{\mathsf{Err}}] + \Pr_0[\overline{\mathsf{bound}_{\mathsf{Err}}}] \\
&\le \Pr_0[\mathsf{Query}|\mathsf{bound}_{\mathsf{Err}}] + 2^{-\lambda+1} \\
&\le \sqrt{\Pr_1[\mathsf{Query}|\mathsf{bound}_{\mathsf{Err}}] \cdot RD_2(\mathsf{Real}'\|\mathsf{Dist}'_1)} + 2^{-\lambda+1} \\
&\le \sqrt{\Pr_1[\mathsf{Query}|\mathsf{bound}_{\mathsf{Err}}] \cdot RD_2(D_1\|x_{\sigma_2})} + 2^{-\lambda+1} \\
&\le \sqrt{\Pr_1[\mathsf{Query}|\mathsf{bound}_{\mathsf{Err}}] \cdot \exp(2\pi n(\beta_{\text{Rényi}}/\sigma_2)^2)} + 2^{-\lambda+1} \\
&\le \sqrt{\frac{\Pr_1[\mathsf{Query}]}{\Pr_1[\mathsf{bound}_{\mathsf{Err}}]} \cdot \exp(2\pi n(\beta_{\text{Rényi}}/\sigma_2)^2)} + 2^{-\lambda+1} \\
&\le \sqrt{\Pr_1[\mathsf{Query}] \cdot \frac{\exp(2\pi n(\beta_{\text{Rényi}}/\sigma_2)^2)}{1-2^{-\lambda+1}}} + 2^{-\lambda+1},
\end{aligned}$$

where the second and last inequalities follow from (3), the third inequality follows from Proposition 1 and the fifth inequality follows from (4). □

In Appendix B, we show that

$$\Pr_1[\mathsf{Query}] \leq \left(N \cdot \mathsf{Adv}^{\mathsf{RLWE}}_{n,q,\chi_{\sigma_1},3}(t_1) + \mathsf{Adv}_{\mathsf{KeyRec}}(t_2) + \frac{q}{2^\lambda} \right),$$

which concludes the proof of Theorem 5.1. □

5.1 Parameter Constraints

Beyond the parameter settings recommended for instantiating Ring-LWE with security parameter λ, parameters $N, n, \sigma_1, \sigma_2, \lambda, \rho$ of the protocol above are also required to satisfy the following inequalities:

$$(N^2 + 2N) \cdot \sqrt{n}\rho^{3/2}\sigma_1^2 + (\frac{N^2}{2} + 1)\sigma_1 + (N-2)\sigma_2 \leq \beta_{\mathsf{Rec}} \quad \text{(Correctness)} \quad (5)$$

$$2N\sqrt{n}\lambda^{3/2}\sigma_1^2 + (N-1)\sigma_1 \leq \beta_{\mathsf{R\acute{e}nyi}} \quad \text{(Security)} \quad (6)$$

$$\sigma_2 = \Omega(\beta_{\mathsf{R\acute{e}nyi}}\sqrt{n}/\log\lambda) \quad \text{(Security)} \quad (7)$$

We comment that once the ring, the noise distributions, and the security parameters λ, ρ are fixed, the maximum number of parties is fixed.

6 Acknowledgments

This material is based on work performed under financial assistance award 70NANB15H328 from the U.S. Department of Commerce, National Institute of Standards and Technology. Work by Dana Dachman-Soled was additionally supported in part by NSF grants #CNS-1840893 and #CNS-1453045, and by a research partnership award from Cisco.

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A Correctness of the Group Key-Exchange Protocol

Theorem 4.1. *If the parameters in the group key exchange protocol Π satisfy the constraints* $(N^2 + 2N) \cdot \sqrt{n}\rho^{3/2}\sigma_1^2 + (\frac{N^2}{2} + 1)\sigma_1 + (N-2)\sigma_2 \leq \beta_{\mathsf{Rec}}$, *then each player derives the same key with probability at least* $1 - 2 \cdot 2^{-\rho}$.

Proof. We begin by introducing the following lemmas to analyze probabilities that each coordinate of $s_i, e_i, e'_i, e'_{N-1}, e'_0$ are "short" for all i, and conditioned on the first event, $s_i e_i$ is "short".

Lemma A.1. *Given $s_i, e_i, e'_i, e'_{N-1}, e'_0$ for all i as defined above, let* bound_ρ *denote the event that for all i and all coordinate indices j, $|(s_i)_j| \le c\sigma_1$, $|(e_i)_j| \le c\sigma_1$, $|(e'_{i\neq 0})_j| \le c\sigma_1$, $|(e'_{N-1})_j| \le c\sigma_1$, and $|(e'_0)_j| \le c\sigma_2$, where $c = \sqrt{\frac{2\rho}{\pi \log e}}$, we have* $\Pr[\mathsf{bound}_\rho] \ge 1 - 2^{-\rho}$.

