

Efficient and Tight Oblivious Transfer from PKE with Tight Multi-User Security*

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Abstract. We propose an efficient oblivious transfer in the random oracle model based on public key encryption with pseudorandom public keys. The construction is as efficient as the state of art though it has a significant advantage. It has a tight security reduction to the multi-user security of the underlying public key encryption. In previous constructions, the security reduction has a multiplicative loss that amounts in at least the amount of adversarial random oracle queries. When considering this loss for a secure parameter choice, the underlying public key encryption or elliptic curve would require a significantly higher security level which would decrease the overall efficiency.

Our OT construction can be instantiated from a wide range of assumptions such as DDH, LWE, or codes based assumptions as well as many public key encryption schemes such as the NIST PQC finalists. Since tight multi-user security is a very natural requirement which many public key encryption schemes suffice, many public key encryption schemes can be straightforwardly plugged in our construction without the need of reevaluating or adapting any parameter choices.

1 Introduction

An oblivious transfer (OT) [Rab81, EGL82] is an interactive protocol between two parties called a sender and a receiver. At the end of the protocol, the sender outputs two messages m_0, m_1 while the receiver outputs b, m_b for a choice bit b . Security requires that the sender does not learn b and the receiver does not learn m_{1-b} . OT is a fundamental building block in cryptography [Kil88], particularly in secure multi-party computation (MPC) [Yao82, Yao86, CvT95, IPS08, IKO⁺11, BL18, GS18], which allows mutually distrusting parties to securely perform joint computations on their privately held data. MPC has a plethora of applications in practice, for example, in securely training machine learning models (e.g. [MR18]), private set intersection (e.g. [KKRT16, PRTY20]) etc. In fact, a significant body of practically efficient MPC protocols do rely primarily on the primitive of OT (e.g. [NNOB12, KOS16]), which makes *efficient secure OT* an important and very natural objective.

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Within the last years, there has been significant progress in making OT more efficient. Chou and Orlandi [CO15] proposed a very efficient OT in the random oracle model [BR93, CGH98] based on the DDH assumption. It turned out, that it does not achieve UC security [GIR17, HL17], but only stand-alone security. Masny and Rindal [MR19] proposed an OT from public key encryption (PKE) with pseudorandom public keys that is as well very efficient but also UC secure and can be instantiated from a variety of assumptions such as LWE or code based assumptions. The construction makes it very easy to plug in PKE schemes such as the NIST PQC candidates [SAB⁺20, DKR⁺20, CDH⁺20, ABC⁺20] which is a significant advantage over more tailored construction of OT based on DDH [CSW20], LWE [PVW08, BD18, BDK⁺20] or McEliece [DvMN08, DN12]. McQuoid, Rosulek and Roy [MRR20, MRR21] gave a more modular analysis of this approach, extended it to PKEs with pseudorandom ciphertexts (PKE B) as well as increased the efficiency when multiple OTs are run in parallel. Masny and Watson [MW21] increased the efficiency by leveraging a PKI.

This approach works as follows. Using a specific query pattern to a random oracle, a receiver can freely chose one public key while a second public key will be completely determined by the random oracle. At the same time, a sender can reproduce the same queries and public keys and then encrypt one OT string under each of the public keys. Though he will not be able to determine which of the keys has been freely chosen by the receiver. At the same time, the receiver can only recover the string under the freely chosen public key but not the other. Unfortunately, this approach has some drawbacks, namely the receiver could repeat the query pattern to the random oracle until he finds a public keys that might be easier to break than the average public key and then try to recover both strings. Typically, a PKE is hard to break for a random public key with overwhelming probability and therefore it should not cause an issue. Nevertheless, it limits how tightly one can prove the security of the OT protocol based on a PKE scheme.

This drawback can be resolved by using a PKE that is tightly secure in the multi-user setting. Tight multi user security has received significant attention in the context of key exchange, PKE and signatures [Hås88, BBM00, HJ12, Zav12, BHJ⁺15, KMP16, CKMS16, GKP18, GJ18, PR20, LLGW20, JKRS21]. Bellare, Boldyreva and Micali [BBM00] showed that ElGamal is tightly secure even when multiple challenge ciphertexts are given to the adversary. There are numerous works that focus on tight multi-user secure PKE [Hås88, HJ12, Zav12, CKMS16, GKP18]. The tightness requirement does not put significant restrictions on known PKEs. Tight multi-user security is a very natural property that a PKE should typically have since usually the security of all users and not just of a single user needs to be considered. Non-tightness would demand an increase in the bit security level of a PKE when used across many users which would render the PKE significantly less efficient.

Unfortunately, using a tightly secure PKE in the multi user setting is not sufficient. The security analysis of [MR19, MRR20, MRR21] also involves re-programming the random oracle and guessing which query a malicious receiver

will later use during the OT protocol. This comes at the cost of a security loss which is multiplicative in the amount of adversarial random oracle queries. This issue seems to require a more in-depth analysis of this approach of constructing OT and opens the question whether a similar construction could achieve tight security. In this paper, we answer the following question:

Can we construct efficient OT that is tightly secure in the ROM from public-key encryption?

1.1 Our Contribution

We propose a new construction of OT in the random oracle model which can be proven tightly secure based on the multi-user security of the underlying PKE. This approach follows the paradigm of Masny and Rindal [MR19, MRR20, MRR21, MW21] by specifying a pattern of random oracle queries which allows a malicious receiver to choose one public key freely while a second one is determined by the random oracle.

We use a mild notion of multi-user security which is weaker than the notion proposed in previous literature such as [BBM00]. In our notion, we require that an adversary receives n user public keys and then decides for which he wants to see a challenge ciphertext. The notion of [BBM00] allows an adversary to see challenge ciphertext for all of the public keys. Nevertheless, there are many PKEs that even achieve the stronger notion of [BBM00] with a tight security proof under the DDH or LWE [Reg05] assumption. We recap the most basic PKEs and their tight reductions to DDH and LWE in Section 3. The results extend straightforwardly to code based schemes, the ring or module LWE [LPR10, BGV12, LS15] setting or elliptic curves.

