

5

Pseudorandom functions from pseudorandom generators

We have seen that PRF's (pseudorandom functions) are extremely useful, and we'll see some more applications of them later on. But are they perhaps too amazing to exist? Why would someone imagine that such a wonderful object is feasible? The answer is the following theorem:

Theorem 5.1 — The PRF Theorem. Suppose that the PRG Conjecture is true, then there exists a secure PRF collection $\{f_s\}_{s \in \{0,1\}^*}$ such that for every $s \in \{0,1\}^n$, f_s maps $\{0,1\}^n$ to $\{0,1\}^n$.

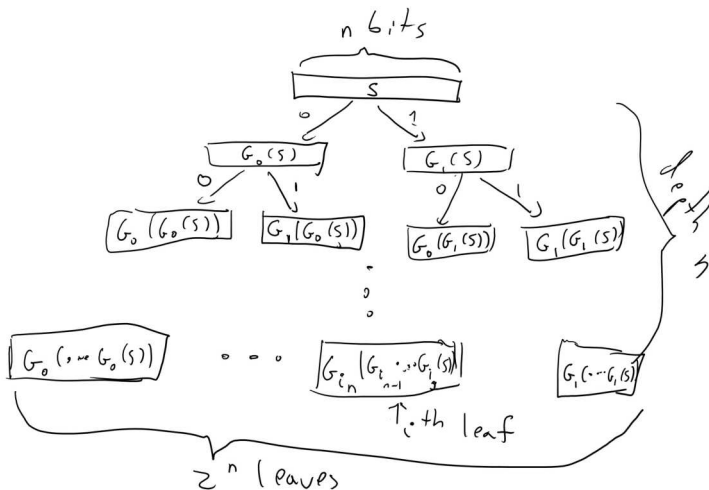


Figure 5.1: The construction of a pseudorandom function from a pseudorandom generator can be illustrated by a depth n binary tree. The root is labeled by the seed s and for every internal node v labeled by a string $x \in \{0,1\}^n$, we label the two children of v by $G_0(x)$ and $G_1(x)$ respectively. The output of the function f_s on input i is the label of the i^{th} leaf.

Proof. If the PRG Conjecture is true then in particular by the length extension theorem there exists a PRG $G : \{0,1\}^n \rightarrow \{0,1\}^{2n}$ that

maps n bits into $2n$ bits. Let's denote $G(s) = G_0(s) \circ G_1(s)$ where \circ denotes concatenation. That is, $G_0(s)$ denotes the first n bits and $G_1(s)$ denotes the last n bits of $G(s)$.

For $i \in \{0,1\}^n$, we define $f_s(i)$ as

$$G_{i_n}(G_{i_{n-1}}(\dots G_{i_1}(s))). \tag{5.1}$$

This might be a bit hard to parse and is easier to understand by looking at Fig. 5.1.

By the definition above we can see that to evaluate $f_s(i)$ we need to evaluate the pseudorandom generator n times on inputs of length n , and so if the pseudorandom generator is efficiently computable then so is the pseudorandom function. Thus, "all" that's left is to prove that the construction is secure and this is the heart of this proof.

I've mentioned before that the first step of writing a proof is convincing yourself that the statement is true, but there is actually an often more important zeroth step which is understanding what the statement actually *means*. In this case what we need to prove is the following:

Given an adversary A that can distinguish in time T a black box for $f_s(\cdot)$ from a black-box for a random function with advantage ϵ , we need to come up with an adversary D that can distinguish in time $\text{poly}(T)$ an input of the form $G(s)$ (where s is random in $\{0,1\}^n$) from an input of the form y where y is random in $\{0,1\}^{2n}$ with bias at least $\epsilon/\text{poly}(T)$.

