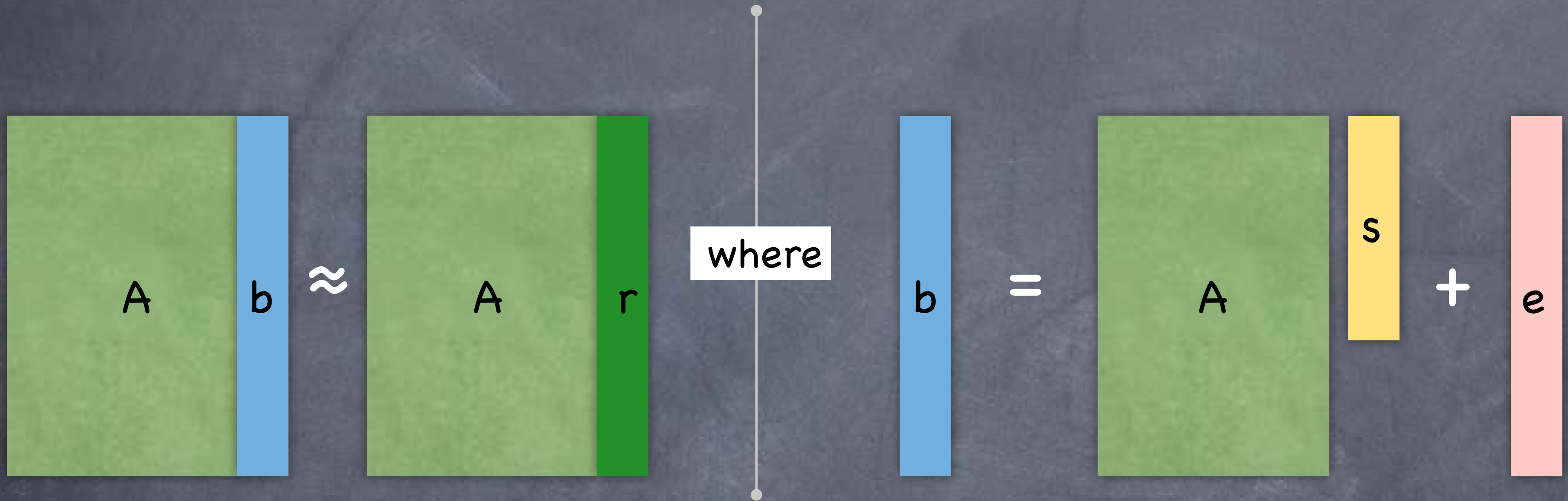
Lattice Cryptography: Towards Fully Homomorphic Encryption