For our OT, we require a second property that is the pseudorandomness of the public keys. This requirement is the same as in [MR19] with the exception that it holds tightly based on the underlying assumption even when n keys are seen. We recap this property as well in Section 3 for the PKEs of interest.

In Figure 1, we compare our result with previous works. Since the main difference of our construction to [MR19] is how the random oracle is used, the efficiency of our OT is very similar to [MR19]. On one hand, we need to compute 3 additional hash evaluations. The hash evaluations are standard evaluations mapping onto $\{0, 1\}^*$ and when using elliptic curves, not to curve points. On the other hand, we are actually, similar to [MRR20] able to reduce the communication complexity on the receivers side from $2|\mathbf{pk}|$ ([MRR20]) to $|\mathbf{pk}| + 2\lambda$. In particular when instantiating the OT with lattice or code based schemes [SAB⁺20, DKR⁺20, CDH⁺20, ABC⁺20] which have rather long keys, this is a significant reduction. Even when instantiating the OT with ElGamal encryption, we need to sample one random group element less which requires an exponentiation. In the elliptic curve setting, our construction is compatible with the performance optimizations of [MRR21] and would therefore be competitive with the currently fastest implementations of UC OT reported in [MRR21]. Further, our OT is based on a PKE with pseudorandom public keys (PKE A) which,

	UC	Loss	Model	Com(\mathcal{R})	Com(\mathcal{S})
[CO15]	χ	-	ROM	$\log \mathcal{G} $	$\log \mathcal{G} $
[MR19]	PKE A	$\mathbf{O}(q)$	ROM	$2 \mathbf{pk} $	$2 \mathbf{ct} $
[CSW20]	DDH	$\mathbf{O}(q^2)$	ROM, CRS	$2 \log \mathcal{G} $	$\log \mathcal{G} $
[MRR20]	PKE B	$\mathbf{O}(q^2)$	ROM	$ \mathbf{ct} + \lambda$	$ \mathbf{pk} $
[MRR20]*	PKE A	$\mathbf{O}(q)$	ROM	$ \mathbf{ct} + \lambda$	$ \mathbf{pk} $
[MRR21]	PKE B	$\mathbf{O}(q)$	Ideal Cipher	$ \mathbf{ct} + \lambda$	$ \mathbf{pk} $
[MRR21]*	PKE A	$\mathbf{O}(1)$	Ideal Cipher	$ \mathbf{ct} + \lambda$	$ \mathbf{pk} $
Ours	PKE A	$\mathbf{O}(1)$	ROM	$ \mathbf{pk} + 2\lambda$	$2 \mathbf{ct} $

Fig. 1: We compare our construction with previous works. The depicted loss assumes tight multi-user security of the underlying PKE. We emphasize that the listed works realize different OT functionalities and therefore the comparison between the communication should be interpreted with caution. PKE A stands for PKE with pseudorandom public keys and PKE B stands for PKE with uniform ciphertexts. q is the amount of adversarial random oracle queries (hash evaluations). [MRR20]*, [MRR21]* are slight adaptations of the original works to make them compatible with the PKE A setting.

unlike PKEs with uniform ciphertexts (PKE B), can be efficiently instantiated with post-quantum PKEs, e.g. from codes or lattices. We could also use our techniques to construct an OT from a PKE with pseudorandom ciphertexts (PKE B), though it is unclear whether the tightness would still hold and it might require stronger assumptions such as the interactive DDH assumption [MR19] or oracle assumptions [BCJ⁺19, MRR21].

As shown in Figure 1, our OT is currently the only OT among the most efficient OTs that is tightly secure in the random oracle model. The main challenge is typically security against a malicious receiver. Previous works suffer at least a loss of $O(q)$ where q is the amount of adversarial hash evaluations. For a conservative parameter choice, previous works need to start with a significantly higher security level of the PKE or elliptic curve which negatively impacts efficiency and communication complexity or alternatively, use a stronger model such as the ideal cipher model as in case of [MRR21]*.

We emphasize that Figure 1 states the loss for [MRR20]* and [MRR21]* when using a PKE with tight multi user security. Using (plain) single user secure PKE, the loss increases by a factor of q . This also holds for the loss stated for [MR19].

1.2 Technical Overview

We follow an approach by Masny and Rindal [MR19]. They construct a two round OT in which the receiver starts by sending a message r_0, r_1 from this message the sender can derive two public keys under which he encrypts the two OT strings. The public keys are $\mathbf{pk}_0 := r_1 + \mathbf{H}(r_0)$ and $\mathbf{pk}_1 := r_0 + \mathbf{H}(r_1)$. When following this approach, proving security against a malicious sender is typically easy since the random oracle can be programmed such that the simulator knows

the secret keys for both public keys which can then be used to extract the malicious sender’s string. The more challenging part is to prove security against a malicious receiver \mathcal{R}^* . Given that \mathcal{R}^* makes only two random oracle queries, r_0 and r_1 , the simulator can observe the first query, let it be r_b . Then, b is the extracted choice bit. Further, when the second query is made, the simulator could pick a public key pk^* of its choice and program the oracle H such that $H(r_{1-b}) := pk^* - r_b$ and thus $pk_{1-b} = pk^*$. If \mathcal{R}^* learns information about the OT string s_{1-b} , he would then break the security of the PKE.

Unfortunately, when the malicious receiver makes many queries, it is not clear how to program $H(r_{1-b})$ since any of the q previous queries $\tilde{r}_1, \dots, \tilde{r}_q$ could be the r_b query. This would lead to the potential public keys $pk_{1-b,1} := \tilde{r}_1 + H(r_{1-b}), \dots, pk_{1-b,q} := \tilde{r}_q + H(r_{1-b})$. We could guess $j \in [q]$ such that $r_b = \tilde{r}_j$ but this would cause a loss of q .