Proof. Using the fact that $\mathrm{erfc}(x) = \frac{2}{\sqrt{\pi}} \int_x^\infty e^{-t^2} dt \le e^{-x^2}$, we obtain

$$\Pr[|v| \ge c\sigma + 1; v \leftarrow D_{\mathbb{Z}_q,\sigma}] \le 2 \sum_{x=\lfloor c\sigma+1 \rfloor}^{\infty} D_{\mathbb{Z}_q,\sigma}(x) \le \frac{2}{\sigma} \int_{c\sigma}^{\infty} e^{-\frac{\pi x^2}{\sigma^2}} dx$$
$$= \frac{2}{\sqrt{\pi}} \int_{\sqrt{\pi}c}^{\infty} e^{-t^2} dt \le e^{-c^2\pi}.$$

Note that there are $3nN$ coordinates sampled from distribution $D_{\mathbb{Z}_q,\sigma_1}$, and n coordinates sampled from distribution $D_{\mathbb{Z}_q,\sigma_2}$ in total. Assume $3nN + n \le e^{c^2\pi/2}$, since all the coordinates are sampled independently, we bound $\Pr[\mathsf{bound}_\rho]$ as follow:

$$\Pr[\mathsf{bound}_\rho] = \left(1 - \Pr[|v| \ge c\sigma_1 + 1; v \leftarrow D_{\mathbb{Z}_q,\sigma_1}]\right)^{3nN}$$
$$\cdot \left(1 - \Pr[|e'_0| \ge c\sigma_2 + 1; e'_0 \leftarrow D_{\mathbb{Z}_q,\sigma_2}]\right)^{n}$$
$$\ge 1 - (3nN + n)e^{-c^2\pi} \ge 1 - e^{-c^2\pi/2} \ge 1 - 2^{-\rho}.$$

The last inequality follows as $c = \sqrt{\frac{2\rho}{\pi \log e}}$. □

Lemma A.2. *Given $s_i, e_i, e'_i, e'_{N-1}, e'_0$ for all i as defined above, and* bound_ρ *as defined in Lemma A.1, let* $\mathsf{products}_{s_i,e_j}$ *denote the event that, for all coefficient indices v, $|(s_i e_j)_v| \le \sqrt{n}\rho^{3/2}\sigma_1^2$. we have*

$$\Pr[\mathsf{products}_{s_i,e_j} \mid \mathsf{bound}_\rho] \ge 1 - 2n \cdot 2^{-2\rho}.$$

Proof. For $t \in \{0, \ldots, n-1\}$, Let $(s_i)_t$ denote the t^{th} coefficient of $s_i \in R_q$, namely, $s_i = \sum_{t=0}^{n-1} (s_i)_t X^t$. $(e_j)_t$ is defined analogously. Since we have $X^n + 1$ as modulo of R, it is easy to see that $(s_i e_j)_v = c_v X^v$, where $c_v = \sum_{u=0}^{n-1} (s_i)_u (e_j)^*_{v-u}$, and $(e_j)^*_{v-u} = (e_j)_{v-u}$ if $v - u \ge 0$, $(e_j)^*_{v-u} = -(e_j)_{v-u+n}$, otherwise. Thus, conditioned on $|(s_i)_t| \le c\sigma_1$ and $|(e_j)_t| \le c\sigma_1$ (for all i, j, t) where $c = \sqrt{\frac{2\rho}{\pi \log e}}$, by Hoeffding's Inequality [20], we derive

$$\Pr[|(s_i e_j)_v| \ge \delta \mid \mathsf{bound}_\rho] = \Pr\left[\left|\sum_{u=0}^{n-1} (s_i)_u (e_j)^*_{v-u}\right| \ge \delta\right] \le 2\exp\left(\frac{-2\delta^2}{n(2c^2\sigma_1^2)^2}\right),$$

as each product $(s_i)_u (e_j)^*_{v-u}$ in the sum is an independent random variable with mean 0 in the range $[-c^2\sigma_1^2, c^2\sigma_1^2]$. By setting $\delta = \sqrt{n}\rho^{3/2}\sigma_1^2$, we obtain

$$\Pr[|(s_i e_j)_v| \ge \sqrt{n}\rho^{3/2}\sigma_1^2 \mid \mathsf{bound}_\rho] \le 2^{-2\rho+1}. \quad (8)$$

Finally, by Union Bound,

$$\Pr[\mathsf{products}_{s_i,e_j}|\mathsf{bound}_\rho] = \Pr[\forall v : |(s_i e_j)_v| \le \sqrt{n}\rho^{3/2}\sigma_1^2] \ge 1 - 2n \cdot 2^{-2\rho}. \tag{9}$$

□

Now we begin analyzing the chance that not all parties agree on the same final key. The correctness of KeyRec guarantees that this group key exchange protocol has agreed session key among all parties $\forall i, k_i = k_{N-1}$, if $\forall j$, the j^{th} coefficient of $|b_{N-1} - b_i| \le \beta_{\mathsf{Rec}}$.

For better illustration, we first write $X_0, \ldots, X_{N-1}$ in form of linear system as follows. $\boldsymbol{X} = [X_0 \;\; X_1 \;\; X_2 \;\; \cdots \;\; X_{N-1}]^T$

$$= \underbrace{\begin{pmatrix} 1 & 0 & 0 & 0 & \ldots & 0 & -1 \\ -1 & 1 & 0 & 0 & \ldots & 0 & 0 \\ 0 & -1 & 1 & 0 & \ldots & 0 & 0 \\ 0 & 0 & -1 & 1 & \ldots & 0 & 0 \\ \vdots & & & & \ddots & & \vdots \\ 0 & 0 & 0 & 0 & \ldots & -1 & 1 \end{pmatrix}}_{\boldsymbol{M}} \underbrace{\begin{pmatrix} as_0s_1 \\ as_1s_2 \\ as_2s_3 \\ as_3s_4 \\ \vdots \\ as_{N-2}s_{N-1} \\ as_{N-1}s_0 \end{pmatrix}}_{\boldsymbol{S}} + \underbrace{\begin{pmatrix} s_0e_1 - s_0e_{N-1} + e'_0 \\ s_1e_2 - s_1e_0 + e'_1 \\ s_2e_3 - s_2e_1 + e'_2 \\ s_3e_4 - s_3e_2 + e'_3 \\ \vdots \\ s_{N-2}e_{N-3} - s_{N-2}e_{N-3} + e'_{N-2} \\ s_{N-1}e_0 - s_{N-1}e_{N-2} + e'_{N-1} \end{pmatrix}}_{\boldsymbol{E}} . \tag{10}$$