Lecture 20

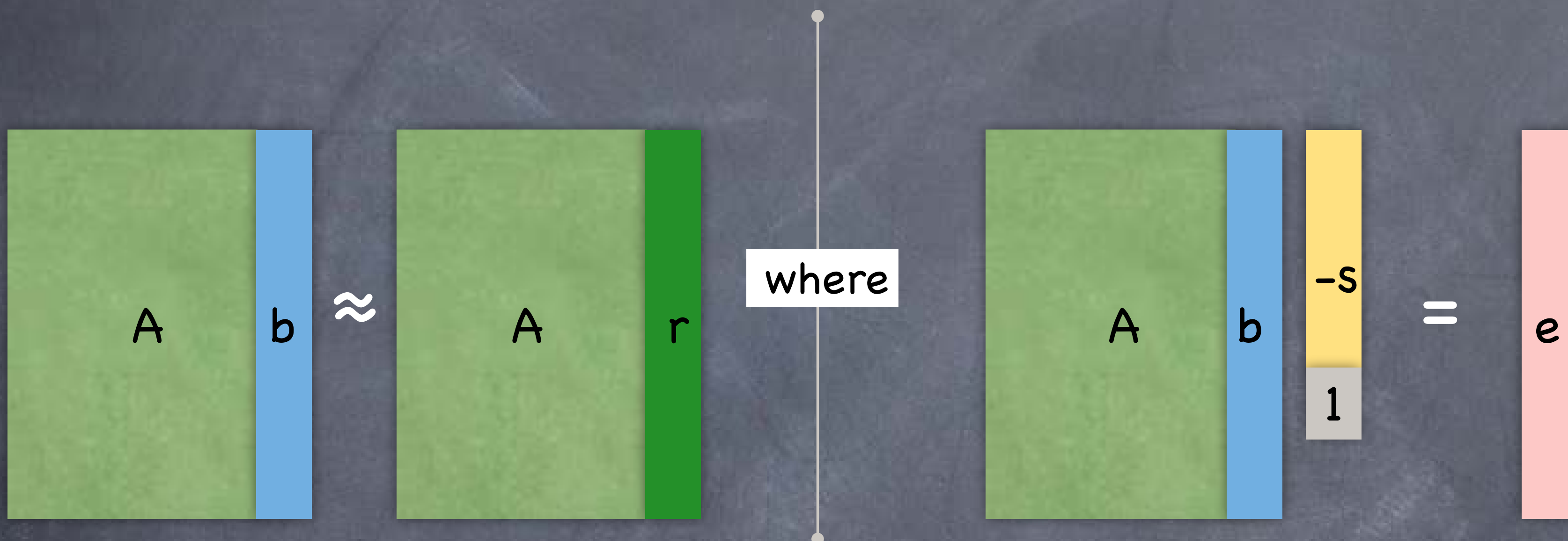
Recall

Learning With Errors



- LWE (decision version): $(A, A\underline{\mathbf{s}}+\underline{\mathbf{e}}) \approx (A, \underline{\mathbf{r}})$, where A random matrix in $A \in \mathbb{Z}_q^{m \times n}$, $\underline{\mathbf{s}}$ uniform, $\underline{\mathbf{e}}$ has "small" entries from a Gaussian distribution, and $\underline{\mathbf{r}}$ uniform.
- Average-case solution for LWE $\Rightarrow$ Worst-case solution for GapSVP (for appropriate choice of parameters)

Learning With Errors



- LWE (decision version): $(A, A\underline{s}+\underline{e}) \approx (A, \underline{r})$, where A random matrix in $A \in \mathbb{Z}_q^{m\times n}$, $\underline{s}$ uniform, $\underline{e}$ has "small" entries from a Gaussian distribution, and $\underline{r}$ uniform.
- Average-case solution for LWE $\Rightarrow$ Worst-case solution for GapSVP (for appropriate choice of parameters)

Learning With Errors

- i.e., a pseudorandom matrix $M \in \mathbb{Z}_q^{m\times n'}$ and non-zero $\underline{\mathbf{z}} \in \mathbb{Z}_q^{n'}$ s.t. entries of $M\underline{\mathbf{z}}$ are all small (writing $n'=n+1$)