Before explaining our construction, we first take an intermediate step. The MR OT has similarities with a sequential OR proof [RST01, AOS02]. Instead we could follow the parallel OR proof paradigm [CDS94]. The public keys would be then derived from a message r, c_0, c_1 and defined as $pk_0 := r + H(c_0)$ and $pk_1 := r + H(c_1)$. This construction has similarities with the McQuoid, Rosulek and Roy OT [MRR20]. As an additional constraint, we ask that $\hat{H}(r) = c_0 + c_1$, where \hat{H} is a second random oracle. When proving security against \mathcal{R}^* , whenever \mathcal{R}^* makes a query to \hat{H} , the simulator samples a random \hat{c} and programs $H(\hat{c} + c_j) = pk_j^* - r$ for any previous query c_j to H for a public key of its choice. Since \hat{c} is uniform, it is very unlikely that H has been programmed on this input for a previous query. Now we could just rely on the multi-user security of the PKE rather than trying to guess which of the previous queries corresponds to r_b . Nevertheless, \mathcal{R}^* could first query \hat{H} for r and then query H for c_0, c_1 such that $\hat{H}(r) = c_0 + c_1$. This would cause an issue in the programming strategy which assumes that the adversary queries first c_0 or c_1 to H . Further, this strategy does not seem to help \mathcal{R}^* since by using a guessing strategy, we could show that by the security of the PKE, \mathcal{R}^* cannot learn any of the OT strings. However, it seems that we cannot show this via a tight reduction.

We resolve the issue via the following approach. We let the receiver send (r, c_0, c_1) and the public keys are defined as $pk_0 := r + \hat{H}(\hat{c}_0)$ and $pk_1 := r + \hat{H}(\hat{c}_1)$, where $\hat{c}_0 := c_1 + H(r, c_0)$ and $\hat{c}_1 := c_0 + H(r, c_1)$. \hat{c}_0 and \hat{c}_1 could be seen as the r_0, r_1 values of the MR OT. But rather than using them directly, we apply an additional random oracle on them as “correlation breaker”. A PKE scheme is typically not tightly secure in a setting where an adversary A can first suggest q shifts $\tilde{r}_1, \dots, \tilde{r}_q$, then receives public key pk and finally tries to break IND-CPA security under public key $pk - \tilde{r}_j$ where $j \in [q]$ is chosen by A . Though a correlation robust hash function \hat{H} [IKNP03] is tailored to such a setting and maps all inputs $pk - r_1, \dots, pk - r_q$ to strings that do not collide as long as pk is uniform and independent of $\tilde{r}_1, \dots, \tilde{r}_q$. In our setting, we need something stronger than correlation robustness since we also need programmability such that we can program these disjunct strings to different public keys. Fortunately, a random oracle provides both properties such that for any choice of r, c_0, c_1 among the

random oracle queries of \mathcal{R}^* , at least one of the public keys pk_0 and pk_1 will correspond to a programmed key chosen by the simulator. When q is the total amount of random oracle queries, there are at most q^2 choices for r, c_0, c_1 among the queries. This is due to the fact, that for any $b \in \{0, 1\}$, c_b is uniquely defined by r and c_{1-b} . Therefore, there will be at most q^2 choices of public keys pk_0, pk_1 and hence the multi-user security of PKE for q^2 user is sufficient to prove security against a malicious receiver.

For the proof, it would sufficient to just hash r_0, r_1 of the MR OT, though in the actual protocol, we need to allow the receiver to control one of the public keys. For this reason we introduce r to the protocol. Interestingly, our protocol could be seen as a combination of sequential and parallel OR proof techniques.

2 Preliminaries

Notation. For $n \in \mathbb{N}$, we use $[n]$ to denote the set $\{1, \dots, n\}$. We use λ to denote the security parameter. And $x \leftarrow \mathcal{X}, x \leftarrow X$ to sample x from a distribution \mathcal{X} or uniformly random from a set X .

Let Π be a protocol between two parties \mathcal{S} and \mathcal{R} . For two (interactive) algorithms $\mathcal{S}', \mathcal{R}'$ that do not necessarily follow the protocol description of Π , we use $[\mathcal{S}', \mathcal{R}']_{\Pi}$ to denote the interaction between \mathcal{S}' and \mathcal{R}' in protocol Π , where \mathcal{S}' takes the role of \mathcal{S} and \mathcal{R}' the role of \mathcal{R} . For an environment D , we use $\mathsf{D}([\mathcal{S}', \mathcal{R}']_{\Pi})$ to denote an interaction of D with $\mathcal{S}', \mathcal{R}'$ who interact in Π . Here, we follow the simple UC framework of [CCL15].

For a cyclic group \mathcal{G} of order $p \in \mathbb{N}$ with generator g , we use $\llbracket 1 \rrbracket$ to denote g and for $a, b \in \mathbb{N}$, $\llbracket a \rrbracket + b\llbracket 1 \rrbracket = \llbracket a + b \rrbracket$. For $a, b \in \mathbb{Z}_q^\eta$, we use $\langle a, b \rangle$ to denote the inner product between a and b . For an oracle \mathcal{O} and an algorithm A , we use $\mathsf{A}^{\mathcal{O}}$ to denote A when A has query access to \mathcal{O} .

Cryptographic Assumptions.

We recap the DDH and LWE problems below. Since we consider the UC setting, we need to consider non-uniform algorithms which receive an auxiliary input.

Definition 1 (Decisional Diffie-Hellman (DDH)). A ppt algorithm A solves the decisional Diffie-Hellman (DDH) problem for a group \mathcal{G} of order $p \in \mathbb{N}$ with generator $\llbracket 1 \rrbracket$ with probability ϵ if for any polynomial auxiliary input z ,

$$|\Pr[\mathsf{A}(z, \llbracket 1 \rrbracket, \llbracket a \rrbracket, \llbracket b \rrbracket, \llbracket ab \rrbracket) = 1] - \Pr[\mathsf{A}(z, \llbracket 1 \rrbracket, \llbracket a \rrbracket, \llbracket b \rrbracket, \llbracket c \rrbracket) = 1]| \geq \epsilon,$$

where $a, b, c \leftarrow \mathbb{Z}_p$.