We denote the matrices above by $\boldsymbol{M}$, $\boldsymbol{S}$, $\boldsymbol{E}$ from left to right and have the linear system as $\boldsymbol{X} = \boldsymbol{MS} + \boldsymbol{E}$. By setting $\boldsymbol{B}_i = [i-1 \; i-2 \; \cdots \; 0 \; N-1 \; N-2 \; \cdots \; i]$ as a N-dimensional vector, we can then write b_i as $\boldsymbol{B}_i \cdot \boldsymbol{X} + N(as_is_{i-1} + s_ie_{i-1}) = \boldsymbol{B}_i\boldsymbol{MS} + \boldsymbol{B}_i\boldsymbol{E} + N(as_is_{i-1} + s_ie_{i-1})$, for $i \ne N-1$ and write b_{N-1} as $\boldsymbol{B}_{N-1}\boldsymbol{MS} + \boldsymbol{B}_{N-1}\boldsymbol{E} + N(as_{N-1}s_{N-2} + s_{N-1}e_{N-2}) + e''_{N-1}$. It is straightforward to see that, entries of $\boldsymbol{MS}$ and Nas_is_{i-1} are eliminated through the process of computing $b_{N-1} - b_i$. Thus we get

$$\begin{aligned} b_{N-1} - b_i &= (\boldsymbol{B}_{N-1} - \boldsymbol{B}_i)\boldsymbol{E} + N(s_{N-1}e_{N-2} - s_ie_{i-1}) + e''_{N-1} \\ &= (N-i-1) \cdot \left(\sum_{\substack{j \in \mathbb{Z}\cap[0,i-1] \\ \text{and } j=N-1}} \left(s_je_{j+1} - s_je_{j-1} + e'_j \right) \right) + e''_{N-1} \\ &\quad + (-i-1) \left(\sum_{j=i}^{N-2} s_je_{j+1} - s_je_{j-1} + e'_j \right) + N(s_{N-1}e_{N-2} - s_ie_{i-1}) \end{aligned}$$

Observe that for an arbitrary $i \in [N]$, there are at most $(N^2 + 2N)$ terms in form of s_ue_v, at most $N^2/2$ terms in form of e'_w where $e'_w \leftarrow \chi_{\sigma_1}$, at most $N-2$ terms of e'_0, where $e'_0 \leftarrow \chi_{\sigma_2}$, and one term in form of e''_{N-1} in any coordinate of the sum above. Let $\mathsf{product}_{\mathsf{ALL}}$ denote the event that for all the terms in form of s_ue_v observed above, each coefficient of such term is bounded by $\sqrt{n}\rho^{3/2}\sigma_1^2$.

By Union Bound and by assuming $2n(N^2+2N) \le 2^{\rho}$, it is straightforward to see $\Pr[\mathsf{product}_{\mathsf{ALL}} | \mathsf{bound}_\rho] \le (N^2+2N)\cdot 2n 2^{-2\rho} \le 2^{-\rho}$.

Let bad be the event that not all parties agree on the same final key. Given the constraint $(N^2+2N)\cdot \sqrt{n}\rho^{3/2}\sigma_1^2 + (\frac{N}{2}+1)\sigma_1 + (N-2)\sigma_2 \le \beta_{\mathsf{Rec}}$ satisfied, we have

$$\begin{aligned}\Pr[\mathsf{bad}] &= \Pr[\mathsf{bad}|\mathsf{bound}_\rho]\cdot\Pr[\mathsf{bound}_\rho] + \Pr[\mathsf{bad}|\overline{\mathsf{bound}_\rho}]\cdot\Pr[\overline{\mathsf{bound}_\rho}] && (11)\\ &\le \Pr[\mathsf{product}_{\mathsf{ALL}}]\cdot 1 + 1\cdot\Pr[\overline{\mathsf{bound}_\rho}] \le 2\cdot 2^{-\rho}, && (12)\end{aligned}$$

which completes the proof.

□

B Concluding the Proof of Theorem 5.1

Theorem 5.1 (Restated). *If the parameters in group key exchange protocol Π satisfy the constraints that $2\sqrt{N}\, n\lambda^{3/2}\sigma_1^2 + (N-1)\sigma_1 \le \beta_{R\acute{e}nyi}$, $\sigma_2 = \Omega(\beta_{R\acute{e}nyi}\sqrt{n/\log\lambda})$, and H is modeled as a classical random oracle, then for any algorithm running in time t, making at most q queries to the random oracle, the maximum advantage of A in breaking GKE security is as follows:*

$$\mathsf{Adv}_{\Pi}^{\mathsf{GKE},\mathsf{O}^H}(t,\mathsf{q}) \le 2^{-\lambda+1} + \sqrt{\left(N\cdot\mathsf{Adv}^{\mathsf{RLWE}}_{n,q,\chi_{\sigma_1},3}(t_1) + \mathsf{Adv}_{\mathsf{KeyRec}}(t_2) + \frac{\mathsf{q}}{2^\lambda}\right)\cdot\frac{\exp\left(2\pi n(\beta_{R\acute{e}nyi}/\sigma_2)^2\right)}{1-2^{-\lambda+1}}},$$

where t_1 and t_2 equal to $t + \mathsf{O}(N)\cdot t_{\mathsf{ring}}$ and t_{ring} is the time to perform operations in R_q.