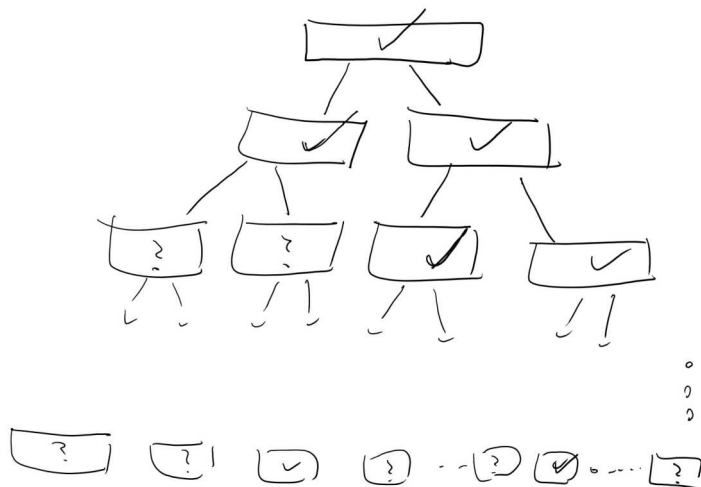


Figure 5.2: In the "lazy evaluation" implementation of the black box to the adversary, we label every node in the tree only when we need it. In this figure check marks correspond to nodes that have been labeled and question marks to nodes that are still unlabeled.

Let us consider the “lazy evaluation” implementation of the black box for A illustrated in Fig. 5.2. That is, at every point in time there are nodes in the full binary tree that are labeled and nodes which we haven’t yet labeled. When A makes a query i , this query corresponds to the path $i_1 \dots i_n$ in the tree. We look at the lowest (furthest away from the root) node v on this path which has been labeled by some value y , and then we continue labelling the path from v downwards until we reach i . In other words, we label the two children of v by $G_0(y)$ and $G_1(y)$, and then if the path i involves the first child then we label its children by $G_0(G_0(y))$ and $G_1(G_0(y))$, and so on and so forth (see Fig. 5.3).

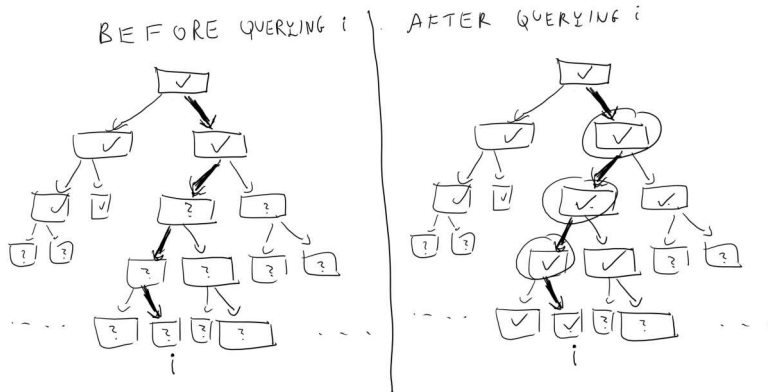


Figure 5.3: When the adversary queries i , the oracle takes the path from i to the root and computes the generator on the minimum number of internal nodes that is needed to obtain the label of the i^{th} leaf

A moment’s thought shows that this is just another (arguably cumbersome) way to describe the oracle that simply computes the map $i \mapsto f_s(i)$. And so the experiment of running A with this oracle produces precisely the same result as running A with access to $f_s(\cdot)$. Note that since A has running time at most T , the number of times our oracle will need to label an internal node is at most $T' \leq 2nT$ (since we label at most $2n$ nodes for every query i).

We now define the following T' hybrids: in the j^{th} hybrid, we run this experiment but in the first j times the oracle needs to label internal nodes, then instead of labelling the b^{th} child of v by $G_b(y)$ (where y is the label of v), the oracle simply labels it by a random string in $\{0,1\}^n$.

Note that the 0^{th} hybrid corresponds to the case where the oracle implements the function $i \mapsto f_s(i)$ will in the T'^{th} hybrid all labels are random and hence the oracle implements a random function. By the hybrid argument, if A can distinguish between the 0^{th} hybrid and the T'^{th} hybrid with bias ϵ then there must exist some j such that it distinguishes between the j^{th} hybrid and the $(j+1)^{\text{st}}$ hybrid with bias

at least ϵ/T' . We will use this j and A to break the pseudorandom generator.

We can now describe our distinguisher D for the pseudorandom generator. On input a string $y \in \{0,1\}^{2n}$ D will run A and the j^{th} oracle inside its belly with one difference- when the time comes to label the j^{th} node, instead of doing this by applying the pseudorandom generator to the label of its parent v (which is what should happen in the j^{th} oracle) it uses its input y to label the two children of v .