Definition 2 (Learning with Errors (LWE)). A ppt algorithm A solves the Learning with Errors (LWE) problem for parameters $q, \eta \in \mathbb{N}$ and noise distribution \mathcal{X} with probability ϵ if for any polynomial auxiliary input z

$$|\Pr[\mathsf{A}^{\mathcal{O}^{\text{LWE}}}(z) = 1] - \Pr[\mathsf{A}^{\mathcal{O}^{\text{u}}}(z) = 1]| \geq \epsilon,$$

where \mathcal{O}_{LWE} is an oracle that outputs samples of the form $a, \langle a, s \rangle + e$ with $a \leftarrow \mathbb{Z}_q^n$, $e \leftarrow \mathcal{X}$ and each sample uses the same secret $s \leftarrow \mathbb{Z}_q^n$. \mathcal{O}_U is the oracle that outputs a, u with $a \leftarrow \mathbb{Z}_q^n$, $u \leftarrow \mathbb{Z}_q$.

Public Key Encryption. We define public key encryption and its multi-user security below. We emphasize that we consider a setting with only a single challenge ciphertext which is a weaker security notion than the commonly used multi-user security setting in which an adversary receives a challenge ciphertext for each public key.

Definition 3 (Public Key Encryption). A public key encryption (PKE) is a triplet of algorithms $(\text{Gen}, \text{Enc}, \text{Dec})$ and a message space \mathcal{M} with the following syntax.

Gen: Takes as input 1^λ and outputs a key pair (sk, pk) .

Enc: Takes as input pk and a message $m \in \mathcal{M}$ and outputs a ciphertext ct .

Dec: Takes as input sk and a ciphertext ct and outputs a message m .

We require correctness and M-IND-CPA security.

Correctness: For any $m \in \mathcal{M}$

$$\Pr[\text{Dec}(\text{sk}, \text{Enc}(\text{pk}, m)) = m] \geq 1 - \text{negl},$$

where $(\text{sk}, \text{pk}) \leftarrow \text{Gen}(1^\lambda)$.

n-Multi-User IND-CPA (M-IND-CPA): For any ppt adversary $A := (A_1, A_2)$ and any polynomial auxiliary input z

$$|\Pr[A_2(\text{st}, \text{ct}_0^*) = 1] - \Pr[A_2(\text{st}, \text{ct}_1^*) = 1]| \leq \text{negl},$$

where for all $i \in [n]$, $(\text{sk}_i, \text{pk}_i) \leftarrow \text{Gen}(1^\lambda)$, $(\text{st}, i^*, m_0, m_1) \leftarrow A_1(z, \text{pk}_1, \dots, \text{pk}_n)$ and for all $b \in \{0, 1\}$ $\text{ct}_b^* \leftarrow \text{Enc}(\text{pk}_{i^*}, m_b)$.

In addition to the multi-user IND-CPA security, we also need that public keys are indistinguishable from uniform in the multi-user setting.

Definition 4 (PKE with Pseudorandom Public Keys). For $n \in \mathbb{N}$, we call a PKE scheme n -multi-user public key indistinguishable (M-IND-PK) over group \mathcal{G} if for any ppt A and polynomial auxiliary input z

$$|\Pr[A(z, \text{pk}_1, \dots, \text{pk}_n) = 1] - \Pr[A(z, u_1, \dots, u_n) = 1]| \leq \text{negl},$$

where for all $i \in [n]$, $(\text{sk}_i, \text{pk}_i) \leftarrow \text{Gen}(1^\lambda)$ and $u_i \leftarrow \mathcal{G}$.

Oblivious Transfer. We use the simplified UC framework which is sufficient for full UC [CCL15]. Below, we define UC secure OT.

Definition 5 (Ideal Oblivious Transfer Functionality). An ideal OT functionality \mathcal{F}_{OT} interacts with two ppt parties \mathcal{S} and \mathcal{R} as follows. \mathcal{F}_{OT} takes s_0, s_1 from \mathcal{S} . \mathcal{F}_{OT} takes b from \mathcal{R} and returns s_b .

Definition 6 (Oblivious Transfer). We call a protocol Π between two ppt parties, a sender \mathcal{S} and a receiver \mathcal{R} , oblivious transfer (OT) if at the end of the protocol they have established a correlation in which \mathcal{S} holds strings (s_0, s_1) and \mathcal{R} holds (b, s_b) . For security, we require two properties with respect to a functionality \mathcal{F}_{OT} .

Security Against a Malicious Sender: For any ppt adversary A , there exists a ppt adversary A' such that for any ppt environment D and any polynomial size auxiliary input z

$$|\Pr[D(z, [A, \mathcal{R}]_{\Pi}) = 1] - \Pr[D(z, [A', \mathcal{F}_{\text{OT}}]_{\Pi}) = 1]| = \text{negl},$$

where all algorithms receive input 1^λ . \mathcal{R} additionally receives input b .

Security Against a Malicious Receiver: For any ppt adversary A , there exists a ppt adversary A' such that for any ppt environment D and any polynomial size auxiliary input z

$$|\Pr[D(z, [\mathcal{S}, A]_{\Pi}) = 1] - \Pr[D(z, [\mathcal{F}_{\text{OT}}, A']_{\Pi}) = 1]| = \text{negl},$$

where all algorithms receive input 1^λ .

3 Public Key Encryption in the Multi User Setting

We use this section to recap commonly known public key encryption schemes that are tightly secure in the multi-user setting. As a proof of concept, we consider ElGamal, Regev encryption and dual Regev encryption.

Definition 7 (ElGamal). The ElGamal PKE over group \mathcal{G} with order $p \in \mathbb{N}$ and generator $[[1]]$ with message space $M := \mathcal{G}$ has the following syntax.

$\text{Gen}([1]) \rightarrow (\text{pk}, \text{sk})$: Sample $x \leftarrow \mathbb{Z}_p$ and output $\text{pk} := [[x]]$ and $\text{sk} := x$.