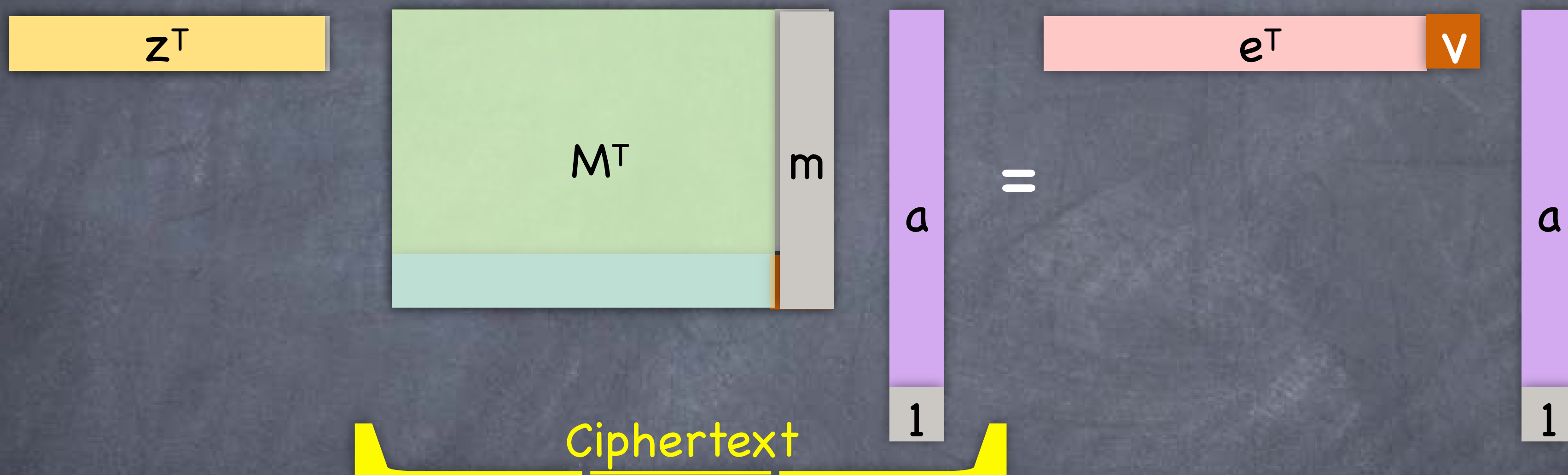
Recall

PKE from LWE

$$\begin{bmatrix} -s^T & 1 \end{bmatrix} \underbrace{\begin{bmatrix} A^T & 0 \\ b^T & v \end{bmatrix} \begin{bmatrix} a \\ 1 \end{bmatrix}}_{\text{Ciphertext}} = \begin{bmatrix} e^T & v \end{bmatrix} \begin{bmatrix} a \\ 1 \end{bmatrix}$$

Recall

PKE from LWE



- Ciphertext = $M^T\mathbf{a} + \mathbf{m}$ where $\mathbf{m}$ encodes the message and $\mathbf{a} \in \{0,1\}^m$
- Decryptng: From $\mathbf{z}^T(M^T\mathbf{a} + \mathbf{m}) = \mathbf{e}^T\mathbf{a} + \mathbf{z}^T\mathbf{m}$ where $\mathbf{e}^T\mathbf{a}$ is small. To allow decoding from this for, say $\mu \in \{0,1\}$, note $\mathbf{z}^T\mathbf{m} = v \approx \mu(q/2)$.
- CPA security: $M^T\mathbf{a}$ is pseudorandom
  - Claim: If $M \in \mathbb{Z}_q^{m \times n'}$ is truly random, $\mathbf{a} \in \{0,1\}^m \setminus \{0^m\}$, $m >> n' \log q$, then $M^T\mathbf{a}$ is very close to being uniform

Randomness Extraction

- Entries in $\underline{\mathbf{a}}$ are not uniformly random over $\mathbb{Z}_q^m$, but concentrated on a small subset $\{0,1\}^m$. We need $M^T\underline{\mathbf{a}}$ to be uniform over $\mathbb{Z}_q^{n'}$
- Follows from two more generally useful facts:
  - $H_M(\underline{\mathbf{a}}) = M^T\underline{\mathbf{a}}$ is a 2-Universal Hash Function (for non-zero $\underline{\mathbf{a}}$)
  - If H is a 2-UHF, then it is a good **randomness extractor**
    - If m >> n′ log q, the entropy of $\underline{\mathbf{a}}$ (m bits) is significantly more than that of a uniform vector in $\mathbb{Z}_q^{n'}$ and a good randomness extractor will produce an almost uniform output

Universal Hashing

- Combinatorial HF: $A \rightarrow (x,y)$; $h \leftarrow \mathcal{H}$. $h(x)=h(y)$ w.n.p
- Even better: 2-Universal Hash Functions
  - "Uniform" and "Pairwise-independent"
  - $\forall x,z \ \Pr_{h \leftarrow \mathcal{H}} [\ h(x)=z\ ] = 1/|Z|$ (where $h:X \rightarrow Z$)
  - $\forall x \neq y, w, z \ \Pr_{h \leftarrow \mathcal{H}} [\ h(x)=w,\ h(y)=z\ ] = 1/|Z|^2$
    - $\Rightarrow \forall x \neq y \ \Pr_{h \leftarrow \mathcal{H}} [\ h(x)=h(y)\ ] = 1/|Z|$