Proof. (Continued) Recall that *Experiment 0* is the real world experiment. We have that $\mathsf{Adv}_{\Pi}^{\mathsf{GKE},\mathsf{O}^H}(t,\mathsf{q}) \le \Pr_0[\mathsf{Query}]$ (see Equation 1), where Query is the event that k_{N-1} is among the adversary A's random oracle queries and $\Pr_i[\mathsf{Query}]$ is the probability that event Query happens in *Experiment i*.

In *Experiment 1*, we switched from X_0 as sampled in the real world to $X_0' = -\sum_{i=1}^{N-1} X_i + e_0'$ and showed (see Equation 2) that

$$\Pr_0[\mathsf{Query}] \le \sqrt{\Pr_1[\mathsf{Query}]\cdot\frac{\exp(2\pi n(\beta_{R\acute{e}nyi}/\sigma_2)^2)}{1-2^{-\lambda+1}}} + 2^{-\lambda+1}.$$

Therefore, to prove the theorem, it remains to show that

$$\Pr_1[\mathsf{Query}] \le \left(N\cdot\mathsf{Adv}^{\mathsf{RLWE}}_{n,q,\chi_{\sigma_1},3}(t_1) + \mathsf{Adv}_{\mathsf{KeyRec}}(t_2) + \frac{\mathsf{q}}{2^\lambda}\right).$$

We do so by considering a sequence of experiments as follows:

Experiment 2. This experiment proceeds exactly the same as *Experiment 1*, except that z_0 is generated uniformly at random, instead of being generated as an Ring-LWE instance. The corresponding distribution is as follows, denoted Dist_2:

$$\mathsf{Dist}_2 := \left\{ \begin{array}{l} a \leftarrow \mathsf{U}(R_q);\ \forall i \geq 1 : s_i, e_i \leftarrow \chi_{\sigma_1}; \\ z_0 \leftarrow \mathsf{U}(R_q), \forall i \geq 1 : z_i = as_i + e_i; \\ e'_1, \ldots, e'_{N-1} \leftarrow \chi_{\sigma_1}; e'_0 \leftarrow \chi_{\sigma_2} \\ X_0 = -\sum_{i=1}^{N-1} X_i + e'_0, \forall i \geq 1 : X_i = (z_{i+1} - z_{i-1})s_i + e'_i \\ e''_{N-1} \leftarrow \chi_{\sigma_1}; \\ b_{N-1} = z_{N-2}Ns_{N-1} + e''_{N-1} + X_{N-1} \cdot (N-1) + \\ \quad X_0 \cdot (N-2) + \cdots + X_{N-3}; \\ (m^{\mathsf{rec}}_{N-1}, k_{N-1}) = \mathsf{recMsg}(b_{N-1}); \mathsf{sk} = \mathsf{H}(k_{N-1}); \\ \mathsf{T} = (z_0, \ldots, z_{N-1}, X_0, \ldots, X_{N-1}, m^{\mathsf{rec}}_{N-1}). \end{array} : (\mathsf{T}, \mathsf{sk}) \right\}.$$

Bounding the difference of $|\Pr_2[\mathsf{Query}] - \Pr_1[\mathsf{Query}]|$:

Given algorithm A running in time t attacking Π, let B be an algorithm running in time t_1 that takes as input (a, z_0), generates $(\mathsf{T}, \mathsf{sk})$ based on distribution Dist'_1 which is identical to Dist_1 except for (a, z_0) given as input, runs A as subroutine and outputs whatever A outputs. It is straightforward to see that if (a, z_0) is sampled from the Ring-LWE distribution $A_{n,q,\chi_{\sigma_1}}$, then Dist'_1 is identical to Dist_1, and if (a, z_0) is sampled from $\mathsf{U}(R_q^2)$, $\mathsf{Dist}'_1$is identical to Dist_2. Note that t_1 is equal to t plus a minor overhead for the simulation of the security experiment for A.

Therefore we conclude that the difference of algorithm A's success probability in *Experiment 1* and *Experiment 2* is bounded by probability that B running in time t_1 distinguishes $A_{n,q,\chi_{\sigma_1}}$ from $\mathsf{U}(R_q)$ given one sample. Since $\mathsf{Adv}^{\mathsf{RLWE}}_{n,q,\chi_{\sigma_1},3}(t_1) \geq \mathsf{Adv}^{\mathsf{RLWE}}_{n,q,\chi_{\sigma_1},2}(t_1) \geq \mathsf{Adv}^{\mathsf{RLWE}}_{n,q,\chi_{\sigma_1},1}(t_1)$, for simplicity, we have

$$|\Pr_2[\mathsf{Query}] - \Pr_1[\mathsf{Query}]| \leq \mathsf{Adv}^{\mathsf{RLWE}}_{n,q,\chi_{\sigma_1},3}(t_1). \tag{13}$$

Recall that in the previous experiment, we switched z_0 to be uniformly distributed in R_q. In next two experiments, we switch z_1, X_1 to be elements uniformly distributed in R_q.