$\text{Enc}([1], \text{pk}, m) \rightarrow (\text{ct}_1, \text{ct}_2)$: Sample $r \leftarrow \mathbb{Z}_p$ and output $\text{ct}_1 := [[r]]$, $\text{ct}_2 := r\text{pk} + m$.

$\text{Dec}([1], \text{sk}, \text{ct}) \rightarrow m$: Output $m := \text{ct}_2 - \text{sk} \cdot \text{ct}_1$.

It is straightforward to see that ElGamal is perfectly correct. Let us recap that it is tightly secure in the multi-user setting. Due to the fact that the public keys are uniform over \mathcal{G} , ElGamal is perfectly n -M-IND-PK secure.

Lemma 1. Let \mathcal{G} be of prime order and DDH be ϵ hard over \mathcal{G} and n polynomial, then ElGamal over \mathcal{G} is 2ϵ n -M-IND-CPA secure.

Proof. The proof follows straightforwardly from the random selfreducibility of the DDH assumption. The reduction for parameter $d \in \{0, 1\}$ receives a DDH challenge $[[a]]$, $[[b]]$, $[[c]]$ and samples for all $i \in [n]$ $r_i \leftarrow \mathbb{Z}_p$. It forwards z and $\text{pk}_1 := r_1[[a]]$, \dots , $\text{pk}_n := r_n[[a]]$ to A that tries to break ElGamal. When A send i^* , m_0 , m_1 , the reduction sends $\text{ct} := ([[b]], r_{i^*}[[c]] \cdot m_d)$. The reduction outputs the output of A .

When $\llbracket c \rrbracket = \llbracket ab \rrbracket$, ct is an encryption of m_d , i.e. $\text{ct} := \text{ct}_d$, while when c is uniform, ct encrypts a uniform message, i.e. $\text{ct} := \text{ct}_U$. If A distinguishes ct_d from ct_U with probability ϵ' , the reduction solves DDH with probability ϵ' . Assuming that DDH is ϵ hard, A cannot distinguish ct_d from ct_U with $\epsilon' > \epsilon$ for any $d \in \{0, 1\}$ and it cannot distinguish ct_0 from ct_1 with $\epsilon' > 2\epsilon$. \square

Definition 8 (Regev Encryption [Reg05]). *Regev encryption with the parameters $q, \eta, m \in \mathbb{N}$ with $m \geq \eta \log q$ and message space $\{0, 1\}^m$ has the following syntax.*

$\text{Gen}(1^\lambda) \rightarrow (\text{pk}, \text{sk})$: Sample $s \leftarrow \mathbb{Z}_q^\eta$, $A \leftarrow \mathbb{Z}_q^{m \times \eta}$, $e \leftarrow \mathcal{X}^m$ and output $\text{pk} := (A, As + e)$ and $\text{sk} := s$.
 $\text{Enc}(\text{pk}, m) \rightarrow (\text{ct}_1, \text{ct}_2)$: Sample $R \leftarrow \{0, 1\}^{m \times m}$ and output $\text{ct}_1 := R\text{pk}_1$, $\text{ct}_2 := R\text{pk}_2 + m \lfloor \frac{q}{2} \rfloor$.
 $\text{Dec}(\text{sk}, \text{ct}) \rightarrow m$: Compute $\hat{m} := \text{ct}_2 - \text{ct}_1 \cdot \text{sk}$ and output $m := \lfloor \lfloor \frac{2}{q} \hat{m} \rfloor \rfloor$.

For a proper choice of q, m and \mathcal{X} , Regev encryption will be correct.

Lemma 2. *Let LWE be ϵ hard and n polynomial, then Regev encryption is 2ϵ n -M-IND-CPA and ϵ n -M-IND-PK secure.*

Proof. We first show M-IND-CPA security. The reduction for parameter $d \in \{0, 1\}$ receives access to an oracle \mathcal{O} that it uses to generate A_i, b_i for all $i \in [n]$. It sets $\text{pk}_i := (A_i, b_i + A_i s_i)$ for $s_i \leftarrow \mathbb{Z}_q^\eta$ and forwards them to A . After A sends (i^*, m_0, m_1) , the reduction samples $R \leftarrow \{0, 1\}^{m \times m}$ and sends $\text{ct} := (RA_i, R(b_i + A_i s_i) + m \lfloor \frac{q}{2} \rfloor)$. The reduction outputs the output of A .

When $\mathcal{O} = \mathcal{O}_{\text{LWE}}$, ct is an encryption of m_d , i.e. $\text{ct} := \text{ct}_d$, while when $\mathcal{O} = \mathcal{O}_U$, ct is by the leftover hash lemma uniform, i.e. $\text{ct} := \text{ct}_U$. If A distinguishes ct_d from ct_U with probability ϵ' , the reduction solves LWE with probability ϵ' . Assuming that LWE is ϵ hard, A cannot distinguish ct_d from ct_U with $\epsilon' > \epsilon$ for any $d \in \{0, 1\}$ and it cannot distinguish ct_0 from ct_1 with $\epsilon' > 2\epsilon$.

Let us now consider the M-IND-PK security. The reduction defines pk_i as previously. When $\mathcal{O} = \mathcal{O}_{\text{LWE}}$, then pk_i is a proper public key and when $\mathcal{O} = \mathcal{O}_U$, then the public key is uniform. If A can distinguish them, it solves LWE. \square

Definition 9 (Dual Regev Encryption [GPV08]). *Dual Regev encryption with the parameters $q, \eta, m \in \mathbb{N}$ with $m \geq \eta \log q$ and message space $\{0, 1\}^m$ has the following syntax.*

$\text{Gen}(1^\lambda) \rightarrow (\text{pk}, \text{sk})$: Sample $R \leftarrow \{0, 1\}^{m \times m}$, $A \leftarrow \mathbb{Z}_q^{m \times \eta}$ and output $\text{pk} := (A, RA)$ and $\text{sk} := R$.
 $\text{Enc}(\text{pk}, m) \rightarrow (\text{ct}_1, \text{ct}_2)$: Sample $s \leftarrow \mathbb{Z}_q^n$, $e_1, e_2 \leftarrow \mathcal{X}^m$, $R' \leftarrow \{0, 1\}^{m \times m}$ and outputs $\text{ct}_1 := \text{pk}_1 \cdot s + e_1$, $\text{ct}_2 := \text{pk}_2 \cdot s + R' e_2 + m \lfloor \frac{q}{2} \rfloor$.
 $\text{Dec}(\text{sk}, \text{ct}) \rightarrow m$: Compute $\hat{m} := \text{ct}_2 - \text{sk} \cdot \text{ct}_1$ and output $m := \lfloor \lfloor \frac{2}{q} \hat{m} \rfloor \rfloor$.