Now, if y was completely random then we get exactly the distribution of the $j + 1^{\text{st}}$ oracle, and hence in this case D simulates internally the $j + 1^{\text{st}}$ hybrid. However, if $y = G(s)$ for a random $s \in \{0,1\}^n$ it might not be obvious if we get the distribution of the j^{th} oracle, since in that oracle the label for the children of v was supposed to be the result of applying the pseudorandom generator to the label of v and not to some other random string. However, because v was labeled *before* the j^{th} step then we know that it was actually labeled by a random string. Moreover, since we use lazy evaluation we know that step j is the *first* time where we actually use the value of the label of v . Hence, if at this point we *resampled* this label and used a completely independent random string then the distribution would be *identical*. Hence if $y = G(U_n)$ then D actually does simulate the distribution of the j^{th} hybrid in its belly, and thus if A had advantage ϵ in breaking the PRF $\{f_s\}$ then D will have advantage ϵ/T' in breaking the PRG G thus obtaining a contradiction.

This proof is ultimately not very hard but is rather confusing. I urge you to also look at the proof of this theorem as is written in Section 7.5 (pages 265-269) of the KL book. ■

R **PRF's in practice** While this construction reassures us that we can rely on the existence of pseudorandom functions even on days where we remember to take our meds, this is not the construction people use when they need a PRF in practice because it is still somewhat inefficient, making n calls to the underlying pseudorandom generators. There are constructions (e.g., HMAC) based on hash functions that require stronger assumptions but can use as few as two calls to the underlying function. We will cover these constructions when we talk about hash functions and the random oracle model. One can also obtain practical constructions of PRFs from *block ciphers*, which we'll see later in this lecture.

5.1 Securely encrypting many messages - chosen plaintext security

Let's get back to our favorite task of *encryption*. We seemed to have nailed down the definition of secure encryption, or did we?



Try to think what kind of security guarantees are *not* provided by the notion of computational secrecy we saw in [Definition 2.4](#)

Our current definition requires talks about encrypting a *single* message, but this is not how we use encryption in the real world. Typically, Alice and Bob (or Amazon and Boaz) setup a shared key and then engage in many back and forth messages between one another. At first, we might think that this issue of a single long message vs. many short ones is merely a technicality. After all, if Alice wants to send a sequence of messages (m_1, m_2, \dots, m_t) to Bob, she can simply treat them as a single long message. Moreover, the way that *stream ciphers* work, Alice can compute the encryption for the first few bits of the message she decides what will be the next bits and so she can send the encryption of m_1 to Bob and later the encryption of m_2 . There is some truth to this sentiment, but there are issues with using stream ciphers for multiple messages. For Alice and Bob to encrypt messages in this way, they must maintain a *synchronized shared state*. If the message m_1 was dropped by the network, then Bob would not be able to decrypt correctly the encryption of m_2 .

There is another way in which treating many messages as a single tuple is unsatisfactory. In real life, Eve might be able to have some impact on *what* messages Alice encrypts. For example, the Katz-Lindell book describes several instances in world war II where Allied forces made particular military manouvers for the sole purpose of causing the axis forces to send encryptions of messages of the Allies' choosing. To consider a more modern example, today Google uses encryption for all of its search traffic including (for the most part) the *ads* that are displayed on the page. But this means that an attacker, by paying Google, can cause it to encrypt arbitrary text of their choosing. This kind of attack, where Eve *chooses* the message she wants to be encrypted is called a *chosen plaintext attack*. You might think that we are already covering this with our current definition that requires security for *every* pair of messages and so in particular this pair could chosen by Eve. However, in the case of multiple messages, we would want to allow Eve to be able to choose m_2 *after* she saw the

encryption of m_1 .

All that leads us to the following definition, which is a strengthening of our definition of computational security:

Definition 5.2 — Chosen Plaintext Attack (CPA) secure encryption. An encryption scheme (E, D) is *secure against chosen plaintext attack (CPA secure)* if for every polynomial time *Eve*, *Eve* wins with probability at most $1/2 + \text{negl}(n)$ in the game defined below:

1. The key k is chosen at random in $\{0, 1\}^n$ and fixed.
2. *Eve* gets the length of the key 1^n as input.¹
3. *Eve* interacts with E for $t = \text{poly}(n)$ rounds as follows: in the i^{th} round, *Eve* chooses a message m_i and obtains $c_i = E_k(m_i)$.
4. Then *Eve* chooses two messages m_0, m_1 , and gets $c^* = E_k(m_b)$ for $b \leftarrow_R \{0, 1\}$.
5. *Eve* *wins* if she outputs b .