Experiment 3. the experiment proceeds exactly the same as *Experiment 2*, except for setting $z_0 = z_2 - r_1$, $X_1 = r_1 s_1 + e'_1$, where r_1 is sampled from $\mathsf{U}(R_q)$. The corresponding distribution is as follows, denoted as Dist_3.

Bounding the difference of $|\Pr_3[\mathsf{Query}] - \Pr_2[\mathsf{Query}]|$: Since r_1 is sampled uniformly, $z_2 - r_1$ is also a uniformly distributed random value, then we claim that *Experiment 3* is identical to *Experiment 3* up to variable substitution, namely

$$\Pr_3[\mathsf{Query}] = \Pr_2[\mathsf{Query}]. \tag{14}$$

$$\mathsf{Dist}_3 := \left\{ \begin{array}{l} a \leftarrow \mathsf{U}(R_q), r_1 \leftarrow \mathsf{U}(R_q);\ \forall i \geq 1 : s_i, e_i \leftarrow \chi_{\sigma_1}; \\ z_0 = z_2 - r_1;\ \forall i \geq 1 : z_i = as_i + e_i; \\ \forall i \geq 1 : e'_i \leftarrow \chi_{\sigma_1};\ e'_0 \leftarrow \chi_{\sigma_2}; \\ X_0 = -\sum_{i=1}^{N-1} X_i + e'_0; X_1 = r_1 s_1 + e'_1; \\ \forall i \geq 2 : X_i = (z_{i+1} - z_{i-1})s_i + e'_i; e'_{N-1} \leftarrow \chi_{\sigma_1}; \\ b_{N-1} = z_{N-2}Ns_{N-1} + e'_{N-1} + X_{N-1} \cdot (N-1) + \\ \quad X_0 \cdot (N-2) + \cdots + X_{N-3}; \\ (m^{\mathsf{rec}}_{N-1}, k_{N-1}) = \mathsf{recMsg}(b_{N-1}); \mathsf{sk} = \mathsf{H}(k_{N-1}); \\ \mathsf{T} = (z_0, \ldots, z_{N-1}, X_0, \ldots, X_{N-1}, m^{\mathsf{rec}}_{N-1}). \end{array} : (\mathsf{T}, \mathsf{sk}) \right\}.$$

Experiment 4. This experiment proceeds exactly the same as *Experiment 3*, except that z_1, X_1 are uniformly distributed in R_q. The corresponding distribution is as follows, denoted as Dist_4.

$$\mathsf{Dist}_4 := \left\{ \begin{array}{l} a, r_1 \leftarrow \mathsf{U}(R_q);\ \forall i \geq 2 : s_i, e_i \leftarrow \chi_{\sigma_1}; \\ z_0 = z_2 - r_1, z_1 \leftarrow \mathsf{U}(R_q);\ \forall i \geq 2 : z_i = as_i + e_i; \\ e'_2, \ldots, e'_{N-1} \leftarrow \chi_{\sigma_1}; e'_0 \leftarrow \chi_{\sigma_2}; \\ X'_0 = -\sum_{i=1}^{N-1} X_i + e'_0, X_1 \leftarrow \mathsf{U}(R_q); \\ \forall i \geq 2 : X_i = (z_{i+1} - z_{i-1})s_i + e'_i, \\ e'_{N-1} \leftarrow \chi_{\sigma_1}; \\ b_{N-1} = z_{N-2}Ns_{N-1} + e'_{N-1} + X_{N-1} \cdot (N-1) + \\ \quad X_0 \cdot (N-2) + \cdots + X_{N-3}; \\ (m^{\mathsf{rec}}_{N-1}, k_{N-1}) = \mathsf{recMsg}(b_{N-1}); \mathsf{sk} = \mathsf{H}(k_{N-1}); \\ \mathsf{T} = (z_0, \ldots, z_{N-1}, X_0, \ldots, X_{N-1}, m^{\mathsf{rec}}_{N-1}) \end{array} : (\mathsf{T}, \mathsf{sk}) \right\}.$$

Bounding the difference of $|\Pr_4[\mathsf{Query}] - \Pr_3[\mathsf{Query}]|$:

Given an algorithm A running in time t attacking Π, let B be an algorithm running in time t_1 that takes as input $(a, z_1), (r_1, X_1)$, generates $(\mathsf{T}, \mathsf{sk})$ based on distribution Dist'_3 which is identical to Dist_3 except for $(a, z_1), (r_1, X_1)$ given as input. B runs A as a subroutine and outputs whatever A outputs. Note that t_1 is equal to t plus a minor overhead for the simulation of the security experiment for A.

It is clear to see that if (a, z_1) and (r_1, X_1) are sampled from the Ring-LWE distribution $A_{n,q,\chi_{\sigma_1}}$, then Dist'_3 is identical to Dist_3. If (a, z_1) and (r_1, X_1) are sampled from $\mathsf{U}(R_q^2)$, Dist'_3 is identical to Dist_4.

Therefore we conclude that the difference of algorithm A successful probability in winning *Experiment 4* and *Experiment 3* is bounded by the advantage of adversary B running in time t_1 in distinguishing $A_{n,q,\chi_{\sigma_1}}$ from $\mathsf{U}(R_q^2)$ given two samples. Thus,

$$|\Pr_4[\mathsf{Query}] - \Pr_3[\mathsf{Query}]| \leq \mathsf{Adv}^{\mathsf{RLWE}}_{n,q,\chi_{\sigma_1},3}(t_1). \tag{15}$$

Experiment 5. This experiment proceeds exactly the same as *Experiment 4*, except that z_0 is sampled directly from $\mathsf{U}(R_q)$. We leave the formal definition of Dist_5 implicit for simplicity.