Correctness follows in the same way as in Regev encryption. By the leftover hash lemma, the public key is statistically indistinguishable from uniform and therefore dual Regev encryption is M-IND-PK secure.

Lemma 3. *Let LWE be ϵ hard and n polynomial, then dual Regev encryption is 2ϵ n -M-IND-CPA secure.*

Proof. The reduction for parameter $d \in \{0, 1\}$ receives access to an oracle \mathcal{O} that it uses to generate A_i, b_i for all $i \in [n]$. It sets $\mathbf{pk}_i := (A_i, R_i A_i)$ for $R_i \leftarrow \{0, 1\}^{m \times m}$ and forwards them to A . After A sends (i^*, m_0, m_1) , the reduction sends $\text{ct} := (b_i, R_i b_i + \mathbf{m} \lfloor \frac{q}{2} \rfloor)$. The reduction outputs the output of A .

When $\mathcal{O} = \mathcal{O}_{\text{LWE}}$, ct is an encryption of \mathbf{m}_d , i.e. $\text{ct} := \text{ct}_d$, while when $\mathcal{O} = \mathcal{O}_{\text{U}}$, ct is by the leftover hash lemma (with leakage Re_2) uniform, i.e. $\text{ct} := \text{ct}_{\text{U}}$. If A distinguishes ct_d from ct_{U} with probability ϵ' , the reduction solves LWE with probability ϵ' . Assuming that LWE is ϵ hard, A cannot distinguish ct_d from ct_{U} with $\epsilon' > \epsilon$ for any $d \in \{0, 1\}$ and it cannot distinguish ct_0 from ct_1 with $\epsilon' > 2\epsilon$. \square

We remark that our security proofs for multi user security require more LWE samples than the proofs of the standard PKE security notions. We emphasize that there are well known techniques to generate many LWE samples from a fixed amount of LWE samples [Reg05, ILL89]. Since such a rerandomization increases the noise level, one needs to start with a lower noise level which decreases the hardness of LWE slightly such that the approximation factor of the underlying SVP instance increases by a factor of $\tilde{O}(n^{1/2})$.

4 Oblivious Transfer from PKE

Theorem 1. *Let PKE be a M-IND-CPA and M-IND-PK secure and correct. Then Protocol 2 is a UC secure OT in the ROM.*

Proof. Given the correctness of PKE, an honest sender and receiver will establish correlation $(s_0, s_1), (b, s_b)$ with overwhelming probability.

We now focus on security against a malicious sender.

Lemma 4. *Let PKE be ϵ_u 1-M-IND-PK secure. Then, for any ppt adversary A , there exists a ppt adversary A' such that for any ppt environment D and any polynomial size auxiliary input z*

$$|\Pr[D(z, [A, \mathcal{R}]_{\Pi}) = 1] - \Pr[D(z, [A', \mathcal{F}_{\text{OT}}]_{\Pi}) = 1]| \leq \epsilon_u,$$

where all algorithms receive input 1^λ . \mathcal{R} additionally receives input b .

Proof. We construct a receiver \mathcal{R}' follows the description of \mathcal{R} by sampling $(\mathbf{pk}_b, \mathbf{sk}_b) \leftarrow \text{Gen}(1^\lambda)$, $\hat{c}_b \leftarrow \{0, 1\}^\lambda$, $c_b \leftarrow \{0, 1\}^\lambda$, computing $r := \mathbf{pk}_b - \hat{H}_b(\hat{c}_b)$, $c_{1-b} := \hat{c}_b \oplus H(r, c_b)$. Unlike \mathcal{R} , \mathcal{R}' computes $\hat{c}_{1-b} := c_b \oplus H_{1-b}(r, c_{1-b})$, samples $(\mathbf{pk}_{1-b}, \mathbf{sk}_{1-b}) \leftarrow \text{Gen}(1^\lambda)$ and programs $\hat{H}_{1-b}(\hat{c}_{1-b}) := \mathbf{pk}_{1-b} - r$. Otherwise, \mathcal{R}' follows the description of \mathcal{R} .

Notice that in case of \mathcal{R} , $r + \hat{H}_{1-b}(\hat{c}_{1-b})$ is uniform while in case of \mathcal{R}' , it has the distribution of a public key generated by Gen . If D can distinguish $[A, \mathcal{R}']$ from $[A, \mathcal{R}]$, then D can be used to break the 1-M-IND-PK security of PKE

Oblivious Transfer Protocol

Primitives:

- PKE scheme ($\text{Gen}, \text{Enc}, \text{Dec}$) with pseudorandom public keys in \mathbb{G} .
- Random oracles
 - $\mathbf{H}_0, \mathbf{H}_1 : \mathbb{G} \times \{0, 1\}^\lambda \rightarrow \{0, 1\}^\lambda$.
 - $\hat{\mathbf{H}}_0, \hat{\mathbf{H}}_1 : \{0, 1\}^\lambda \rightarrow \mathbb{G}$.

Common input: 1^λ .

Sender \mathcal{S} input: s_0, s_1 .

Receiver \mathcal{R} input: $b \in \{0, 1\}$.