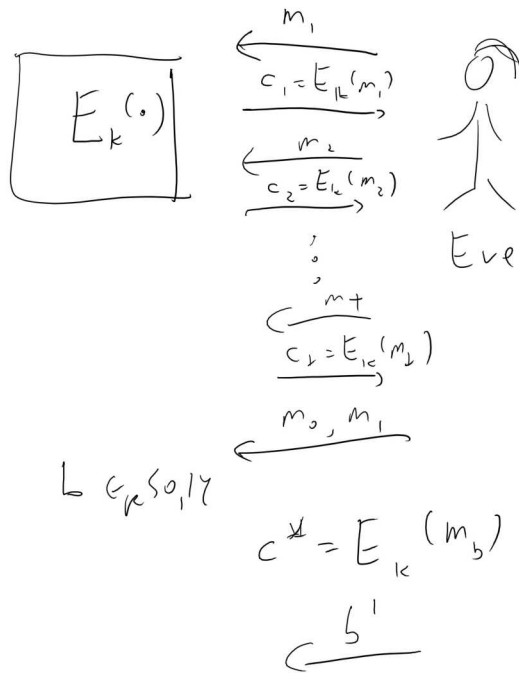


Figure 5.4: In the CPA game, *Eve* interacts with the encryption oracle and at the end chooses m_0, m_1 , gets an encryption $c^* = E_k(m_b)$ and outputs b' . She *wins* if $b' = b$

Definition 5.2 is illustrated in Fig. 5.4. Our previous notion of computational secrecy (i.e., Definition 2.4) corresponds to the case that we skip Step 3 above. Since Step 3 only gives the adversary more power (and hence is only more likely to win), CPA security (Defini-

¹ Giving *Eve* the key as a sequence of n 1's as opposed to in binary representation is a common notational convention in cryptography. It makes no difference except that it makes the input length for *Eve* of length n , which makes sense since we want to allow *Eve* to run in $\text{poly}(n)$ time.

tion 5.2) is *stronger* than computational secrecy (Definition 2.4), in the sense that every CPA secure encryption (E, D) is also computationally secure. It turns out that CPA security is *strictly stronger*, in the sense that without modification, our stream ciphers cannot be CPA secure. In fact, we have a stronger, and initially somewhat surprising theorem:

Theorem 5.3 — CPA security requires randomization. There is no CPA secure (E, D) where E is *deterministic*.

Proof. The proof is very simple: Eve will only use a single round of interacting with E where she will ask for the encryption c_1 of 0^ℓ . In the second round, Eve will choose $m_0 = 0^\ell$ and $m_1 = 1^\ell$, and get $c^* = E_k(m_b)$ she will then output 0 if and only if $c^* = c_1$. ■