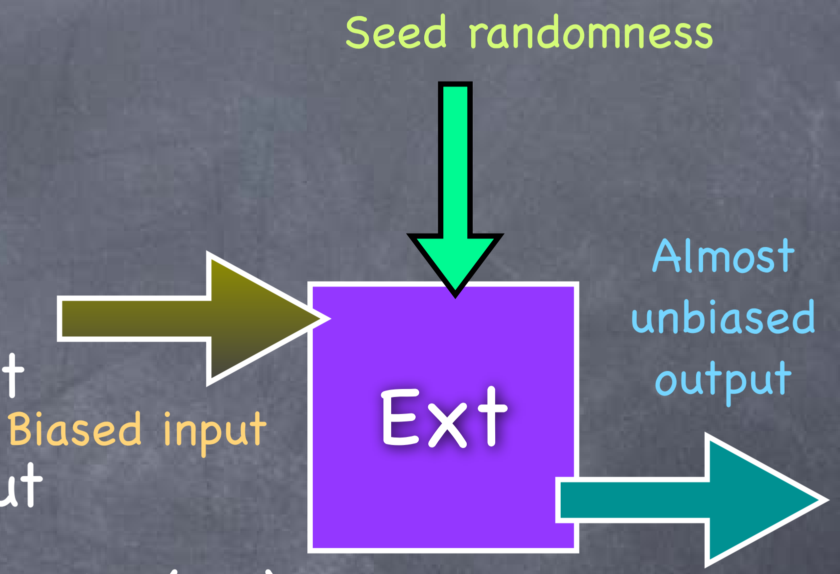
x	$h_1(x)$	$h_2(x)$	$h_3(x)$	$h_4(x)$
0	0	0	1	1
1	0	1	0	1
2	1	0	0	1

Negligible collision-probability if super-polynomial-sized range

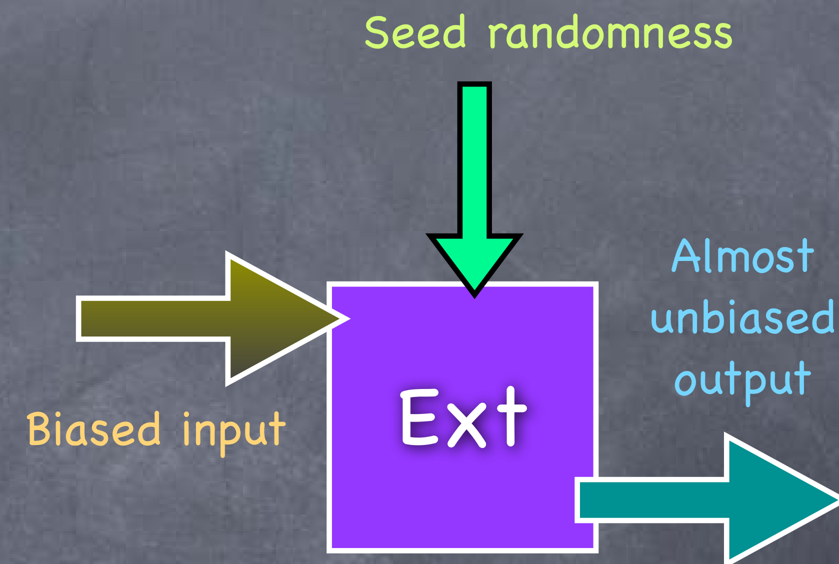
- e.g. $h_{a,b}(x) = ax+b$ (in a finite field, $X=Z$)
  - $\Pr_{a,b} [\ ax+b = z\ ] = \Pr_{a,b} [\ b = z-ax\ ] = 1/|Z|$
  - $\Pr_{a,b} [\ ax+b = w,\ ay+b = z] = ?$ Exactly one (a,b) satisfying the two equations (for $x \neq y$)
    - $\Pr_{a,b} [\ ax+b = w,\ ay+b = z] = 1/|Z|^2$
- Exercise: $M\underline{\mathbf{x}}$ (M random matrix) is a 2-UHF for non-zero boolean $\underline{\mathbf{x}}$

Randomness Extractor

- Input has high "min-entropy"
  - i.e., probability of any particular input string is very low
- Seed uniform and independent of input
- Output vector is shorter than the input
  - Need input min-entropy > output length (1+ε)
- Ext(inp,seed)) ≈ Uniform
  - Statistical closeness
- A **strong extractor**: (seed, Ext(inp,seed)) ≈ (seed,Uniform)
  - i.e., for any input distribution with enough min-entropy, most choices of seed yield a good deterministic extractor