1. \mathcal{R} samples $(\mathbf{pk}_b, \mathbf{sk}_b) \leftarrow \text{Gen}(1^\lambda)$, $\hat{c}_b \leftarrow \{0, 1\}^\lambda$, $c_b \leftarrow \{0, 1\}^\lambda$, computes
 - $r := \mathbf{pk}_b - \mathbf{H}_b(\hat{c}_b)$
 - $c_{1-b} := \hat{c}_b \oplus \mathbf{H}_b(r, c_b)$
 and sends (r, c_0, c_1) .
2. \mathcal{S} computes
 - $\hat{c}_0 := c_1 \oplus \mathbf{H}_0(r, c_0)$, $\hat{c}_1 := c_0 \oplus \mathbf{H}_1(r, c_1)$,
 - $\mathbf{pk}_0 := r + \hat{\mathbf{H}}_0(\hat{c}_0)$, $\mathbf{pk}_1 := r + \hat{\mathbf{H}}_1(\hat{c}_1)$,
 - $\mathbf{ct}_0 := \text{Enc}(\mathbf{pk}_0, s_0)$, $\mathbf{ct}_1 := \text{Enc}(\mathbf{pk}_1, s_1)$,
 and sends $(\mathbf{ct}_0, \mathbf{ct}_1)$.
3. \mathcal{R} computes $s_b := \text{Dec}(\mathbf{sk}_b, \mathbf{ct}_b)$.

Fig. 2: Oblivious Transfer in the Random Oracle Model. $(+, -)$ are used to denote the operations in \mathbb{G} . \oplus is the xor operation over $\{0, 1\}^*$.

with probability ϵ_u as follows. The reduction receives a 1-M-IND-PK challenge \mathbf{pk} and sets $\mathbf{pk}_{1-b} := \mathbf{pk}$. When \mathbf{pk} is uniform, it simulates \mathcal{R} and otherwise \mathcal{R}' . Therefore,

$$|\Pr[\mathbf{D}(z, [\mathbf{A}, \mathcal{R}]_\Pi) = 1] - \Pr[\mathbf{D}(z, [\mathbf{A}, \mathcal{R}']_\Pi) = 1]| \leq \epsilon_u.$$

Based on \mathcal{R}' , we can construct an adversary \mathbf{A}' which interacts with \mathbf{A} , relays all interaction between \mathbf{A} and \mathbf{D} and needs to submit s_0 and s_1 to \mathcal{F}_{OT} . \mathbf{A}' follows the process of \mathcal{R}' when constructing r, c_0, c_1 that defines \mathbf{pk}_0 and \mathbf{pk}_1 . As \mathcal{R}' , \mathbf{A}' knows both, \mathbf{sk}_0 and \mathbf{sk}_1 which \mathbf{A}' uses to decrypt \mathbf{ct}_0 and \mathbf{ct}_1 to obtain s_0 and s_1 . Since, \mathbf{A}' follows the description of \mathcal{R}' , it leads to the same interaction between \mathbf{A} and \mathbf{D} . Therefore

$$\Pr[\mathbf{D}(z, [\mathbf{A}, \mathcal{R}']_\Pi) = 1] = \Pr[\mathbf{D}(z, [\mathbf{A}', \mathcal{F}_{\text{OT}}]_\Pi) = 1],$$

which concludes the proof of the lemma. □

We conclude the theorem with the following lemma that establishes security against a malicious receiver.

Lemma 5. *Let PKE be ϵ_u q^2 -M-IND-PK and ϵ_t q^2 -M-IND-CPA secure. Then, for any ppt adversary \mathbf{A} making at most q random oracle queries to $\mathbf{H}_0, \mathbf{H}_1, \hat{\mathbf{H}}_0$*

and \hat{H}_1 combined, there exists a ppt adversary A' such that for any ppt environment D and any polynomial size auxiliary input z

$$|\Pr[D(z, [\mathcal{S}, A]_{\Pi}) = 1] - \Pr[D(z, [\mathcal{F}_{\text{OT}}, A']_{\Pi}) = 1]| \leq \epsilon_u + \epsilon_t + \frac{q^2}{2^\lambda},$$

where all algorithms receive input 1^λ .

Proof. For simplicity, we assume that when A sends r, c_0, c_1 during the protocol to the sender, it has queried the random oracles for $H_0(r, c_0)$, $H_1(r, c_1)$, $\hat{H}_0(\hat{c}_0)$ and $\hat{H}_1(\hat{c}_1)$. We can assume this without loss of generality by making at most 4 additional queries and setting the amount of queries to $\hat{q} = q + 4$. Since this is not significant for our overall bound, we identify \hat{q} with q in the following. We also assume without loss of generality that A queries an oracle only once per input.

We define three intermediate algorithms $\mathcal{S}_1, \mathcal{S}_2, \mathcal{S}_3$ playing the role of sender \mathcal{S} . \mathcal{S}_1 is identical to \mathcal{S} except that it simulates random oracles H_0, H_1 as follows. For all $i \in [q]$ and $j \in [q]$ (where q is the amount of queries), it samples $\text{pk}_{i,j} \leftarrow \mathcal{G}$. Whenever A makes a query $r_i, c_{i,d}$ to H_d for $i \in [q]$ and $d \in \{0, 1\}$, \mathcal{S}_1 samples $H_d(r_i, c_{i,d}) \leftarrow \{0, 1\}^\lambda$ and does the following for any $j \in [q]$ with $j < i$ for which the j th query is a query $r_j, c_{j,1-d}$ to H_{1-d} with $r_j = r_i$.

1. Compute $\hat{c}_{i,j,d} := c_{j,1-d} \oplus H_d(r_i, c_{i,d})$.
2. If $\hat{H}_d(\hat{c}_{i,j,d})$ is defined (through programming or a query), abort. Otherwise, program $\hat{H}_d(\hat{c}_{i,j,d}) := \text{pk}_{i,j} - r_i$.

Afterwards, \mathcal{S}_1 answers the query with $H_d(r_i, c_{i,d})$.

When A sends r, c_0, c_1 , \mathcal{S}_1 computes pk_0, pk_1 in the same way as \mathcal{S} . \mathcal{S}_1 defines b^* such that $\text{pk}_{1-b^*} = \text{pk}_{i,j}$ for a $i \in [q]$ and $j \in [q]$. If no such b^*, i, j exists, \mathcal{S}_1 aborts. Otherwise, it concludes the protocol according to the description of \mathcal{S} .