Bounding the difference of $|\Pr_5[\mathsf{Query}] - \Pr_4[\mathsf{Query}]|$: It is easy to see that the corresponding distribution Dist_5 is identical to Dist_4 by substituting variable z_0 for $z_2 - r_1$. Thus,

$$\Pr_5[\mathsf{Query}] = \Pr_4[\mathsf{Query}]. \tag{16}$$

In the case that $N \geq 3$, we present the following sequence of experiments from *Experiment 6* to *Experiment* $3N-4$. For $i = 2, 3, \ldots, N-2$, we define three experiments *Experiment* $3i$, *Experiment* $3i+1$, *Experiment* $3i+2$. It is ensured that in the experiments prior to *Experiment* $3i$, we already switched z_j, X_j for all $0 \leq j \leq i-1$. In *Experiment* $3i$, *Experiment* $3i+1$ and *Experiment* $3i+2$, we replace z_i and X_i by random elements uniformly distributed in R_q. *Experiment* $3i$, *Experiment* $3i+1$, *Experiment* $3i+2$ are formally defined as follows:

Experiment $3i$. The experiment proceeds exactly the same as *Experiment* $3i-1$, except for setting $z_{i-1} = z_{i+1} - r_i$, $X_i = r_i s_i + e'_i$, where r_1 is sampled from $\mathsf{U}(R_q)$. The corresponding distribution is as follows, denoted Dist_{3i}

$$\mathsf{Dist}_{3i} := \left\{ \begin{array}{l} a, r_i \leftarrow \mathsf{U}(R_q);\ \ \forall j \geq i : s_j, e_j \leftarrow \chi_{\sigma_1}; \\ z_0, \ldots, z_{i-2} \leftarrow \mathsf{U}(R_q), z_{i-1} = z_{i+1} - r_i; \\ \forall j \geq i : z_j = a s_j + e_j; \\ e'_i, \ldots, e'_{N-1} \leftarrow \chi_{\sigma_1}, e'_0 \leftarrow \chi_{\sigma_2}; \\ X_0 = -\sum_{i=1}^{N-1} X_i + e'_0, X_1, \ldots, X_{i-1} \leftarrow \mathsf{U}(R_q); \\ X_i = r_i s_i + e'_i;\ \ \forall j \geq i : X_{j+1} = (z_{j+2} - z_j) s_{j+1} + e'_{j+1} \\ e''_{N-1} \leftarrow \chi_{\sigma_1}; \\ b_{N-1} = z_{N-2} N s_{N-1} + e''_{N-1} + X_{N-1} \cdot (N-1) + \\ \qquad X_0 \cdot (N-2) + \cdots + X_{N-3}; \\ (m^{\mathsf{rec}}_{N-1}, k_{N-1}) = \mathsf{recMsg}(b_{N-1}); \mathsf{sk} = \mathsf{H}(k_{N-1}); \\ \mathsf{T} = (z_0, \ldots, z_{N-1}, X_0, \ldots, X_{N-1}, m^{\mathsf{rec}}_{N-1}) \end{array} : (\mathsf{T}, \mathsf{sk}) \right\}.$$

Experiment $3i+1$. This experiment proceeds exactly the same as *Experiment 3i*, except that z_i, X_i are uniformly distributed in R_q. The corresponding distribution is as follows, denoted Dist_{3i+1}:

$$
\mathsf{Dist}_{3i+1} := \left\{ \begin{array}{l} a, r_i \leftarrow \mathsf{U}(R_q);\ \ \forall j \geq i+1: s_j, e_j \leftarrow \chi_{\sigma_1} \\ z_0, \ldots, z_{i-2} \leftarrow \mathsf{U}(R_q), z_{i-1} = z_{i+1} - r_i, z_i \leftarrow \mathsf{U}(R_q), \\ \forall j \geq i+1 : z_j = a s_j + e_j; \\ e'_1, \ldots, e'_{N-1} \leftarrow \chi_{\sigma_1}; e'_0 \leftarrow \chi_{\sigma_2} \\ X_0 = -\sum_{i=1}^{N-1} X_i + e'_0, X_1, \ldots, X_i \leftarrow \mathsf{U}(R_q), \\ \forall j \geq i+1, X_j = (z_{j+1} - z_{j-1}) s_j + e'_j; \\ e''_{N-1} \leftarrow \chi_{\sigma_1}; \\ b_{N-1} = z_{N-2} N s_{N-1} + e''_{N-1} + X_{N-1} \cdot (N-1) + \\ \qquad X_0 \cdot (N-2) + \cdots + X_{N-3}; \\ (m^{\mathsf{rec}}_{N-1}, k_{N-1}) = \mathsf{recMsg}(b_{N-1}); \mathsf{sk} = \mathsf{H}(k_{N-1}); \\ \mathsf{T} = (z_0, \ldots, z_{N-1}, X_0, \ldots, X_{N-1}, m^{\mathsf{rec}}_{N-1}) \end{array} : (\mathsf{T}, \mathsf{sk}) \right\}.
$$

Experiment $3i+2$**.** This experiment proceeds exactly the same as *Experiment* $3i+1$, except that z_{i-1} is directly sampled from $\mathsf{U}(R_q)$. The corresponding distribution is denoted as Dist_{3i+2}. We leave the formal definition of Dist_{3i+2} implicit for simplicity.