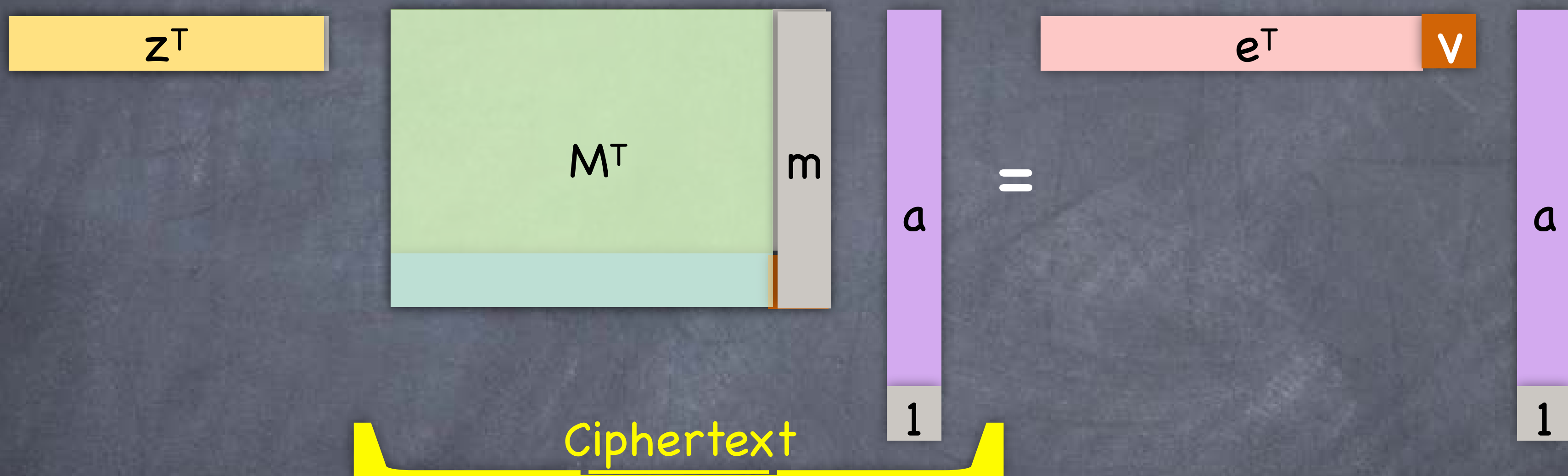
Randomness Extractor



- Leftover Hash Lemma:
  - Any 2-UHF is a strong extractor that can extract almost all of the min-entropy in the input
- A very useful result
  - We need only a special case here:
    - Only for a particular 2-UHF ($H_M(\underline{\mathbf{x}}) = M\underline{\mathbf{x}}$)
    - Only for a particular input distribution ($\underline{\mathbf{x}}$ uniform over $\{0,1\}^m$)
      - Input min-entropy = m bits
      - If $M \in \mathbb{Z}_q^{m \times n}$ then need $m > n \log q\ (1+\varepsilon)$

Recall

PKE from LWE



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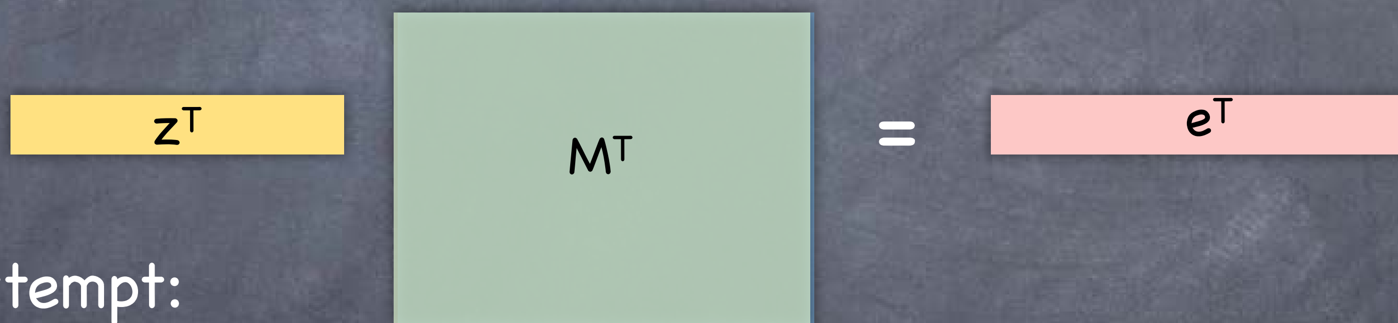
Gentry-Sahai-Waters

- Want to allow homomorphic operations on the ciphertext
- Idea: Ciphertext is a matrix masked by a pseudorandom matrix that can be "annihilated" with secret key. Addition and multiplication of messages given by addition and multiplication of ciphertexts.
- Recall from LWE: $M \in \mathbb{Z}_q^{m \times n}$ and $\underline{\mathbf{z}} \in \mathbb{Z}_q^n$ s.t. $M\underline{\mathbf{z}}$ has small entries



- First attempt:
  - Public-Key = M, Secret-key = $\underline{\mathbf{z}}$
  - $\mathrm{Enc}(\mu) = M^T R + \mu I$ where $\mu \in \{0,1\}$, $R \leftarrow \{0,1\}^{m \times n}$, and $I_{n \times n}$ identity
  - Security: LWE (and LHL) $\Rightarrow$ $M^T R$ is pseudorandom
  - $\mathrm{Dec}_z(C)$: $z^T C = e^T R + \mu z^T$ has "error" $\underline{\delta}^T = e^T R$. Can recover μ since error has small entries (w.h.p.)