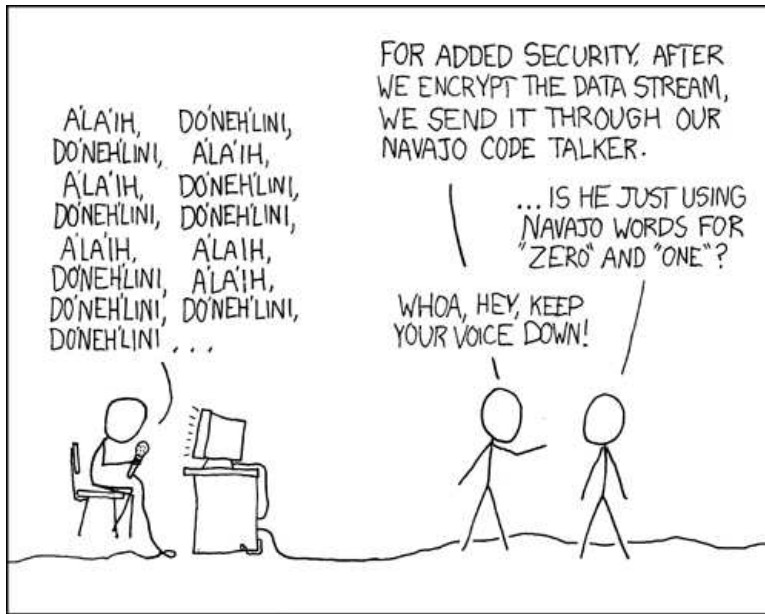


Figure 5.5: Insecurity of deterministic encryption

This proof is so simple that you might think it shows a problem with the definition, but it is actually a real problem with security. If you encrypt many messages and some of them repeat themselves, it is possible to get significant information by seeing the repetition pattern (see the XKCD cartoon again, see Fig. 5.5). To avoid this issue we need to use a *randomized* (or *probabilistic*) encryption, such that if we encrypt the same message twice we *won't* see two copies of the same ciphertext.² But how do we do that? Here pseudorandom functions come to the rescue:

² If the messages are guaranteed to have *high entropy* which roughly means that the probability that a message repeats itself is negligible, then it is possible to have a secure deterministic private-key encryption, and this is sometimes used in practice. (Though often some sort of randomization or padding is added to ensure this property, hence in effect creating a randomized encryption.) Deterministic encryptions can sometimes be useful for applications such as efficient queries on encrypted databases. See [this lecture](#) in Dan Boneh's coursera course.

Theorem 5.4 — CPA security from PRFs. Suppose that $\{f_s\}$ is a PRF collection where $f_s : \{0, 1\}^n \rightarrow \{0, 1\}^\ell$, then the following is a CPA secure encryption scheme: $E_s(m) = (r, f_s(r) \oplus m)$ and $D_s(r, z) = f_s(r) \oplus z$.

Proof. I leave to you to verify that $D_s(E_s(m)) = m$. We need to show the CPA security property. As is usual in PRF-based constructions, we first show that this scheme will be secure if f_s was an actually random function, and then use that to derive security.

Consider the game above when played with a completely random function and let r_i be the random string chosen by E in the i^{th} round and r^* the string chosen in the last round. We start with the following simple but crucial claim:

Claim: The probability that $r^* = r_i$ for some i is at most $T/2^n$.

Proof of claim: For any particular i , since r^* is chosen independently of r_i , the probability that $r^* = r_i$ is 2^{-n} . Hence the claim follows from the union bound. QED

Given this claim we know that with probability $1 - T/2^n$ (which is $1 - \text{negl}(n)$), the string r^* is distinct from any string that was chosen before. This means that by the lazy evaluation principle, if $f_s(\cdot)$ is a completely random function then the value $f_s(r^*)$ can be thought of as being chosen at random in the final round independently of anything that happened before. But then $f_s(r^*) \oplus m_b$ amounts to simply using the one-time pad to encrypt m_b . That is, the distributions $f_s(r^*) \oplus m_0$ and $f_s(r^*) \oplus m_1$ (where we think of r^*, m_0, m_1 as fixed and the randomness comes from the choice of the random function $f_s(\cdot)$) are both equal to the uniform distribution U_n over $\{0, 1\}^n$ and hence Eve gets absolutely no information about b .

This shows that if $f_s(\cdot)$ was a random function then Eve would win the game with probability at most $1/2$. Now if we have some efficient Eve that wins the game with probability at least $1/2 + \epsilon$ then we can build an adversary A for the PRF that will run this entire game with black box access to $f_s(\cdot)$ and will output 1 if and only if Eve wins. By the argument above, there would be a difference of at least ϵ in the probability it outputs 1 when $f_s(\cdot)$ is random vs when it is pseudorandom, hence contradicting the security property of the PRF. ■

5.2 Pseudorandom permutations / block ciphers

Now that we have pseudorandom functions, we might get greedy and want such functions with even more magical properties. This is where the notion of *pseudorandom permutations* comes in.

Definition 5.5 — Pseudorandom permutations. Let $\ell : \mathbb{N} \rightarrow \mathbb{N}$ be some function that is polynomially bounded (i.e., there are some $0 < c < C$ such that $n^c < \ell(n) < n^C$ for every n). A collection of functions $\{f_s\}$ where $f_s : \{0, 1\}^\ell \rightarrow \{0, 1\}^\ell$ for $\ell = \ell(|s|)$ is called a *pseudorandom permutation (PRP) collection* if:

1. It is a pseudorandom function collection (i.e., the map $s, x \mapsto f_s(x)$ is efficiently computable and there is no efficient distinguisher between $f_s(\cdot)$ with a random s and a random function).
2. Every function f_s is a permutation of $\{0, 1\}^\ell$ (i.e., a one to one and onto map).
3. There is an efficient algorithm that on input s, y returns $f_s^{-1}(y)$.