Bounding the difference of $|\Pr_{3i}[\mathsf{Query}] - \Pr_{3i-1}[\mathsf{Query}]|$, $|\Pr_{3i+1}[\mathsf{Query}] - \Pr_{3i}[\mathsf{Query}]|$, and $|\Pr_{3i+2}[\mathsf{Query}] - \Pr_{3i+1}[\mathsf{Query}]|$ follows exactly the same logic as *bounding the differences of* $|\Pr_3[\mathsf{Query}] - \Pr_2[\mathsf{Query}]|$, $|\Pr_4[\mathsf{Query}] - \Pr_3[\mathsf{Query}]|$, and $|\Pr_5[\mathsf{Query}] - \Pr_4[\mathsf{Query}]|$, respectively. Then we have

$$\Pr_{3i}[\mathsf{Query}] = \Pr_{3i-1}[\mathsf{Query}]; \tag{17}$$

$$|\Pr_{3i+1}[\mathsf{Query}] - \Pr_{3i}[\mathsf{Query}]| \leq \mathsf{Adv}^{\mathsf{RLWE}}_{n,q,\chi_{\sigma_1},3}(t_1); \tag{18}$$

$$\Pr_{3i+2}[\mathsf{Query}] = \Pr_{3i+1}[\mathsf{Query}]; \tag{19}$$

Note that in *Experiment* $3N-4$, the last experiment of the experiment sequence above, we already switched all the z_i, X_i up to z_{N-1}, X_{N-1}. We construct the next two experiments to switch $z_{N-1}, X_{N-1}, b_{N-1}$.

Experiment $3N-3$**.** The experiment proceeds exactly the same as *Experiment* $3N-4$, except that we let $z_{N-2} = r_2$, $X_{N-1} = r_1 s_{N-1} + e'_{N-1}$, $z_0 = r_1 + r_2$, where r_1, r_2 are uniformly distributed in R_q. The corresponding distribution is as follows, denoted Dist_{3N-3}.

Bounding the difference of $|\Pr_{3N-3}[\mathsf{Query}] - \Pr_{3N-4}[\mathsf{Query}]|$:

Since r_1, r_2 is sampled uniformly, $r_1 + r_2$ is also uniformly distributed in R_q. Then we claim that *Experiment* $3N-3$ is identical to *Experiment* $3N-4$ up to variable substitution, written as

$$\Pr_{3N-3}[\mathsf{Query}] = \Pr_{3N-4}[\mathsf{Query}]; \tag{20}$$

$$\text{Dist}_{3N-3} := \left\{ \begin{array}{l} a, r_1, r_2 \leftarrow \mathsf{U}(R_q), s_{N-1}, e_{N-1} \leftarrow \chi_{\sigma_1}; z_0 = r_1 + r_2, \\ z_1, \ldots, z_{N-3} \leftarrow \mathsf{U}(R_q), z_{N-2} = r_2, \\ z_{N-1} = as_{N-1} + e_{N-1}; e'_0 \leftarrow \chi_{\sigma_2}; e'_{N-1} \leftarrow \chi_{\sigma_1}; \\ X'_0 = -\sum_{i=1}^{N-1} X_i + e'_0, X_1, \ldots, X_{N-2} \leftarrow \mathsf{U}(R_q), \\ X_{N-1} = r_1 s_{N-1} + e'_{N-1}; e''_{N-1} \leftarrow \chi_{\sigma_1}; \\ b_{N-1} = r_2 N s_{N-1} + e''_{N-1} + X_{N-1} \cdot (N-1) + \\ \quad X_0 \cdot (N-2) + \cdots + X_{N-3}; \\ (m^{rec}_{N-1}, k_{N-1}) = \text{recMsg}(b_{N-1}); \text{sk} = \mathsf{H}(k_{N-1}); \\ \mathsf{T} = (z_0, \ldots, z_{N-1}, X_0, \ldots, X_{N-1}, m^{rec}_{N-1}). \end{array} : (\mathsf{T}, \text{sk}) \right\}.$$

Experiment $3N-2$**.** This experiment proceeds exactly the same as *Experiment* $3N-3$, except that z_{N-1}, X_{N-1}, b_{N-1} are generated from $\mathsf{U}(R_q)$. The corresponding distribution is as follows, denoted Dist_{3N-2}:

$$\text{Dist}_{3N-2} := \left\{ \begin{array}{l} a \leftarrow \mathsf{U}(R_q), z_0, z_1 \ldots, z_{N-2} \leftarrow \mathsf{U}(R_q), \\ z_{N-1} \leftarrow \mathsf{U}(R_q); e'_0 \leftarrow \chi_{\sigma_2}; r_1, r_2 \leftarrow \mathsf{U}(R_q) \\ X'_0 = -\sum_{i=1}^{N-1} X_i + e'_0, X_1, \ldots, X_{N-1} \leftarrow \mathsf{U}(R_q) \\ b_{N-1} \leftarrow \mathsf{U}(R_q); \\ (m^{rec}_{N-1}, k_{N-1}) = \text{recMsg}(b_{N-1}); \text{sk} = \mathsf{H}(k_{N-1}); \\ \mathsf{T} = (z_0, \ldots, z_{N-1}, X_0, \ldots, X_{N-1}, m^{rec}_{N-1}). \end{array} : (\mathsf{T}, \text{sk}) \right\}.$$

Bounding the difference of $|\Pr_{3N-2}[\text{Query}] - \Pr_{3N-3}[\text{Query}]|$:

Let $b_{rlwe} = r_2 N s_{N-1} + e''_{N-1}$, then $b_{N-1} = b_{rlwe} + X_{N-1} \cdot (N-1) + X_0 \cdot (N-2) + \cdots + X_{N-3}$. As r_2 is sampled uniformly at random and N is invertible over R_q, $r_2 N$ is uniformly distributed in R_q.