Gentry-Sahai-Waters

- First attempt:
  - $\text{Enc}(\mu) = M^T R + \mu I$
  - $\text{Dec}_z(C)$: $z^T C = e^T R + \mu z^T$ has error $\underline{\delta}^T = e^T R$
  - $C_1+C_2 = M^T(R_1+R_2) + (\mu_1+\mu_2) I$ has error $\underline{\delta}^T = \underline{\delta}_1^T + \underline{\delta}_2^T$
    - Error adds up with each operation
    - OK if there is an a priori bound on the <u>depth</u> of computation: Levelled Homomorphic Encryption
  - $C_1 \times C_2$: Error = ?
    - $z^T C_1 C_2 = (\underline{\delta}_1^T + \mu_1 z^T) C_2 = \underline{\delta}_1^T C_2 + \mu_1(\underline{\delta}_2^T + \mu_2 z^T)$
    - Error = $\underline{\delta}_1^T C_2 + \mu_1\, \underline{\delta}_2^T$
    - Problem: Entries in $\underline{\delta}_1^T C_2$ may not be small, as entries in C_2 are not small! (Since $\mu_1 \in \{0,1\}$, $\mu_1 \underline{\delta}_2^T$ does have small entries)

Gentry-Sahai-Waters

- Problem: Entries in $\underline{\delta}_1^T C_2$ may not be small
- Solution Idea: Represent ciphertext as bits!
  - But homomorphic operations will be affected
  - Observation: Reconstructing a number from bits is a linear operation
  - If $\alpha \in \mathbb{Z}_q^m$ has bit-representation $B(\alpha) \in \{0,1\}^{km}$ $(k=O(\log q))$, then $G\ B(\alpha) = \alpha$, where $G \in \mathbb{Z}_q^{m \times km}$ (all operations in $\mathbb{Z}_q$)
    - B can be applied to matrices also as $B : \mathbb{Z}_q^{m \times n} \rightarrow \mathbb{Z}_q^{km \times n}$ and we have $G\ B(\alpha) = \alpha$

Gentry-Sahai-Waters

The Actual Scheme

- Supports messages $\mu \in \{0,1\}$ and NAND operations up to an a priori bounded depth of NANDs
- Public key: Pseudorandom $M \in \mathbb{Z}_q^{m\times n}$ s.t. $m \gg n \log q$
  Private key: non-zero $\underline{z}$ s.t. $M\underline{z}$ has small entries
- **Enc(μ)** $= M^TR + \mu G$ where $R \leftarrow \{0,1\}^{m\times kn}$ and $G \in \mathbb{Z}_q^{n\times kn}$

  (G is the matrix to reverse bit-decomposition)
- **Dec$_z$(C)** : $\underline{z}^TC = \underline{\delta}^T + \mu\underline{z}^TG$ where $\underline{\delta}^T = e^TR$

  Decrypting G yields 1
- **NAND(C_1,C_2)** : $G - C_1 \cdot B(C_2)$
  - $\underline{z}^TC_1 \cdot B(C_2) = \underline{z}^TC_1 \cdot B(C_2) = (\underline{\delta}_1^T + \mu_1\underline{z}^TG)\ B(C_2)$

    $= \underline{\delta}_1^TB(C_2) + \mu_1\underline{z}^TC_2 = \underline{\delta}^T + \mu_1\mu_2\underline{z}^TG$

    where $\underline{\delta}^T = \underline{\delta}_1^TB(C_2) + \mu_1\underline{\delta}_2^T$ has small entries

    Only "left depth" counts, since $\underline{\delta} \leq k\cdot n\cdot\underline{\delta}_1 + \underline{\delta}_2$
  - In general, error gets multiplied by kn. Allows depth $\approx \log_{kn} q$