The parameter n is known as the *key length* of the pseudorandom permutation collection and the parameter $\ell = \ell(n)$ is known as the *input length* or *block length*. Often, $\ell = n$ and so in most cases you can safely ignore this distinction.

P At first look [Definition 5.5](#) might seem not to make sense, since on one hand it requires the map $x \mapsto f_s(x)$ to be a permutation, but on the other hand it can be shown that with high probability a random map $H : \{0, 1\}^\ell \rightarrow \{0, 1\}^\ell$ will *not* be a permutation. How can then such a collection be pseudorandom? The key insight is that while a random map might not be a permutation, it is not possible to distinguish with a polynomial number of queries between a black box that computes a random function and a black box that computes a random permutation. Understanding why this is the case, and why this means that [Definition 5.5](#) is reasonable, is crucial to getting intuition to this notion, and so I suggest you pause now and make sure you understand these points.

As usual with a new concept, we want to know whether it is possible to achieve and is useful. The former is established by the following theorem:

Theorem 5.6 — PRP's from PRFs. If the PRF conjecture holds (and hence by [Theorem 5.1](#) also if the PRG conjecture holds) then there exists a pseudorandom permutation collection.

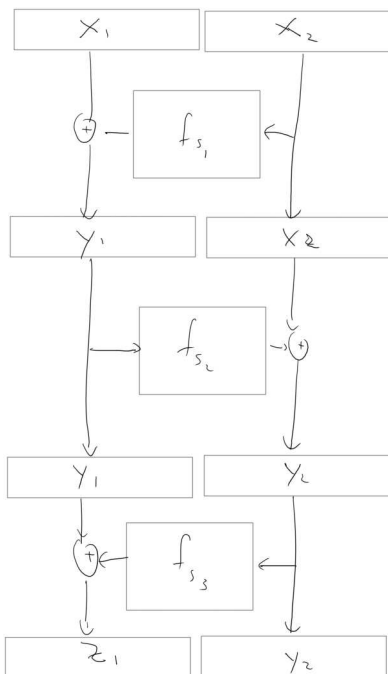


Figure 5.6: We build a PRP p on $2n$ bits from three PRFs $f_{s_1}, f_{s_2}, f_{s_3}$ on n bits by letting $p_{s_1, s_2, s_3}(x_1, x_2) = (z_1, y_2)$ where $y_1 = x_1 \oplus f_{s_1}(x_2)$, $y_2 = x_2 \oplus f_{s_2}(y_1)$ and $z_1 = f_{s_3}(y_2) \oplus y_1$.

We will not show the proof of this theorem here, but [Fig. 5.6](#) illustrates how the construction of a pseudorandom permutation from a pseudorandom function looks like. The construction (known as the Luby-Rackoff construction) uses several rounds of what is known as the **Feistel Transformation** that maps a function $f : \{0, 1\}^n \rightarrow \{0, 1\}^n$ into a permutation $g : \{0, 1\}^{2n} \rightarrow \{0, 1\}^{2n}$ using the map $(x, y) \mapsto (x, f(x) \oplus y)$. For an overview of the proof see Section 4.5 in Boneh Shoup or Section 7.6 in Katz-Lindell.

The more common name for a pseudorandom permutation is *block cipher* (though typically block ciphers are expected to meet additional security properties on top of being PRPs). The constructions for block ciphers used in practice don't follow the construction of [Theorem 5.6](#) (though they use some of the ideas) but have a more ad-hoc nature.

One of the first modern block ciphers was the **Data Encryption Standard (DES)** constructed by IBM in the 1970's. It is a fairly good cipher- to this day, as far as we know, it provides a pretty good number of security bits compared to its key. The trouble is that its

key is only 56 bits long, which is no longer outside the reach of modern computing power. (It turns out that subtle variants of DES are far less secure and fall prey to a technique known as **differential cryptanalysis**; the IBM designers of DES were aware of this technique but kept it secret at the behest of the NSA.)