Given an algorithm A running in time t attacking group key exchange protocol Π, let B be an algorithm that takes as input (a, z_{N-1}), (r_1, X_{N-1}), and $(r_2 N, b_{rlwe})$, generates (T, sk) based on distribution Dist'_{3N-3} which is identical to Dist_{3N-3} except for (a, z_{N-1}), (r_1, X_{N-1}), and $(r_2 N, b_{rlwe})$ given as input. B runs A as subroutine and outputs whatever A outputs. Note that running time t_1 of B equals to t plus a minor overhead for the simulation of the security experiment for A.

It is straightforward to see that if (a, z_{N-1}), (r_1, X_1), and $(r_2 N, b_{rlwe})$ are sampled from the Ring-LWE distribution $A_{n,q,\chi_{\sigma_1}}$, then Dist'_{3N-3} is identical to Dist_{3N-3}. If (a, z_{N-1}), (r_1, X_{N-1}), and $(r_2 N, b_{rlwe})$ are sampled from $\mathsf{U}(R_q^2)$, then Dist'_{3N-3} is identical to Dist_{3N-2}, since when b_{rlwe} is sampled uniformly at random, $b_{rlwe} + X_{N-1} \cdot (N-1) + X_0 \cdot (N-2) + \cdots + X_{N-3}$ is also uniformly distributed over R_q.

Therefore we conclude that the difference of algorithm A^{GKE}'s success probability in *Experiment 3N - 2* and *Experiment 3N - 3* is bounded by the advantage of adversary $\mathcal{B}$ running in time t_1 in distinguishing Ring-LWE from (R_q) given three samples. Thus, we conclude that

$$|\Pr_{3N-2}[\mathsf{Query}] - \Pr_{3N-3}[\mathsf{Query}]| \leq \mathsf{Adv}^{\mathsf{RLWE}}_{n,q,x_{o_1},3}(t_1). \tag{21}$$

Experiment $3N-1$**.** This experiment proceeds exactly the same as *Experiment* $3N-2$, except that k_{N-1} is directly sampled uniformly from $\{0,1\}^\lambda$. Note that the corresponding distribution is exactly the distribution Ideal.

$$\mathsf{Ideal} := \left\{ \begin{array}{l} a \leftarrow \mathsf{U}(R_q);\ z_0,\ldots,z_{N-1} \leftarrow \mathsf{U}(R_q);\ e_0 \leftarrow x_{o_1}; \\ X'_0 = -\sum_{i=1}^{N-1} X_i + e'_0,\ X_1,\ldots,X_{N-1} \leftarrow \mathsf{U}(R_q) \\ b_{N-1} \leftarrow \mathsf{U}(R_q);\ (m^{\mathsf{rec}}_{N-1}, k_{N-1}) = \mathsf{recMsg}(b_{N-1}) \\ k'_{N-1} \leftarrow \{0,1\}^\lambda;\ \mathsf{sk} = \mathsf{H}(k'_{N-1}); \\ \mathsf{T} = (z_0,\ldots,z_{N-1}, X'_0,\ldots,X_{N-1}, m^{\mathsf{rec}}_{N-1}); \end{array} : (\mathsf{T},\mathsf{sk}) \right\}.$$

Bounding the difference of $|\Pr_{3N-1}[\mathsf{Query}] - \Pr_{3N-2}[\mathsf{Query}]|$:

Given transcript T, and b_{N-1} which is uniformly distributed, using a straight forward reduction, we obtain advantage of adversary B running in time t_2 in distinguishing k_{N-1} computed by $\mathsf{recMsg}(b_{N-1})$ from a uniform bit string k'_{N-1} with length λ is at least $|\Pr_{3N-1}[\mathsf{Query}] - \Pr_{3N-2}[\mathsf{Query}]|$, namely,

$$|\Pr_{3N-1}[\mathsf{Query}] - \Pr_{3N-2}[\mathsf{Query}]| \leq \mathsf{Adv}_{\mathsf{KeyRec}}(t_2). \tag{22}$$

Note that t_2 equals to the running time of adversary $\mathcal{A}$ attacking the protocol Π, plus a minor overhead for simulating experiment for $\mathcal{A}$.

Finally, since adversary attacking the GKE protocol Π makes at most q queries to the random oracle, $\Pr_{3N-1}[\mathsf{Query}] = \frac{\mathsf{q}}{2^\lambda} \in \mathsf{negl}(\lambda)$. Combining Equations (13) - (22), we have

$$\Pr_1[\mathsf{Query}] \leq N \cdot \mathsf{Adv}^{\mathsf{RLWE}}_{n,q,x_{o_1},3}(t_1) + \mathsf{Adv}_{\mathsf{KeyRec}}(t_2) + \frac{\mathsf{q}}{2^\lambda}. \tag{23}$$

The theorem now follows immediately from Equations (1), (2), and (23). □