Let us now consider whether an environment D can distinguish $[A, \mathcal{S}]$ from $[A, \mathcal{S}_1]$. Since $\text{pk}_{i,j}$ are uniform in \mathcal{G} , the output distribution of \hat{H}_d , in particular for every point $\hat{H}_d(\hat{c}_{i,j,d}) := \text{pk}_{i,j} - r_i$ is uniform over \mathcal{G} , in both settings. Other than that, \mathcal{S}_1 differs from \mathcal{S} by two abort conditions - one during queries to H_d and one after seeing (r, c_0, c_1) . Let us assume that \mathcal{S}_1 aborts during a query to H_d . This implies that either A has queried \hat{H}_d for $\hat{c}_{i,j,d} = c_{j,1-d} \oplus H_d(r_i, c_{i,d})$ for an $j \in [q]$ or there exists a $j \in [q]$ and a $j' \in [q] \setminus \{j\}$ with $c_{j,1-d} \oplus H_d(r_i, c_{i,d}) = c_{j',1-d} \oplus H_d(r_i, c_{i,d})$. In the former case, A would predict $H_d(r_i, c_{i,d}) = c_{j,1-d} \oplus \hat{c}_{i,j,d}$ which happens for each query with probability at most $\frac{q}{2^\lambda}$. In the latter case, $c_{j,1-d} = c_{j',1-d}$ and thus A would make the same query twice which we have excluded w.l.o.g. since every adversary queries any input at most once.⁴

The second abort condition never triggers for the following reason. Since A sends r, c_0, c_1 , he will query r, c_0 to H_0 and r, c_1 to H_1 . Let $b^* \in \{0, 1\}$ such

⁴In case an adversary is allowed to query inputs multiple time, \mathcal{S}_1 would simply not try to program the oracle on an input that the adversary has queried already and send the output that is consistent with the previous query for that input.

that A makes query r, c_{b^*} before r, c_{1-b^*} . When A makes query c_{1-b^*}, c_{b^*} will therefore be defined and \mathcal{S}_1 will program $\hat{H}_{1-b^*}(c_{b^*} \oplus H_d(r, c_{1-b^*})) = \text{pk}_{i,j} - r$ for some $i, j \in [q]$. By the definition of pk_{1-b^*} , $\text{pk}_{1-b^*} = \text{pk}_{i,j}$. Thus, we obtain the bound

$$|\Pr[\text{D}(z, [\mathcal{S}, A]_{\Pi}) = 1] - \Pr[\text{D}(z, [\mathcal{S}_1, A]_{\Pi}) = 1]| \leq \frac{q^2}{2\lambda}.$$

\mathcal{S}_2 is identical to \mathcal{S}_1 except that it samples $(\text{pk}_{i,j}, \text{sk}_{i,j}) \leftarrow \text{Gen}(1^\lambda)$ for any $i, j \in [q]$. If there is an environment D that can distinguish $[A, \mathcal{S}_2]$ from $[A, \mathcal{S}_1]$, then we can break the q^2 -M-IND-PK security, i.e. public keys are hard to distinguish from uniform, of PKE as follows. The reduction receives q^2 challenge public keys $\widehat{\text{pk}}_{i,j}$ for $i, j \in [q]$. Instead of sampling $\text{pk}_{i,j}$, it sets $\text{pk}_{i,j} := \widehat{\text{pk}}_{i,j}$.

When the challenge public keys are uniform, the reduction simulates \mathcal{S}_1 and otherwise (when the challenge public keys are distributed according to Gen) \mathcal{S}_2 . Therefore,

$$|\Pr[\text{D}(z, [\mathcal{S}_1, A]_{\Pi}) = 1] - \Pr[\text{D}(z, [\mathcal{S}_2, A]_{\Pi}) = 1]| \leq \epsilon_u.$$

Our next intermediate sender \mathcal{S}_3 follows the description of \mathcal{S}_2 except that after receiving r, c_0, c_1 from A , it defines $\text{ct}_{1-b^*} := \text{Enc}(\text{pk}_{1-b^*}, 0)$. If there is an environment D that can distinguish $[A, \mathcal{S}_2]$ from $[A, \mathcal{S}_3]$, we can break the q^2 -M-IND-CPA security of PKE as follows. The reduction receives q^2 challenge public keys $\widehat{\text{pk}}_{i,j}$ for $i, j \in [q]$. As previously, it sets $\text{pk}_{i,j} := \widehat{\text{pk}}_{i,j}$. It then follows the description of \mathcal{S}_2 until it defines b^* and can compute $\text{pk}_{1-b^*} = \text{pk}_{i,j}$ for some $i, j \in [q]$. The reduction sends $((i, j), \mathbf{m}_0 := s_{1-b^*}, \mathbf{m}_1 := 0)$ to the M-IND-CPA challenger and receives back ct^* . It then sets $\text{ct}_{1-b^*} := \text{ct}^*$. When ct^* encrypts s_{1-b^*} , the reduction simulates \mathcal{S}_2 and otherwise \mathcal{S}_3 . Therefore,

$$|\Pr[\text{D}(z, [\mathcal{S}_2, A]_{\Pi}) = 1] - \Pr[\text{D}(z, [\mathcal{S}_3, A]_{\Pi}) = 1]| \leq \epsilon_t.$$

Based on \mathcal{S}_3 , we can define A' which interacts with A , relays all interaction between A and D and submits b^* to \mathcal{F}_{OT} and then receives s_{b^*} which is used to generate ct_{b^*} . Since A' follows the description of \mathcal{S}_3 , it leads to the same interaction between A and D . Therefore, we can conclude the lemma with

$$\Pr[\text{D}(z, [\mathcal{S}_3, A]_{\Pi}) = 1] = \Pr[\text{D}(z, [\mathcal{F}_{\text{OT}}, A']_{\Pi}) = 1].$$

□

Lemma 4 and Lemma 5 are sufficient to establish Theorem 1. □

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