Between 1997 and 2001, the U.S. national institute of standards (NIST) ran a competition to replace DES which resulted in the adoption of the block cipher Rijndael as the new **advanced encryption standard (AES)**. It has a block size (i.e., input length) of 128 bits and a key size (i.e., seed length) of 128, 196, or 256 bits.

The actual construction of AES (or DES for that matter) is not extremely illuminating, but let us say a few words about the general principle behind many block ciphers. They are typically constructed by repeating one after the other a number of very simple permutations (see Fig. 5.7). Each such iteration is called a *round*. If there are t rounds, then the key k is typically expanded into a longer string, which we think of as a t tuple of strings (k_1, \dots, k_t) via some pseudorandom generator known as the *key scheduling algorithm*. The i -th string in the tuple is known as the *round key* and is used in the i^{th} round. Each round is typically composed of several components: there is a “key mixing component” that performs some simple permutation based on the key (often as simply as XOR’ing the key), there is a “mixing component” that mixes the bits of the block so that bits that were initially nearby don’t stay close to one another, and then there is some non-linear component (often obtained by applying some simple non-linear functions known as “S boxes” to each small block of the input) that ensures that the overall cipher will not be an affine function. Each one of these operations is an easily reversible operations, and hence decrypting the cipher simply involves running the rounds backwards.

5.3 Encryption modes

How do we use a block cipher to actually encrypt traffic? Well we could use it as a PRF in the construction above, but in practice people use other ways.³ The most natural approach would be that to encrypt a message m , we simply use $p_s(m)$ where $\{p_s\}$ is the PRP/block cipher. This is known as the *electronic code book (ECB) mode* of a block cipher (see Fig. 5.8). Note that we can easily decrypt since we can compute $p_s^{-1}(m)$. However, this is a *deterministic* encryption and hence cannot be CPA secure. Moreover, this is actually a real problem of security on realistic inputs (see Fig. 5.9).

³ Partially this is because in the above construction we had to encode a plaintext of length n with a ciphertext of length $2n$ meaning an overhead of 100 percent in the communication.

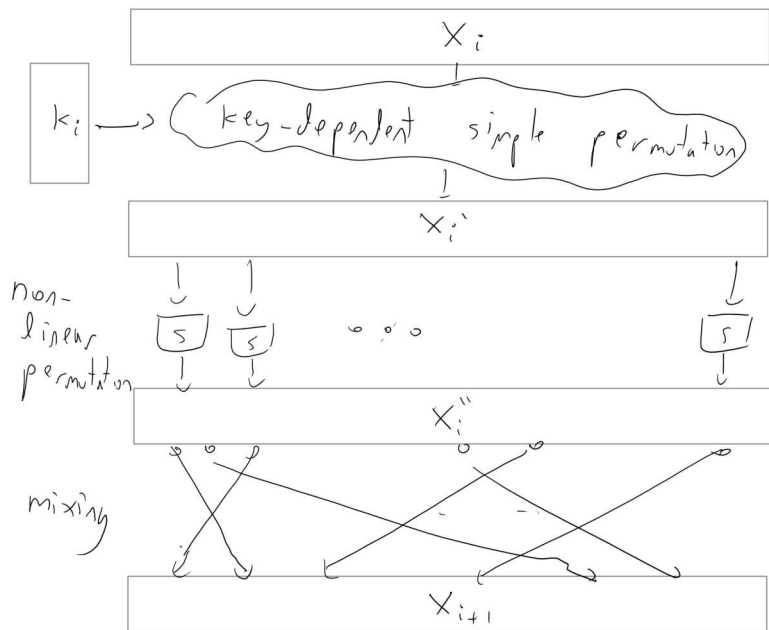


Figure 5.7: A typical round of a block cipher, k_i is the i^{th} round key, x_i is the block before the i^{th} round and x_{i+1} is the block at the end of this round.

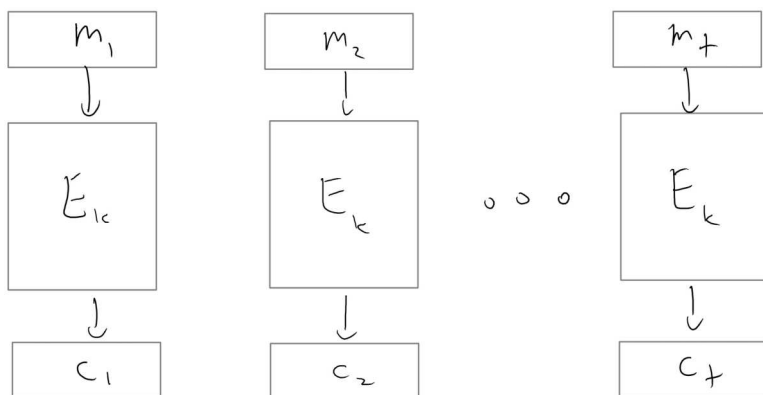


Figure 5.8: In the Electronic Codebook (ECB) mode every message is encrypted deterministically and independently

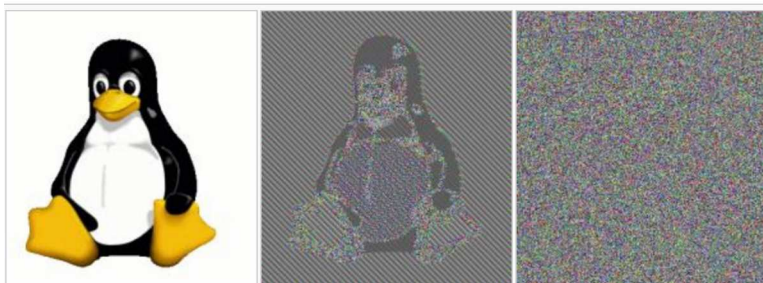


Figure 5.9: An encryption of the Linux penguin (left image) using ECB mode (middle image) vs CBC mode (right image). The ECB encryption is insecure as it reveals much structure about the original image. Image taken from Wikipedia.

A more secure way to use a block cipher to encrypt is the *cipher block chaining mode* where we XOR every message with the previous ciphertext (Fig. 5.10). For the first message we XOR a string known as the *initialization vector* or IV. Note that if we lose a block to traffic in the CBC mode then we are unable to decrypt the next block, but can recover from that point onwards. It turns out that using this mode with a random IV can yield CPA security, though one has to be careful in how you go about it, see the exercises.

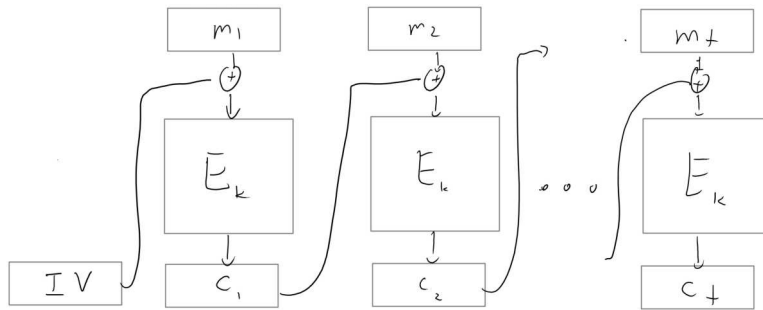


Figure 5.10: In the Cypher-Block-Chaining (CBC) the encryption of the previous message is XOR'ed into the current message prior to encrypting. The first message is XOR'ed with an *initialization vector* (IV) that if chosen randomly, ensures CPA security.

In the *output feedback mode* (OFB) we encrypt the all zero string using CBC mode to get a sequence (y_1, y_2, \dots) of pseudorandom outputs that we can use as a stream cipher. Perhaps the simplest mode is *counter mode* where we convert a block cipher to a stream cipher by using the stream $p_k(IV), p_k(IV + 1), p_k(IV + 2), \dots$ where IV is a random string in $\{0, 1\}^n$ which we identify with $[2^n]$ (and perform addition modulo 2^n). For a modern block cipher this should be no less secure than CBC or OFB and has advantages that we can easily compute it in parallel.

A fairly comprehensive study of the different modes of block ciphers is in [this document](#) by Rogaway. His conclusion is that if we simply consider CPA security (as opposed to the stronger notions of *chosen ciphertext security* we'll see in the next lecture) then counter mode is the best choice, but CBC, OFB and CFB are widely implemented due to legacy reasons. ECB should not be used (except as a building block as part of a construction achieving stronger security).

